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Disruption-free routing convergence

Computing minimal link-state update sequences

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Résumé

Avec le développement des applications temps-réel sur Internet, telles que la télévision, la voix sur IP et les jeux en ligne, les fournisseurs d'accès à Internet doivent faire face à des contraintes de plus en fortes quant aux performances de leurs services. Ces contraintes se traduisent sous la forme de conventions de service et définissent le niveau de service attendu d'un opérateur via divers indicateurs, comme les pertes de paquets ou la disponibilité du réseau. Les interruptions de service sont principalement causées par des changements topologiques (ajout/suppression de lien ou de routeur, changement de poids, ...), lesquels sont pourtant des événements courants dans les réseaux IP. D'une part, la topologie du réseau peut être régulièrement modifiée en fonction des besoins des opérateurs, pour procéder à des remplacements de matériel, des mises à jour système, ou encore dans le cadre de politiques d'ingénierie de trafic. Une étude menée sur l'épine dorsale du réseau Sprint rapporte que 20% des changements topologiques sont causés par des opérations planifiées. De plus, d'autres études révèlent que de telles opérations ont lieu fréquemment, mais celles-ci sont généralement effectuées de nuit, afin de limiter leur impact sur le trafic. Cela représente néanmoins un coût supplémentaire pour l'opérateur, et réduit sa capacité à améliorer le routage en fonction des fluctuations du trafic. D'autre part, les changements topologiques imprévus, tels que les pannes de liens ou de routeurs, sont également une source importante de problèmes de convergence. Cependant, leur impact sur l'acheminement des données peut être limité grâce à des techniques de re-routage rapide largement répandues.

Chacun de ces changements force les routeurs à recalculer leurs tables de routage, faisant ainsi entrer le réseau dans un état transitoire durant lequel des perturbations peuvent apparaître. Les spécifications des protocoles de routage à état des liens, Open Shortest Path First (OSPF) et Intermediate System to Intermediate System (IS-IS), ne fournissent aucun contrôle sur l'ordre de mise à jour des tables de commutation des routeurs. Cet ordre dépend à la fois des dynamiques de diffusion des messages de signalisation et des capacités de calcul de chaque routeur. Ainsi, le plan de commutation global du réseau peut être transitoirement incohérent, certains routeurs ayant déjà pris en compte la modification tandis que d'autres, plus lents ou plus éloignés, considèrent toujours la topologie initiale. Dans certains cas, les décisions de routages consécutives et antérieures à un changement topologique peuvent être conflictuelles, de sorte que plusieurs routeurs se considèrent alors l'un l'autre sur leur plus court chemin respectif vers une même destination. Ce phénomène, connu sous le nom de boucle de routage, augmente les délais d'acheminement des données et, selon le contexte de trafic, peut mener à des problèmes de congestion voire des pertes de paquets. Une telle baisse de performance est particulièrement regrettable lorsqu'elle survient à la suite d'une opération planifiée.

Afin d'évaluer l'ampleur de ce problème sur un réseau de production, nous avons engagé une collaboration avec l'opérateur Internet français RENATER. L'infrastructure réseau nationale de RENATER inclut 72 routeurs fournissant un accès Internet à la plupart des universités et organismes de recherche en France. Certaines de ces institutions participent au projet PlanetLab et, à ce titre, maintiennent des serveurs appelés nœuds PlanetLab que nous pouvons utiliser pour mener des opérations de mesures du réseau. Néanmoins, ces nœuds ne fournissent pas une couverture suffisante pour détecter efficacement la présence de boucles de routage. Comme première étape de notre collaboration avec RENATER, nous avons donc déployé 10 cartes Raspberry Pi pour compléter l'infrastructure PlanetLab existante. Ces appareils sont directement connectés aux routeurs afin d'assurer la fiabilité des mesures. De plus, nous avons mis en place un équipement supportant le protocole de routage IS-IS et capable d'établir une relation d'adjacence avec l'un des routeur de RENATER. Ce *listener* nous permet ainsi de détecter en temps réel et de maintenir un historique de l'ensemble des évènements topologiques affectant le routage sur ce réseau.

Notre première campagne de mesures actives sur le réseau de RENATER a eu lieu du 6 au 27 juin 2014, soit une durée de 21 jours. Pendant cette période, le *listener* nous a permis de détecter 1371 modifications topologiques dans le réseau, représentées par la réception de messages de signalisation non sollicités sur le *listener*. En moyenne, 63 évènements logique ont donc eu lieu chaque jour sur le réseau. Ce chiffre peut sembler très élevé, mais ne reflète pas nécessairement la fréquence des évènements physiques. En effet, le retrait d'un lien physique entre deux routeurs engendre deux évènements logiques, un pour chaque routeur. De même, l'extinction d'un routeur entrainera un nombre d'évènements égal à son degré, chacun de ses voisins détectant la coupure de son adjacence et émettant un message de signalisation pour en informer le reste du réseau.

Pendant ce temps, nos 10 points de mesures s'échangeaient à haute fréquence des messages de type Internet Control Message Protocol (ICMP), dans le but de fournir des données précises sur l'apparition et la durée des perturbations transitoires. Chaque Raspberry Pi était configuré pour envoyer un message vers chacun des autres toutes les 10ms, tout en enregistrant des informations de temps et de time-to-live (TTL) pour tous les messages émis et reçus. Les résultats obtenus permettent non seulement de montrer que des boucles transitoires apparaissent réellement, mais surtout que celles-ci ont un impact non négligeable sur le trafic, pouvant aller jusqu'à compromettre le respect des Service Level Agreements (SLAs) établis entre l'opérateur et ses clients. Nous pouvons conclure de cette campagne de mesures que les évènements topologiques, planifiés ou non, peuvent mener à des interruptions de service d'une durée de l'ordre de la seconde.

Afin de résoudre ce problème, ou d'en atténuer les effets, plusieurs solutions ont déjà été proposées à l'IETF[1]. Néanmoins, toutes se présentent sous la forme d'extensions à apporter aux protocoles existants, impliquant des modifications logicielles ou matérielles. De telles extensions pourraient prendre des années avant d'être effectivement déployés et utilisables. Pour pallier à ce manque, certains opérateurs ont défini des procédures pour dévier *en douceur* le trafic hors d'un lien ou d'un nœud, sur base de poids pseudo infinis, avant de déconnecter ce dernier. De telles procédures permettent d'éviter les pertes de paquets liées à l'absence temporaire de connectivité, mais n'ont aucun effet sur les boucles de routage transitoires.

En se basant sur des travaux de Francois et al.[FSB07], nous proposons des solutions algorithmiques efficaces pour prévenir l'apparition de perturbations transitoires dans le cas d'une modification planifiée sur un lien ou un routeur. Notre approche repose sur les fonctionnalités de base des protocoles de routage à état des liens, et ne requière donc aucune modification de ces derniers. Intuitivement, il s'agit de contrôler implicitement l'ordre de mise à jour des routeurs, à travers une modification progressive du poids d'un sous-ensemble de liens. Ainsi, des augmentations successives du poids d'un lien aura pour effet de forcer les routeurs les plus éloignés de ce composant à se mettre à jour avant les routeurs plus proches. Tout changement topologique peut être modélisé sous la forme d'une reconfiguration des poids attribués à un ensemble de liens du réseau. Par exemple, nous modélisons le retrait d'un lien par l'augmentation de son poids, depuis sa valeur actuelle jusqu'à la valeur minimale à laquelle il n'est plus utilisé pour acheminer des données dans le réseau (ou, plus simplement, jusqu'à la valeur maximale qu'il est possible de lui attribuer). Celui-ci pourra ensuite être retiré du réseau sans impact sur les décisions de routage. De la même manière, un nouveau lien peut se voir attribuée un poids très élevé lors de son ajout dans le réseau, lequel sera ensuite réduit jusqu'à la valeur prévue par l'opérateur. Le retrait d'un routeur peut également être précédé par l'augmentation des poids attribués à l'ensemble de ses liens sortants. Ce routeur ne sera alors plus utilisé comme transit, mais uniquement pour acheminer des données vers ou depuis les réseaux feuilles qui lui sont directement connectés. Enfin, le procédé inverse pourra être utilisé dans le cas de l'ajout d'un routeur. Pour prévenir l'apparition de boucles de routage, notre approche consiste à diviser ces modifications de poids en une séquence de mises à jour *sûres*, sans perturbations. Les mises à jour intermédiaires sont calculées de sorte qu'aucune boucle transitoire ne puisse apparaître lors de leur application, en supposant qu'elles soient appliquées dans l'ordre et séparées d'un intervalle de temps suffisant. Une solution simple répondant à ces critères consisterait à augmenter, ou diminuer, le poids de l'ensemble des liens affectés de 1 à chaque étape. Nous avons prouvé que de telles modifications ne peuvent jamais mener à l'apparition de boucles,

[1] IETF: Internet Engineering Task Force

et permettraient donc d'assurer une convergence sans incidents. Cette solution pourrait néanmoins nécessiter un grand nombre d'étapes intermédiaires, obligeant l'opérateur à attendre une durée considérable avant de pouvoir enfin effectuer l'opération prévue. De plus, ce type de reconfiguration à également un impact au niveau inter-domaine, forant les décisions de routage pour l'ensemble des préfixes BGP à être reconsidérés après chaque modification topologique. Par conséquent, au delà de la seule prévention des boucles, notre objectif est également de fournir les séquences de mises à jour les plus courtes possibles.

Dans [2], nous proposons un algorithme pour calculer des séquences de mises à jour de longueur minimale, prévenant toute boucle transitoire qui pourrait survenir lors de la modification du poids d'un unique lien. Ce premier algorithme fonctionne sur un mode *essai-erreur*, cherchant à maximiser l'amplitude de chaque modification tout en assurant l'absence de boucle, et repose de valeurs pivots, appelées *delta*. Une valeur delta est définie pour chaque routeur pour une destination donnée, comme la différence entre les distances depuis ce routeur vers la destination avant et après le changement topologique. Ainsi, un routeur dont les routes vers une destination ne sont pas affectées par le changement topologique aura une valeur delta nulle pour cette destination. Pour les autres routeurs, cette valeur représente la reconfiguration de poids minimum a appliquer au lien modifié pour que la décision de routage change, pour cette destination. Une modification plus faible sera donc sans effet sur ce routeur, tandis qu'une modification plus importante forcera le routeur à converger pour ne plus utiliser que ses nouveaux chemins vers la destination. Enfin, une modification égale à la valeur delta mènera à un état transitoire et à l'utilisation simultanée des chemins pre et post convergence. Dans le cadre de notre algorithme, les valeurs delta permettent de réduire significativement l'espace de recherche, le limitant à l'ensemble des valeurs delta, pour tous les routeurs et toutes les destinations. Il est donc possible de calculer des séquences valides dans un temps très limité (de l'ordre de la seconde sur du matériel de qualité standard), malgré la nature naïve de notre algorithme. Nous avons prouvé qu'aucune boucle ne pouvait survenir entre deux mises à jour successives de cette séquence, et que celle-ci était de longueur minimale. Nos évaluations, menées sur des topologies représentant des réseaux d'opérateurs réels, montrent que les séquences ainsi obtenues sont très courtes en pratique. Même sur des réseaux de grande taille, approximativement 95% des opérations de retrait de lien nécessitent en effet moins de 3 mises à jours intermédiaires.

Nous généralisons cette approche dans [1] et [3] aux modifications sur un routeur. Notre nouvel algorithme, appelé Greedy Backward Algorithm (GBA), est en effet capable de calculer des séquences de reconfigurations sans boucle pour n'importe modification sur un sous-ensemble des liens sortants d'un routeur, incluant de fait le cas de l'ajout ou du retrait du routeur entier. Notre algorithme fonctionne de la manière suivante. Dans un premier temps, il itère sur l'ensemble des destinations accessibles dans le réseau, détectant pour chacune d'elles la potentialité de boucle transitoire. Si de telles boucles sont détectées, l'algorithme extrait, sur base des valeurs delta des nœuds impliqués dans chaque boucle, un ensemble de contraintes représentant les conditions nécessaires et suffisantes pour prévenir celles-ci.

Ces conditions sont représentées sous la forme d'intervalles vectoriels, dont les composantes représentent les reconfigurations de poids à appliquer sur chacun des liens modifiés. La résolution de ce système de contraintes par une séquence de mises à jour de taille minimale constitue donc un problème multidimensionnel. De plus, les bornes de ces intervalles affichent un caractère asymétrique : s'il est nécessaire pour un vecteur intermédiaire d'être supérieur à la borne inférieure de l'intervalle sur chacune des composantes, il est en revanche suffisant que celui-ci soit inférieur à la borne supérieure sur une seule composante pour satisfaire la contrainte. Cette asymétrie s'explique par la réaction attendue des routeurs suite à l'application d'un vecteur intermédiaire satisfaisant une contrainte. Afin de prévenir l'apparition de la boucle associée à la contrainte, il est en effet nécessaire et suffisant que l'un des routeurs impliqués dans celle-ci se mette à jour (celui-ci n'utilisera alors plus que ses chemins post-convergence pour atteindre la destination), et qu'au moins l'un des autres routeurs de la boucle ne soit pas affecté par le vecteur intermédiaire (celui-là utilise toujours ces chemins initiaux pour joindre la destination). Dans la mesure o il suffit que le vecteur intermédiaire soit inférieur ou égal au delta d'un routeur sur l'une des composantes pour que celui-ci utilise toujours sont chemin initial vers la destination, la première condition nécessite que le vecteur intermédiaire soit strictement supérieur sur toutes les composantes au plus petit delta parmi les routeurs impliqués dans la boucle. A l'inverse, la seconde condition requière qu'au moins l'une des composantes du vecteur soit strictement inférieure au plus grand delta parmi les routeurs impliqués dans la boucle pour que celui-ci n'utilise aucun chemin post-convergence.

Face à de telles contraintes, un algorithme de recherche *en avant*, tel que celui présenté précédemment pour la reconfiguration d'un unique lien, serait confronté à un problème d'indéterminisme lié au choix de la composante permettant de satisfaire chaque contrainte. En effet, un tel algorithme serait incapable de déterminer *a priori* quelle composante devra rester inférieure à la borne supérieure de l'intervalle, afin d'obtenir une séquence de longueur minimale. Pour pallier à ce problème, notre algorithme GBA repose sur un mécanisme de recherche *en arrière*, partant de l'état final et cherchant à chaque étape le plus petit vecteur strictement supérieur aux bornes inférieures des contraintes restantes. Ce procédé permet d'obtenir une séquence de vecteurs de taille minimale satisfaisant l'intégralité des contraintes.

Nos évaluations montrent que les séquences produites par GBA pour le retrait d'un routeur sont à peine plus longues que celles pour un unique lien. Ainsi, même dans le cas d'un réseau d'opérateur de très grande taille, 90% des opérations de retrait de routeur requièrent moins de 5 mises à jour intermédiaires. De plus, diverses améliorations algorithmiques permettent de réduire la complexité temporelle de GBA en $O(N^4)$, voire $O(N^3)$ si la taille des séquences est bornée, et de maintenir un temps de calcul des séquences de l'ordre de quelques secondes au pire.

Cependant, l'application simultanée de mises à jour de poids d'amplitude différente sur plusieurs liens du réseau, nécessaire pour garantir la minimalité de la séquence, requiert de prendre en compte une nouvelle forme de perturbations transitoires. Ce type de mise à jour, qui consiste à augmenter ou diminuer le poids sur certains liens modifiés plus que d'autres, peut en effet mener à des phénomènes d'oscillation de routes, néfastes pour le trafic, ainsi qu'à des boucles non prises en compte par notre algorithme. Dans [3], nous présentons une heuristique modifiant légèrement les séquences produites par notre algorithme afin de prévenir l'apparition de telles boucles. Nos analyses expérimentales montrent que, bien que théoriquement plus longues, les séquences ainsi obtenues sont en pratique très proches, et bien souvent de même longueur que celles produites par GBA. Dans [1], nous étendons cette solution à l'ensemble des perturbations de routage, éliminant du même coup toutes les oscillations de routes qui pourraient survenir lors de l'application de la séquence de reconfigurations. Notre nouvel algorithme, nommé Adjusted Greedy Backward Algorithm (AGBA), permet en effet de définir des conditions nécessaires et suffisantes pour garantir la stabilité du routage malgré l'hétérogénéité des mises à jour intermédiaires. Ces conditions se présentent sous la forme d'un degré de liberté par rapport à une séquence uniforme, laquelle consisterait à appliquer des modifications de même amplitude sur chacun des liens sortants du routeur à une étape donnée. Nous avons prouvé que l'algorithme AGBA produit des séquences de taille minimale considérant ces nouveaux paramètres.

En pratique, les séquences calculées par AGBA se révèlent généralement plus longues que celles obtenues avec GBA, mais l'amplitude de ces différences se limite à 1 ou 2 éléments dans la plupart des cas. En termes de temps de calcul, nous n'avons pas constaté de différences significatives entre les performances des deux algorithmes. Ainsi, bien que notre solution puisse être utilisée via un outil centralisé de management du réseau, nous espérons que ces résultats pratiques encourageront son intégration directement dans les logiciels de routage.

Contents

Introduction

The growing popularity of real-time media services over Internet, such as TV broadcast, voice or video over IP, and gaming have changed the requirements of Internet Service Providers (ISPs) on the performance of routing protocols supporting those services. Non-Internet IP based services such as VPNs have also led to ISPs facing ever more stringent Service Level Agreements (SLAs), defining the performance of an ISP through various metrics such as service availability, packet losses and latency. Breaches in service availability are usually due to side effects of network topological changes, which are common events in large IP networks. On the one hand, the topology can be regularly reconfigured according to the needs of the operators, in order to perform hardware replacement, software upgrades or to apply traffic-engineering policies. Several studies reveal that such operations occur frequently. In particular, a study on the Sprint IP backbone reports that a significant proportion of topological changes are caused by scheduled operations. Maintenance tasks are mainly performed during nightly scheduled windows in order to reduce their impact on the traffic. However, this increases the cost of operating the network, and reduces its flexibility at the time when it is actually most likely to undergo traffic-engineering issues. It is thus not currently possible for operators to optimize routing policies according to traffic fluctuations. On the other hand, unplanned changes such as link or router failures are also a great source of transient convergence problems, yet their impact on the routing data plane can be limited thanks to widely deployed fast-reroute techniques.

Each topological change compels the routers to recompute their shortest path information, putting the network into an inconsistent state during which transient disruptions may occur. Specifications of current link-state routing protocols, Open Shortest Path First (OSPF) and Intermediate System to Intermediate System (IS-IS), provide no control over the routers update order. In practice, this order depends on flooding dynamics of control plane signalization packets and processing capabilities of each router. As a result, the global data plane of a network can be transiently inconsistent, some routers having already applied the modification and forwarding packets according to the new topology, while others still follow the initial one. In some cases, the routing decisions before and after the change may be conflicting, causing several routers to consider each other on the shortest path towards a given destination. This phenomenon, known as a routing loop, increases packet transmission delays and, depending on its duration and the amount of traffic involved, may lead to congestions and packet losses. Such performance drop during convergence is particularly unfortunate in the case of a scheduled operation, with no failed component black-holing traffic. Several methods have been proposed in the scientific literature and at the IETF to solve this problem. However, they require extensions to the OSPF and IS-IS protocols, implying software and/or hardware modifications. Even in a favorable perspective, such changes would possibly take years before being actually deployed. In the meantime, some ISPs have defined pragmatic procedures to smoothly reroute the traffic out of a link or a router, using pseudo infinite weights, before actually shutting it down. While efficiently preventing traffic black-holing due temporary lack of connectivity, this method does not solve transient routing loops and may exacerbate their impacts.

Based on previous works by Francois et al. [FSB07], we generalize the problem formalization and propose practical solutions to prevent transient disruptions caused by operations on a link or a router. These may either be used directly for scheduled events, or combined with fast-reroute technique to handle failures. Our approach only relies on basic principles of link-state routing, thus not requiring any protocol extension. Intuitively, it consists in implicitly controlling the routers update order through progressive weight modifications on a subset of links. For example, subsequent weight increments will force routers farther away from the modified component to update before routers close to it. Should the magnitude of these changes be finely tuned, it could spread the update of routers potentially involved in a loop across multiple steps. That is, to make a subset of the routers switch to their final routes, before they appear on the shortest paths of the others, so that no routing loop occurs. This operation can be repeated until the component is no longer used for transit in order to enable to its safe remove from the network without any routing disruptions.

Let X and Y be two routers, and D be a destination such that X initially reaches D through Y while the opposite holds after a topological change. If Y reacts first to the change, it will start sending its traffic towards D to X, and X will loop it back to Y. In this thesis, we demonstrate that there always exists an intermediate weight modification such that only X updates its route towards D, while Y still follows the initial routing plan. Hence, whatever the order in which X and Y process the modification, this loop cannot occur.

More generally, for all link or router-wide modifications, which could cause transient routing loop if performed abruptly, we propose solutions to associate with any modification a sequence of *safe*, loop-free weight updates. Intermediate updates are computed such that no transient loop could appear as they are applied, provided that two subsequent updates are separated by a *sufficient* amount of time. A basic, provably correct, solution would be to increase, or decrease, the weight on each affected link by 1 at each step. However, such solution would require a large amount of intermediate updates, thus potentially requiring the network operator to wait for a long time before the intended operation is actually performed. Hence, aside from the safety requirement, we also aim at providing update sequences of minimal length.

To this end, we define a theoretical framework for avoiding transient loops with sequences of intermediate weight reconfigurations. This framework is based on a set of loop-constraints, which represents necessary and sufficient conditions to prevent each loop occurrence for all destinations in the network. That is, a sequence prevents a transient loop if and only if it satisfies the associated constraint. These conditions allow us to devise an efficient algorithm for computing sequences of minimal length that provably prevent all transient loops for a given link or router modification. For any system of loop-constraints, we prove that there always exists a minimal sequence whose elements are strictly increasing or decreasing. However, aiming for minimality in the case of router-wide or multi-link operations may require to simultaneously perform different weight modifications on several outgoing links of the modified router. Such heterogeneous updates may jeopardize routing stability, causing route diversions as well as additional transient loops around the modified router that could not have occurred in case of an abrupt operation, i.e. without intermediate updates. We propose several variations of our minimization algorithm to address these problems with different tradeoffs between disruption avoidance and sequence lengths.

In chapter 1, we present the networking context of this work. We first provide a general overview of routing protocols by describing the three main families basics. Then we focus on link-state protocols, which are the most used for intra-domain routing in ISP backbones. We explain how these protocols react to topological changes, planned or not. For each kind of disruption that may occur during the convergence period, we describe existing solutions to prevent or mitigate the impact on the traffic. In particular, we detail the circumstances in which transient routing loops may occur, and analyze their impact on a real ISP network. We finally present the key idea proposed in [FSB07] to prevent these loops in the case of a single link modification. In chapter 2, we explain how this concept can be extended to handle router-wide modifications. We first consider the simple case of uniform weight modifications, i.e. performing the same weight modifications on each outgoing link of the router, whose calculation process is similar to single-link update sequences. We later generalize to the more challenging case of heterogeneous modifications. Focusing on *normal* transient loops first, we detail our main algorithm for computing minimal weight update sequences for any router-wide modifications. We then provide several algorithmic and technical solutions to prevent the additional inconsistencies related to the use of heterogeneous modifications. Finally, we describe several algorithmic improvements to allow for an efficient implementation of our solutions. In chapter 3, we thoroughly evaluate the performances of our solutions on real and inferred network topologies. After having shown how much each evaluation topology is affected by transient routing loops, we analyze and compare the length of the sequences produced by each of our algorithms. We then focus on the time required to compute such sequences, detailing the effects of each implementation improvement of the computing time distribution. We show that both the length of computed sequences and time necessary to obtain them are really limited on our set of evaluation topologies. Based on these observations, we discuss several schemes for a practical deployment of our solutions. Eventually, we conclude in chapter 4 and describe several possibilities to extend and improve this work. In particular, we aim at evaluating the benefits of our solution on real ISP networks.

Chapter 1

Context

Contents

1 Routing protocol basics

Routing is the process of selecting best paths in a network to enable transmitting contents from one or multiple sources to one or multiple destinations. Routing is performed in many kinds of networks, including telephone networks, electronic data networks and transportation networks. In the context of packet switching networks, routing is performed by dedicated devices, called *routers*, which are in charge of computing best paths according to a routing metric, such as bandwidth, delay, reliability or simply hop count. Routers store best path information in *routing tables*, or Routing Information Bases (RIBs), as a list of entries. Each entry associates a network destination with the path, or route, towards it. Although they are generally constructed by routers running *routing protocols*, additional entries denoted *static routes* can be manually supplied. Routing tables are not used directly for traffic forwarding, but instead to populate forwarding tables, or Forwarding Information Bases (FIBs), which are optimized for fast lookup. Forwarding tables contain the minimal information necessary to transmit outgoing traffic on the *best* interface. Each entry associates an address matching one or multiple destinations with an identifier of the next routing capable equipment, or *next-hop*, on the route towards them. In brief, the routing or *control plane* of a router draws a map of the network, while the *forwarding plane* decides how to handle incoming data packets.

Several types of routing exist to be used in different contexts. Very small networks, for example, may choose to rely on *static routing*, which consists in manually configuring the routing table of each router with an entry for every destination in the network. Fallback routes may also be specified in case the first ones become unavailable. This is however not suited for large networks that serve dozens or hundreds of destinations, and may frequently undergo topological modifications. Dynamic routing aims at solving this problem by constructing routing tables automatically, based on topological information carried by routing protocols. This allows the network to dynamically react to topological modifications, attempting to avoid failures and blockages.

Routing protocols consider that each router has a priori knowledge of networks directly attached to it, and define how this information is shared with the rest of the network. Practically, local information of each router is embedded in signalization messages to be transmitted to immediate neighbors. Upon receiving such message, a router updates its view of the network accordingly and retransmits this information to its own neighbors. Topological knowledge is thus recursively flooded throughout the network.

Based on the information they receive, routing protocols locally compute on each router the best path towards every reachable destination and construct the routing table accordingly. Such a best path is not necessarily the one minimizing the number of routers the packet has to cross. It may depend on speed or bandwidth available on each link, processing capabilities of routers, traffic flows passing through the network, as well as specific needs of the operator. It is hence possible to influence the routing decisions by configuring a strictly positive valuation, or weight, on each link. A best path between two nodes is then defined as the one minimizing the sum of the weights on its constituent links, rather than the number of link. These weights could be defined according to various criteria [FT03], such as the inversed capacity of each links. In this case, the more capacity a link has, the more *attractive* it is.

In order to compute such best paths, routing protocols rely on operations research and graph theory algorithms. The network is modelled as a directed weighted graph whose nodes usually represent routers and edges are adjacencies between routers. However, depending on the protocol, the information available on each node is not necessarily a complete view of the network, but may also be aggregated distance information from neighboring nodes. This information is used to compute the shortest paths for every destination and set the corresponding routing table entry. In hop-by-hop routing protocols, which are the most commonly used in IP networks, only the next-hop is actually stored in the table, even if the router has enough information to compute the full path. The *optimality principle* state that, if router J is on the optimal path from router I to K, then the optimal path from J to K falls along the same route. A consequence of this principle is that the routing decision the next-hop will make for a given destination will match the one the current node would have opted for. In the following, we detail the three major classes of routing protocols in IP networks. Distance-vector and link-state protocols are designed for intra-domain routing, i.e. inside an autonomous system, while a path-vector protocol is used for inter-domain routing.

1.1 Distance-vector routing

Distance-vector protocols do not require that routers have a complete knowledge on the network. Instead, they are based on vectors containing the distance from a given node to every destination in the network. Each router periodically informs its immediate neighbors about potential topological changes, by transmitting its own *distance vector*. While not having knowledge of the entire path for a destination, a router knows *how far* the destination is from each neighbor and can select the closest one as next-hop. Protocols based on distance vectors include Routing Internet Protocol (RIP) and Cisco's proprietary Interior Gateway Routing Protocol (IGRP).

In practice, best paths are computed using a distributed variant of the Bellman-Ford algorithm. This algorithm is based on the principle of relaxation, in which an approximation to the correct distance is gradually replaced by more accurate values, until eventually reaching the optimum solution. In a stable network, the approximate distance to each router is always an overestimate of the true distance, and it is updated at each step with the minimum of its old value with the length of a newly found path. Initially, a router only knows the distance to its immediate neighbors, which is the weight configured on each interface, and considers an infinite distance for all other destinations in the network. As distance vectors are spread in the network, a router progressively replaces infinite values with actual hop distances calculated from the vectors it receives. Also, previously stored distances may be updated to lower values as alternate paths are discovered. Every time a distance is modified, the neighbor that originated the message is stored as the new next-hop for this destination. This algorithm has a worst case complexity in $O(|N| \times |E|)$ when performed globally, but only requires $k \times (|N| - 1)$ operations on each node, where k represents the degree of this node.

Despite a fairly low complexity, distance-vector protocols come with significant drawbacks that prevent them from being used in large networks. These include a slow convergence as well as the chance of a long lasting routing loop being triggered after a failure renders a destination unreachable for the rest of the network. Indeed, depending on the signalization messages ordering, multiple routers may consider one another on the shortest path towards the unreachable destination. They gradually increase their distance for this destination until it reaches a pseudo-infinity value (16 in the case of RIP version 1), at which point the algorithm corrects itself, due to the relaxation property of Bellman-Ford.

1.2 Link-state routing

The purpose of link-state routing protocols was, at first, to overcome the limitations of distance-vector routing. Whenever a router is initialized, it floods the state of its links throughout the entire network, not only to its immediate neighbors. With such information, each router can draw a *connectivity map* of the network, showing how routers are connected to each other and which weight is configured on every link. Each node then independently computes the best path from itself to every possible destination in the network, and stores the next-hop for each path in its routing table. Also, if any router notices a topological change, a updated link-state message is sent to all other routers in the network, so that they can adjust their routing tables accordingly. The most commonly used link-state routing protocols are Open Shortest Path First (OSPF), which is supported by the Internet Engineering Task Force (IETF), and Intermediate System to Intermediate System (IS-IS), developed by the International Organization for Standardization (ISO). For the sake of simplicity, we use in the following OSPF

OSPF	IS-IS
Link	Circuit
Host	End System (ES)
Router	Intermediate System (IS)
Packet	Protocol Data Unit (PDU)
Hello packet	IS-to-IS Hello (IIH) PDU
Link-state advertisement (LSA)	Link-state PDU (LSP)

Table 1.1: Link-state protocols terminology

terminology to denote network components and interactions. Equivalences with IS-IS terms are given in Table 1.1.

Link-state protocols paths calculation relies on Dijkstra's algorithm. A router maintains three data structures: a *tree* containing routers that are "done", a set of *unvisited* routers and a *tentative distance* for each router in the network. The algorithm starts with the *tree* structure and the set of *unvisited* routers empty. Then, the initial router, on which the algorithm is performed, is added as the *root* of the *tree* and its *tentative distance* is set to zero. All other routers are marked as *unvisited* with a *tentative distance* set to infinity. Considering the initial router as the first *current* router, the algorithm repeatedly does the following:

- Update the *tentative distance* of each *unvisited* neighbor of the *current* router. Its new value is equal to the minimum of the old value with the distance via the *current* router. Also, if the *tentative distance* was modified, attach the neighbor to the *current* router in the tree (and remove any previous attachment).
- Select the *unvisited* router having the smallest *tentative distance*. Remove this router from the *unvisited* list and mark it as *current*.

These two steps are repeated until there is no more router left *unvisited*. When the algorithm ends, the shortest path from the initial router to any destination in the network is indicated by a path in the *tree*. Such a tree is known as Shortest Path Tree (SPT). The complexity of this algorithm mainly comes from extracting the smallest *tentative distance*, which may require cycling through all elements in the list. If used with a standard list, as in the original version, Dijkstra's algorithm runs in $O(|N|^2)$. However, efficient implementations usually rely on more sophisticated *priority queues*. The asymptotically fastest known variant is based on a Fibonacci heap and runs in $O(|E| + |N|log|N|)$.

Figure 1.1 represents the IP backbone of Internet2 Network [Int], which is freely available online. This network, operated by the not-for-profit organization Internet2, provides network services for many U.S. educational, research and government institutions. Interior

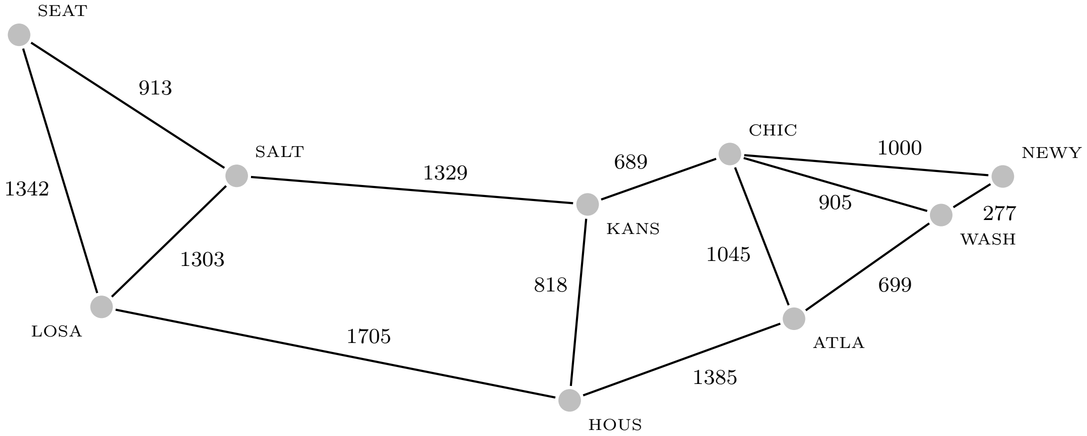


Figure 1.1: Internet2 IP network with IGP metrics (2009)

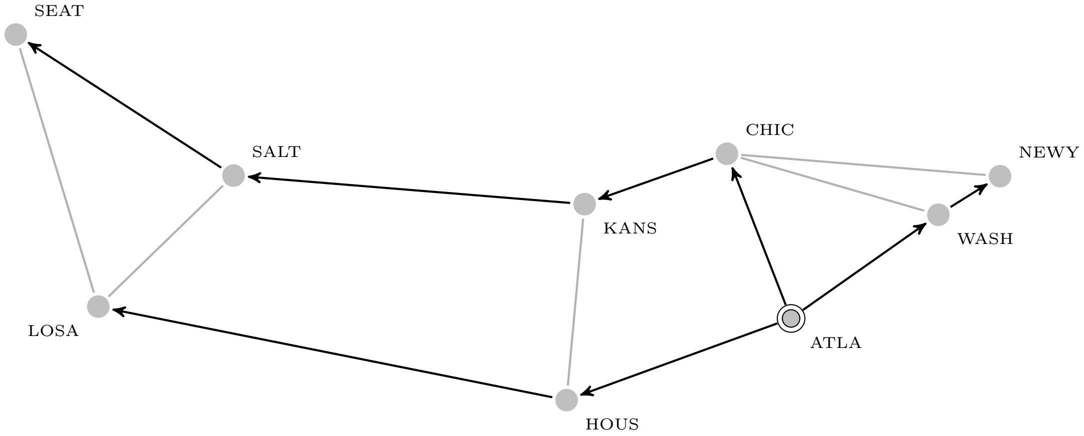


Figure 1.2: Shortest Path Tree rooted at Atlanta

Destination	Next-hop	Distance
SEAT	CHIC	3976
LOSA	HOUS	3090
SALT	CHIC	3063
HOUS	HOUS	1385
KANS	CHIC	1734
CHIC	CHIC	1045
ATLA	–	0
WASH	WASH	699
NEWY	WASH	976

Table 1.2: Routing table computed by the router at Atlanta

Gateway Protocol (IGP) weights, or metrics, configured on this network are directly based on fiber route kilometers, so that best paths calculated by routing algorithms minimize the actual geographic distance traveled by the signal. We chose to illustrate routing mechanisms on this topology for that particular reason; as it is easier the grasp the idea behind shortest path routing with link metrics being euclidean distances. The SPT obtained by running Dijkstra's algorithm on the router located in Atlanta, Georgia, and the associated routing table are represented on Fig. 1.2 and Table 1.2. The routing table states that packets processed at Atlanta that are headed towards Seattle, Salt Lake City or Kansas City shall be forwarded to the router at Chicago, while packets towards Los Angeles are to be sent to Houston and those towards New York City to Washington D.C.

OSPF and IS-IS protocols implement an extension for multi-path routing. This extension is called Equal-Cost Multi-Path (ECMP) and enable packet forwarding over multiple *best paths* that share the same shortest distance for a given destination. Multi-path routing potentially offers substantial increases in bandwidth by load-balancing traffic over multiple paths. Equal-cost paths may however differ on a variety of other metrics, such as Maximum Transmission Unit (MTU), latency and available bandwidth. This may impact the traffic if a single flow is split across several paths, as packets may be reordered and undergo constantly changing maximum size. Multi-path routing is thus generally performed on a per-flow basis. Routers calculate a hash over the packet header fields that identify a flow and forward to the same next-hop packets having the same resulting key.

Routers running link-state protocols having complete knowledge of the network topology, count-to-infinity and routing loops problems cannot occur the way they do with distance-vector protocols. However, it requires all routers to calculate their best paths based on exactly the same view of the network.

1.3 Path-vector routing

Distance-vector and link-state protocols are both designed for intra-domain routing. They are used to compute routing paths inside an Autonomous System (AS), but are not suited for inter-domain routing. Distance-vector protocols quickly become impractical as the number of routers increases, and even link-state ones show limitations when it comes to thousands of routers. Routing table calculations for such very large networks would require huge amount of resources, not to mention the heavy traffic load generated by signalization messages. Most of all, routing policies between ASes, maintained by Internet Service Providers (ISPs) of different types and countries, must consider various parameters aside from arithmetic shortest paths. Compared to intra-domain, inter-domain routing relies on a different perspective of the network. Instead of a plain graph of routers, the Internet is viewed as a hierarchical graph of ASes divided into tiers. Tier-1 networks are at the top of the routing hierarchy. They span across multiple continents and are all interconnected to each other via peering agreements. Tier-2 ASes buy transit from these Tier-1 networks to reach remote parts of the Internet, but may also establish peering relationships with other ASes of the same tier. These tier-2 networks provide Internet access to lower tier ASes and end users.

Path-vector protocols provide mechanisms to deal with these complex interactions and compute compute routing paths through the whole Internet. Similarly to distance-vector, every AS advertises its view of each prefix to its neighbors, except that routing table entries contain full paths towards each destination AS rather than a simple metric. The concept of routing metric as a global level of attractiveness does not make sense for inter-domain routing, for the attractiveness of a relationship varies across ASes. It indeed depends on commercial agreements between ASes as well as geopolitical considerations. Routing decisions of an AS are hence taken based on the full paths announced by neighboring ASes, and are shared by all routers within the AS.

Border Gateway Protocol (BGP) is the most widely deployed protocol for inter-domain routing in the Internet. It is often classified as a path-vector routing protocol, even though it does not completely satisfy the principles described above. In particular, routing tables of a given AS are only partially shared with the neighbors, based on commercial relationships.

2 Convergence of link-state protocols

We now focus on link-state protocols and, in particular, the convergence period that follows each modification of the network topology. These changes can be caused by network failures but also maintenance operations. For example, a study on the Sprint IP backbone [MIB+08] reports that 20% of topological changes are caused by maintenance operations. In IP over optical networks, the topology can be regularly reconfigured according to the need of the operators [PDRG02]. Another possible kind of topological change is the intentional modification of IGP weights for traffic engineering purposes [FT02] in order to optimize routing according to traffic fluctuations.

Formally, we denote as topological change any modification in the network that could have an impact on the intra-domain routing tables calculated by routers within this network. Such change can be a reconfiguration of the IGP weight associated to a router interface, or the addition, or loss, of an adjacency relationship between routers. In the following, we denote the former as a link weight reconfiguration, or simply *weight reconfiguration*. We also use the terms of *weight increment* and *decrement* so as to specify the direction of the modification. As for the latter, we split up the definition in different sub-cases. If an adjacency relationship is simultaneously established between one or several routers in to network with a new router, that was not part of the network before the change, we use the term *node startup*. Respectively, we denote as *node shutdown*, the simultaneous loss of all adjacency relationships with a given router. Finally, we denote as *link startup*, respectively *shutdown*, the addition, or loss, of an adjacency relationship between two routers that does not change the total number of routers in the network.

In networks running link-state protocols, topological changes always triggers a reaction in the *control plane* of the routers. New Link-State Advertisements (LSAs) are flooded and routers update their RIBs accordingly. However, the actual impact of a topological change on the *data plane*, thus on the traffic, depends on the nature of this change and the conditions in which it occurs. If it results from a logical modification that has no impact on the global network connectivity, the router being modified can instantly start spreading updated information to the rest of the network. For example, an operator decides to reconfigure the IGP weight associated to an interface of a given router. At the time the command is passed (or with no significant delay), the router starts recalculating its shortest paths and sends to its neighbors an LSA containing the updated information. In practice, both actions are performed at the same time by separate processes. Until new paths have been calculated and pushed to the FIB, traffic keeps on being forwarded along the initial paths.

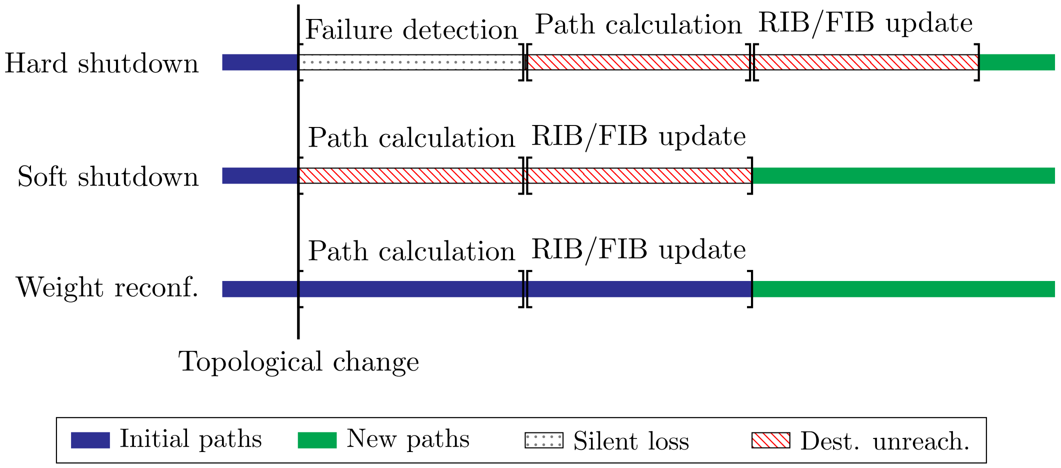


FIGURE 1.3: Traffic forwarding on a router undergoing a local topological change

On the other hand, link or router failures (hard shutdown) are not instantly detected by surrounding devices. In OSPF [Moy98, CFML08], an existing adjacency is considered *down* only if no *Hello* packet is received within the *Router Dead Interval*. By default it is equal to four times the *Hello Interval*, which is itself of either 10 or 30 seconds, depending on the network type. In IS-IS [ISO02], an adjacency is removed if a neighbor is not heard of within the *Holding Time* interval, by default equal to three times the *Hello Timer* (3 or 10 seconds). As long as the logical adjacency exists, traffic continues to be normally sent through the link. This phenomenon is referred to as a *transient black-hole*, as it may cause a large amount of traffic to be lost with no error being triggered.

Once an adjacency comes down on a router, be it due to a component failure or the result of a configuration command issued on this router, every entry relying on it is removed from the RIB and FIB. LSAs are then sent to announce this change to the rest of the network and new paths are calculated for each affected destination. In the meanwhile, further packets received by the router and headed towards such a destination are dropped by the router, which notifies senders with Internet Control Message Protocol (ICMP) *destination unreachable* error messages. Normal traffic forwarding is only resumed when new entries have been pushed to the FIB, assuming that a path towards the destination still exists in the network. The duration between the removal of the initial entries and the addition of the new ones is mainly characterized by two factors: the new paths calculation, which depends on the number of routers in the network; and the RIB/FIB update time, depending on the number of prefixes affected by the change.

The traffic forwarding states, towards an affected destination, on a router that undergoes a local topological modification is summarized on Fig 1.3. In the following sections 2.1

and 2.2 we present existing solutions that aim at reducing the traffic loss period (silent loss and destination unreachable) after a link or node shutdown.

2.1 Fast failure detection

From the explanation provided in the previous section, the delay before a failure is detected directly depends on the interval between *Hello* packets. Sending more frequent *Hello* packets would hence naturally speed up failure detections. However, this also means increasing the signalization traffic. A *Hello interval* too narrow can increase the probabilities of network congestion, possibly causing several consecutive *Hello* packets to be lost. False breakdowns resulting from this situation may be more harmful for the network than a slower detection of actual failures. When an adjacency goes down, every further data packet that ought to be forwarded on it is dropped and new routes are calculated. False positives thus increase the CPU load on the routers and cause traffic losses. The problem of finding better *Hello interval* values, which would provide both fast failure detection and low chances of network congestion, has been investigated in [AJY00] and [GRcF03]. The authors state that *Hello intervals* can be reduced to much lower values than those specified in protocols standards. Even though optimal values depend on the physical constraints of each link, such solution makes it possible to detect link failures within a few seconds in most cases.

Alternatively, failure detection can rely on the link layer in certain circumstances, hence avoiding the need of heavier signalization traffic. Synchronous Digital Hierarchy (SDH) and Synchronous Optical Networking (SONET), which are commonly used in optical networks, provide inbuilt alarm mechanisms that triggers if either no bit transitions are detected (Loss of Signal (LOS)), or the received data does not match the framing pattern (Loss of Frame (LOF)) during a given time interval. The router linecard can thus detect a failure in less than 10 milliseconds and transmit the information to the main CPU. Experiments performed by Francois et al. [FFEB05] show that the total detection delay is lower than 20ms in most cases, and barely exceeds 50ms for worst cases.

Finally, failure detection can be performed with low signalization overhead on any type of network using Bidirectionnal Forwarding Detection (BFD) ([KW10a, KW10b]), a dedicated protocol recently standardized by the IETF. BFD relies on encapsulation to allow for a rapid detection of link failures at any layer and over any media. It was primarily designed to provide faster notification of failing adjacencies for routing protocols, but also has many other use cases, which include virtual circuits, tunnels and MPLS Label Switched Paths. BFD has no neighbor discovery mechanism, but establishes point-to-point sessions between pre-defined systems. When used in conjunction with a routing protocol, BFD sessions are established upon request by the IS-IS or OSPF implementation. Depending on their ability to quickly proceed BFD packets, both systems then agree on the operating mode to be used in the session. BFD has two operative modes, *Asynchronous* and *Demand*, that can be used independently in both directions and modified in real time in order to handle unusual situations. In *Asynchronous* mode, the systems periodically transmit BFD control packets to one another. If a system does not receive any packet for given duration, it assumes that the link broke down. While this mode is similar to the inbuilt failure detection method of routing protocols, it differs in its capacity to dynamically adapt to specific constraints of each link. In *Demand* mode, it is assumed that there exists another way to ensure the connectivity in this session, and no more control packets are sent after the session is established. Either system may still request the connection to be explicitly verified by sending BFD control packets. In addition to these operating modes, BFD also provides an *Echo* function that may be called at any moment, independently of the current mode. This function makes the system transmit a stream of BFD *Echo* packets in such a way that the remote one sends them back through its forwarding plane. If too few of these packets are received, the link is considered down. Overall, BFD might not be as fast as a SDH/SONET alarms, but allows for more flexibility and is usable in any environment.

2.2 Fast reroute mechanisms

When a router detects the failure of an adjacency, it initializes a notification process by sending updated LSAs to its neighbors, and starts calculating new shortest paths. Until these new path are computed and the corresponding entries updates in the FIB, packets for destinations that were previously reached through the failed component are dropped by the router. The amount of lost traffic hence directly depends on the time required to recompute the SPT. As mentioned in Sec. 1.2, an implementation of Dijkstra's algorithm has a computational complexity of $O(|E| + |N|log(|N|))$ at best. Such complete SPT calculation can delay the convergence by a few seconds in large networks, and causes more incoming packets to be dropped by the router. However, in the case of a link or router failure, most of the network topology remains the same. It is thus possible to re-use the previous SPT in order to speed up the convergence. This algorithmic optimization to shortest path calculation is known as Incremental Shortest Path First (ISPF) [MRR79]. ISPF analyses the impact of the topological change on the previously computed SPT in order to minimize the amount of additional computation required. For example, if a link that belongs to the previous SPT goes down, ISPF limits the shortest path computation to the impacted subgraph, and re-uses the non-impacted region of the previous SPT. Also, if the link is not used in the previous SPT, then the whole

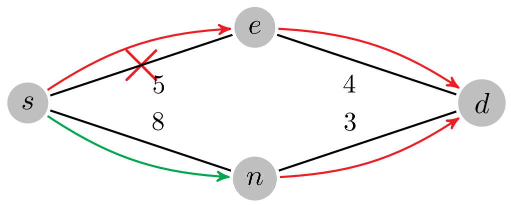


Figure 1.4: Loop-free alternate

shortest path calculation can be skipped as the old SPT is still valid. ISPF can thus greatly reduce the time required to compute new shortest paths towards each affected destination. In addition, recent works on multipath routing [MFB+11] have devised even more efficient algorithms.

In order to further reduce the unreachability period, it is also possible to rely on predetermined backup paths. If a failure occur on a link or router for which a backup path exists, further traffic that should be forwarded via the failed component is sent along the backup path instead. This procedure is known as *fast reroute*, as it prevents the traffic from being dropped while new shortest paths are computed. An alternative paths avoiding a failed component is called a *repair path*, and component for which such a repair path exits are said to be *protected*. Such protection mechanisms are often local, which means that the repair paths for a given protected component originate at the router immediately upstream of that component. This is motivated by the fact that packets continue to be forwarded along the initial forwarding path until a new one has been computed. Hence, it is sufficient that the last router before the failure has a backup solution to reach the destination in order to prevent any packet from being dropped. Local protection is not necessarily optimal in terms of routing, but it limits the number of repair paths to be computed, and still provides a decent alternative to dropping packets. In practice, repair paths can be classified into three main categories: purely local, single-hop and multi-hop.

Purely local repair paths are ECMPs that do not contain the failed component. Such paths are both straightforward and optimal repair paths. They should be used whenever available for they do not require any additional computation and match the new paths that will be calculated by the routing algorithm, thus preventing any further disruption.

Loop-free alternates

If no safe equal-cost path exits on the router adjacent to the failure, but a direct neighbor has a shortest path that does not include the failed component, incoming traffic could be forced towards this neighbor. A direct neighbor providing a single-hop repair path is called a Loop-Free Alternate (LFA) [AZ08]. Considering the simple topology on Fig 1.4, when router s computes its shortest path towards d, it determines to use router e as its primary next-hop. If LFA is enabled, s looks for an alternate next-hop to reach d, and determines that it could also send its traffic towards d through its link to n. Router s thus adds n as its next-hop for destination d. Then, if the adjacency between s and e comes down, s stops sending the traffic towards s to e and immediately switches to the alternate next-hop n. The traffic continues to be forwarded to n until a new SPT is computed and the FIB entries are updated. However, such a suitable LFA does not always exist. It depends on the topology and the component to be protected. Should the weight configured the link from n to d have been 20 instead of 3, using n as an alternate next-hop would have caused the traffic to loop between s and n.

[AZ08] defines inequalities to verify whether or not a given neighbor is a valid LFA. Let $C(A, B)$ denote the shortest distance between two arbitrary routers A and B, a neighbor N provides **link protection** for the primary next-hop E of router S towards destination D if and only if:

$$C(N, D) < C(N, S) + C(S, D) \tag{1.1}$$

This first inequality states that no shortest path from an alternate next-hop N towards D traverses router S.

Besides, the same neighbor N also provides **node protection** for E if and only if:

$$C(N, D) < C(N, E) + C(E, D) \tag{1.2}$$

This second equation ensures that the shortest paths from N to D do not include router E either. Note that, in the example on Fig 1.4, $C(n, d) < C(n, e) + C(e, d) < C(n, s) + C(s, d)$, so that n protects both link (s, e) and node e. In terms of calculation, retrieving the distances mentioned above requires to compute an additional Shortest Path DAG (SPDAG) from the perspective of each direct neighbor of S, which can be done efficiently using techniques presented in [MFB$^+$11]. The applicability of this LFA mechanism is further discussed in [FFS$^+$12]. Besides, real use case illustrations and operational management requirements are provided in [LDF$^+$].

ECMP and LFA offer the simplest repair paths and are usually preferred over any other fast reroute mechanism whenever they are available. [SB10b] indicates that around 80% of failures on real-world IGP networks can be covered using only these two methods. For the remaining 20%, multi-hop repair paths are required. These involve more complex mechanisms, both to compute the repair paths and to forward the traffic along them.

U-turn Alternates[Atl06] is an extension to LFA that aims at increasing the failure coverage by looking for LFAs one hop further away. In addition to its primary next-hops, a router S computes for each prefix an alternate next-hop to be used if the primary

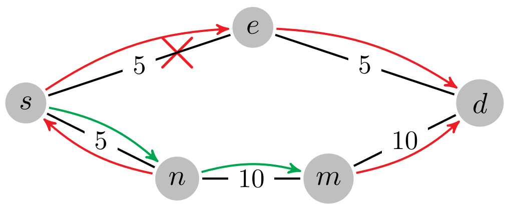


Figure 1.5: U-turn alternate

one fails. This alternate next-hop can either be an LFA or, if no such neighbor exists, a U-turn alternate. A U-turn alternate does not satisfy inequality 1.1, hence uses S as a primary next-hop towards the destination prefix, but has itself a node-protecting LFA for its primary next-hop, i.e. router S. This mechanism requires U-turn alternates to support U-turn themselves, in order to forward U-turn traffic coming from S to their own LFA, rather than sending it back to S. Identification of U-turn traffic, by a U-turn alternate, may be either implicit or explicit. Implicit identification requires no modification to the packets. If a U-turn capable router receives a packet headed towards a given destination from its primary next-hop for this same destination, it identifies the packet as a U-turn packet and forwards it to its LFA. On the other hand, explicit identification requires U-turn packets to be marked as such by the router sending them to the U-turn alternate. Explicit packet marking is used when hardware restrictions or particular deployment conditions make implicit identification unrealistic.
In Fig 1.5, router s has no LFA to protect its next-hop for destination d, because its only other neighbor n uses s as its primary next-hop. However, if both s and n support the U-turn mechanism, s could use n as a U-turn alternate. Then, if its primary next-hop e fails, s can forward the traffic towards d to n, which will send it to m, its LFA protecting s for destination d. Note that the node-protecting condition on the LFA of the U-turn alternate ensures that the traffic never loops back to s. It does not however guarantee that n is a node protecting U-turn alternate for s. For a U-turn alternate to also provide node protection for the primary next-hop e of s, it is necessary to ensure that e is not on the shortest path towards d of the U-turn alternate's LFA.
[SB10b] states that U-turn alternates, and 2-hops repair paths in general, increase the coverage to around 98% of link or node failures in real-world IGP networks.

Bryant et al. [BFP+14] proposed a new extension to LFA, called Remote Loop-Free Alternate (RLFA). RLFA increases the coverage of LFA against link failures by providing additional virtual links to the repairing node. These virtual links are in fact tunnels, based on IP-in-IP [Sim95, HLFT94] or MPLS-LDP [AMT07, RTF+01] encapsulation, which carry the rerouted traffic to some *staging point* in the network. Such points are selected so that, in the absence of concurrent failures, the traffic will travel from the staging point to its destination over normal forwarding paths without looping back. Formally, staging points satisfy the LFA inequality 1.1, whence the term of *remote* LFA.

If no normal LFA exist, a set of suitable staging points is calculated based on the following criteria: a staging point must be reachable from the repairing router S without traversing the failed link; and the shortest paths from a staging point to the destination D must not include the failed link. The set of nodes satisfying the first criterion is denoted *extended P-space* of S, while the nodes that meet the second are in the *Q-space* of D. The intersection of the *extended P-space* and the *Q-space* thus represents the set of valid staging points. In practice, the *Q-space* of the primary next-hop of S is used as a substitute for the *Q-space* of each destination reached through that next-hop. This approximation strongly reduces the complexity of calculating staging points, at the expense of potentially suboptimal repair paths. Any valid staging point, satisfying both criteria, can be chosen as a RLFA. However, it is recommended that the closest one from the repairing router is selected, as this maximizes the load balancing possibilities for the traffic exiting the tunnel.

RLFA extension brings a complete coverage against link failure in symmetrically weighted networks. Since RLFAs do not necessarily satisfy inequality 1.2, protection against node failures is not ensured. Hence, if RLFA is used to provide a repair path in case of node failure, the rerouted traffic may loop. Several hints are given in [BFP+14] as to minimize the probability of a loop, and the problem is further discussed in [SGH+14] and [BFPS07].

More recently, the development of source routing with the *segment routing* framework [FPB+14] has opened new possibilities in terms of repair paths. Using a list of segments, it is possible establish repair paths following the natural backup paths that would be used after the convergence, thus providing complete link and node protection for any topology [FFB+14].

Other fast reroute mechanisms

Failure Insensitive Fast Rerouting (FIFR) [NLY+07] takes the idea behind fast-reroute approaches from a different perspective. Since most of the topological changes due to link failures are transient events, it is likely that the topology will eventually fall back to its initial state, so that accelerating routing convergence would only cause more instability. Hence, instead of providing temporary forwarding paths while the network converges to its new routing state, FIFR prevents the convergence and forwards the traffic along backup paths until the topology returns to normal. This makes the network virtually insensitive to failures, as long as backup paths exist.

In practice, FIFR is similar to a local fast reroute mechanism, but for its ability to hold the LSA a router would transmit upon detecting a failure on an adjacency. This prevents the rest of the network from knowing about the change, and thus from converging. Traffic

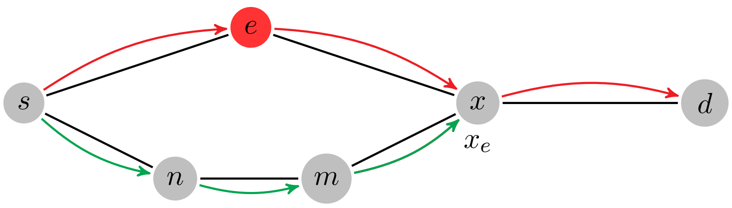


FIGURE 1.6: Not-via addresses

forwarding along backup paths relies on implicit failure detection. Upon receiving a rerouted packet, a router can infer a failure based on the unusual interface the packet came through. This router thus forwards the packet using a precomputed, interface-specific, backup table along a repair path. If the failed adjacency comes up again within a reasonable delay, denoted *suppression interval*, forwarding resumes over the recovered link, otherwise an LSA is propagated so that the network can converge to the new optimal routing state.

FIFR comes with roughly the same limitations as U-turn implicit packet identification. It requires the network to span over a single routing area, with point-to-point links and symmetric IGP weight assignments. Besides, FIFR only provides protection against link failures.

Fast-reroute mechanisms generally assume that only the neighbors of the failed component are aware of that failure. Hence, in order to provide a repair path, these routers have to steer the traffic around the failure despite other routers being unaware of its nature and location. Even though it could be possible for a router to implicitly infer the location of the failure in some situations, such solution limits the applicability of the fast-reroute mechanism. The idea behind Not-Via [BPS13] is to overcome this limitation by explicitly identifying the network component to be avoided. This method relies on a single level on encapsulation adding to each rerouted packet a special address, denoted Not-Via address. The Not-Via address indicates both the component that the repair path must avoid and the end of this repair path, which is the router directly downstream of the failed component on the shortest path towards the destination. At that point, the encapsulation is removed and the packet reaches its destination using normal forwarding paths. Since the decapsulating router is always closer to the destination than the encapsulating router, the packet will not loop.

Consider the network shown in Fig. 1.6 with a uniform valuation on the links. Upon detecting a failure of router e, router s can no longer forward the traffic towards d via its primary next-hop for this destination. Instead, it encapsulates every further packet headed to d with the address x_e, which is a special address of router x that cannot be reached from e. Provided that the network is not partitioned due to the failure of e, the encapsulated packets will reach router x through the best path from s that does not include router e, that is $s \rightarrow n \rightarrow m \rightarrow x$. At router x, these packets will be decapsulated and normally forwarded to d.

The Not-Via mechanism can protect against any link or node failure in the network, as long as the network is still connected after the failure. Rerouted packet are encapsulated to the address of the router just *behind* the failure that is not reachable from the previous router on the shortest path to the destination. Let us denote as A the primary next-hop of router S towards destination D, and B the primary next-hop of A for the same destination. If router A fails, the rerouted traffic from S is sent to B_A, the address of router B that is not reachable from A. On the other hand, if only the link (S, A) breaks down while router A itself is still available, the traffic can be steered to A_S instead, thus avoiding only the failed link. This mechanism however requires that all routers on a repair path have a route to the Not-Via address. Every router hence has to compute $N - 1$ additional SPDAG, one for the case of each other router having to be avoided. These SPDAGs can be calculated efficiently using ISPF, but still represent significant extra computations.

The above solutions aim at providing repair paths for IP traffic passing through the network. However, more and more ISPs now rely on ingress-to-egress Label Switching Path (LSP) tunnels to carry transiting traffic throughout their network. Since these tunnels usually have different service level requirements, it may be interesting to have specific backup paths that ensure continuity of the service level in case of failure.

An example of protection tunnel for the LSP from A to E is shown on Fig 1.7. If a failure occurs on link (B, C), router B will encapsulate the traffic associated to this LSP and send it on the backup tunnel (green dashed line) through F and G. Upon reaching D, the traffic is decapsulated and transmitted to E along the normal LSP. The start point of such a backup tunnel, which is router B in this example, is referred to as the Point of Local Repair (PLR) and the end point is called Merge Point (MP).

The IETF defined procedures to establish backup LSP tunnels for local repair of LSP tunnels [PSA05]. These procedures are based on two repair methods. The first method is called *one-to-one* backup. A backup LSP is established that intersects the original LSP downstream of the point of link or node failure. On each router providing a point of local repair, a separate backup LSP is established for each LSP that is backed up. The second method is called *facility backup*. It takes advantage of the Multiprotocol Label Switching (MPLS) label stack in order to back up multiple LSPs with a single repair LSP tunnel, called a *bypass tunnel*. Such a bypass tunnel can provide fast-reroute for all LSPs that pass through the point of local repair and through a common router downstream of the failed component. In both cases the repair paths can either be defined manually by a network operation, or computed using a Contrained Shortest Path First (CSPF) algorithm [Zie12].

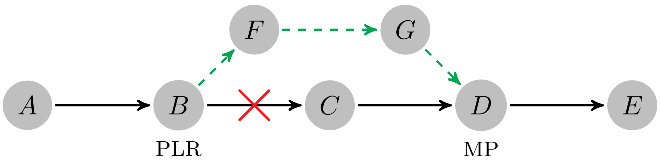


FIGURE 1.7: MPLS fast reroute

Protection schemes have also been investigated in various contexts aside from fast reroute. Optical switching networks, for example, have their own survivability mechanisms implemented at the physical layer [RSS09]. SONET/SDH architectures rely on 1:1 protection mechanisms, known as ring-based schemes, to provide quick recovery times in case of failure. Ring-based schemes however require that half the network capacity be dedicated to protection. Other protection mechanisms include mesh-based schemes, offering better capacity efficiency at the expense of slower recovery times, and preconfigured protection cycles [GS98, KAJ09], denoted *p-cycles*, which aims at combining the benefits of both ring- and mesh-based techniques. These mechanisms complement routing layer protection, for they provide fast and efficient failure recovery in most cases without being, however, able to handle all types of failure.

3 Transient routing loops

As we showed in the previous section, various solutions exist on the forwarding plane to overcome failure events, providing temporary path to safely carry the traffic while the routing plane converges to a new stable topology. However, the distributed nature of link-state routing protocols may lead to transient traffic disruptions during the convergence period [HMMD02].

In a stable situation, the shortest paths used by each router are consistent with the rest of the network, for they are computed based on exactly the same map of this network, represented as a Link-state Database (LSDB). Whenever a topological change occurs, the new information is flooded in the network using LSAs and integrated to the LSDB of each router. Yet these LSA are subject to the same propagation delays as any other data packet passing through the network, and do not spread instantly to every other router. Different versions of the LSDB may thus simultaneously exist in the network. Besides, the time required to recompute the RIBs and update the FIBs may vary across routers, some having more processing capabilities than others. As a result, a subset of the routers may forward their traffic according to the new topology, while the rest still follow the initial one. Depending the update order among affected routers, this situation may lead to routing loops. These are often referred to as *transient loops* for they disappear as soon as the last router involved has converged to the new topology [GSB+12].

A transient routing loop is characterized by several routers, in different update states, considering each other on the shortest path towards a given destination. Traffic headed for the same destination that comes through one of these routers may thus be caught into the loop. At best, that is if the routers involved in the loop update their FIBs fast enough, the traffic will reach the destination with a delay. However, the time-to-live (TTL) of affected packets, whose value is decremented for each router traversed, will quickly reach 0 and cause the packet to be dropped. Moreover, looping packets increase the load on links and routers, eventually causing congestion issues that will also affect the traffic towards other destinations [PZMH07, ZMMW07, WMW+06].

3.1 Illustration

Let us consider again the Internet2 network (Fig. 1.1) and describe step-by-step how a transient loop could occur between Chicago and Atlanta for the traffic headed to Seattle. Initially, the router at Atlanta forwards its traffic towards Seattle on the shortest path passing through Chicago, Kansas City and Salt Lake City. If the link between Chicago and Kansas City breaks down, or if its weight is *sufficiently* increased, a different path

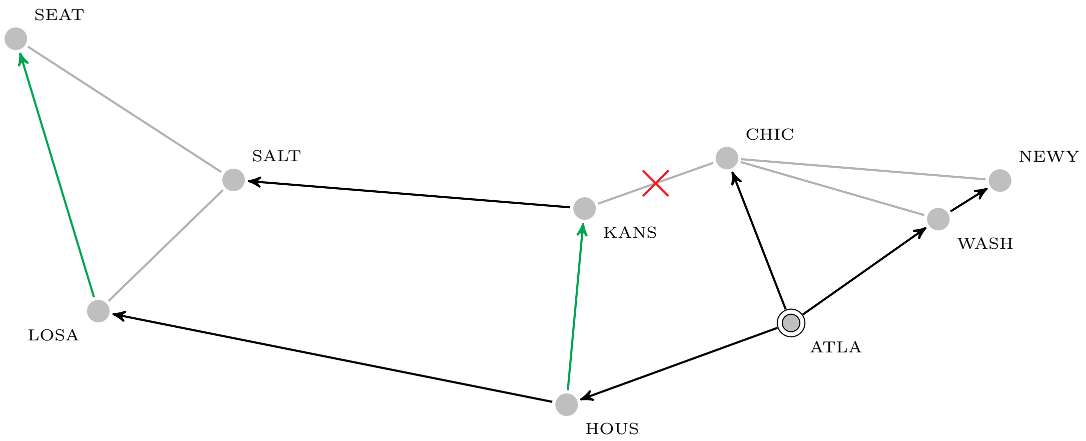
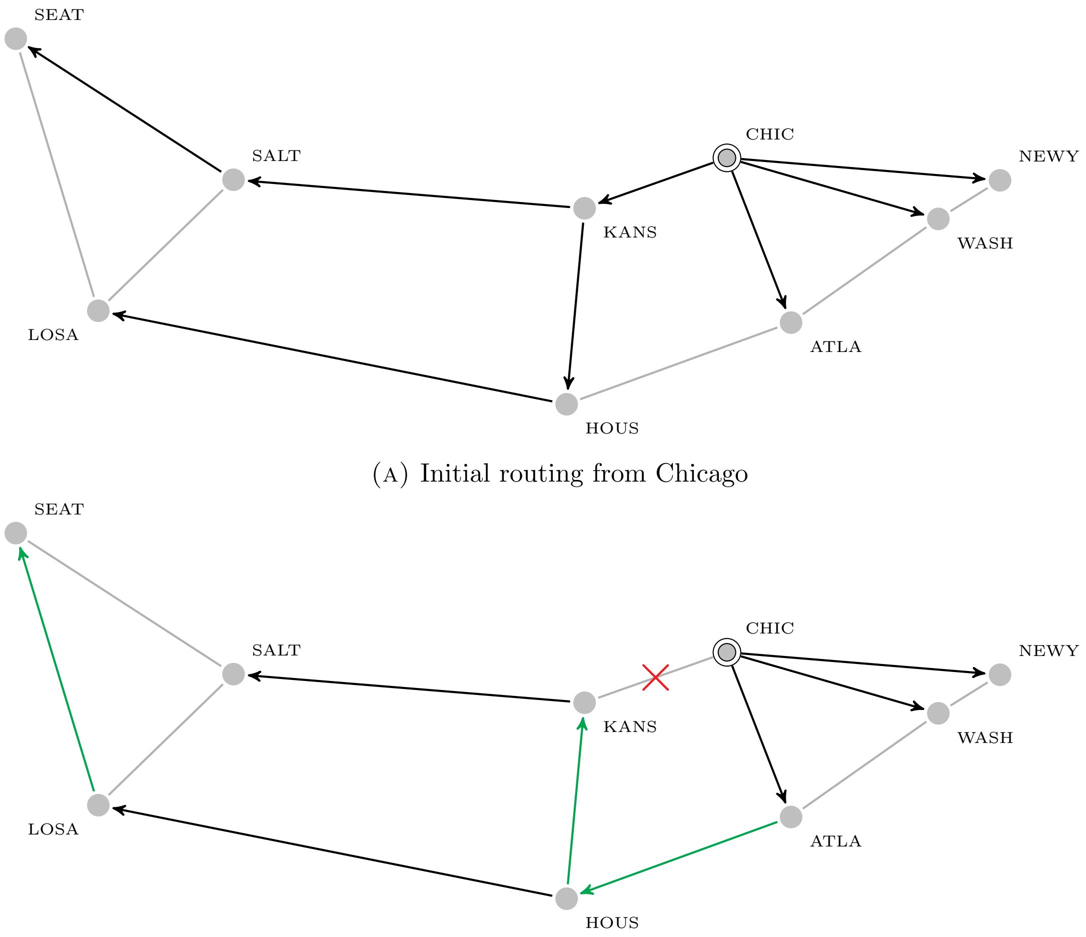


Figure 1.8: Shortest Path Tree rooted at Atlanta

Destination	Next-hop	Distance
SEAT	HOUS	4432
LOSA	HOUS	3090
SALT	HOUS	3532
HOUS	HOUS	1385
KANS	HOUS	2203
CHIC	CHIC	1045
ATLA	–	0
WASH	WASH	699
NEWY	WASH	976

Table 1.3: Routing table computed by the router at Atlanta

has to be calculated. Since Kansas City can no longer be directly reached from Chicago, the new path from Atlanta to Seattle necessarily implies using Houston as next-hop. Then, traffic may either be sent back to Kansas City, as there is still a link between Houston and Kansas City, and forwarded along the initial best path via Salt Lake City, or continue to Los Angeles. Finally, the traffic may reach Seattle either directly from Los Angeles, or via Salt Lake City. In order to select the new best path, we have to compare the respective distances associated to these paths. Passing through Houston, Kansas City and Salt Lake city puts Seattle at a distance of 4445 from Atlanta, while the path through Houston and Los Angeles is only 4432 units long. Intuitively enough, making a detour via Salt Lake City rather than using the direct link from Los Angeles to Seattle further increases the total distance. Routing paths towards Kansas City and Salt Lake City, which contained that same link, are also modified. Eventually, the router at Atlanta only forwards on its link to Chicago the traffic directly headed there. The new SPT and corresponding routing table are presented on Fig 1.8 and Table 1.3, respectively, with the modification compared to the initial routing plan appearing in green.



(A) Initial routing from Chicago

(B) Routing modifications caused by link (CHIC, KANS) failure

FIGURE 1.9: Shortest Path Trees rooted at Chicago

We now change our perspective and consider this same event from the router located at Chicago. As shown on Fig. 1.9a, this router initially forwards traffic along its direct link to Kansas City for most of the destinations in the network. Should it break down, all this traffic would have to be rerouted through another link. Aside from Kansas City, every route heading to the west cost necessarily goes through Atlanta. Besides, the distance using the direct link from Chicago to Atlanta is shorter than the one via Washington, a fortiori via New York. The router at Atlanta would thus be the new next-hop for these destinations. Modifications on the shortest paths from Chicago and their effects on the routing table are reported on Fig. 1.9b and Table 1.4.

By now, one may have noticed that a routing inconsistency could appear as the network converges to a new routing plan. Three destinations, that are Kansas City, Salt Lake City and Seattle, are initially reached from Atlanta through Chicago, while the contrary holds after the link modification. That is, Chicago reaches them via Atlanta. The router at Chicago being closer to the modified link, it will likely be aware of the change before the one at Atlanta. It would then recompute the shortest paths and update

Destination	Next-hop	Distance
SEAT	KANS ⇒ ATLA	2931 ⇒ 5477
LOSA	KANS ⇒ ATLA	3231 ⇒ 4135
SALT	KANS ⇒ ATLA	2018 ⇒ 4577
HOUS	KANS ⇒ ATLA	1507 ⇒ 2430
KANS	KANS ⇒ ATLA	689 ⇒ 3248
CHIC	–	0
ATLA	ATLA	1045
WASH	WASH	907
NEWY	NEWY	1000

TABLE 1.4: Routing table computed by the router at Chicago

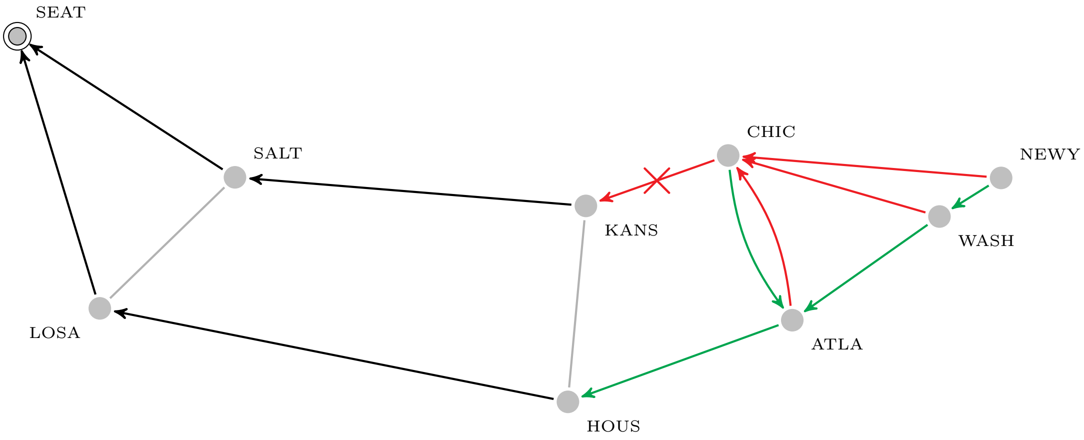


FIGURE 1.10: Merged Reverse Shortest Path Tree towards Seattle

its forwarding information while the router at Atlanta still forwards traffic along the initial routing plan, thus forming a transient routing loop between Chicago and Atlanta. New data packets arriving at Chicago and headed to any of these three destinations are transmitted to Atlanta, and from Atlanta back to Chicago. This phenomenon appears clear on a representation of the union of the shortest paths *towards* Seattle before and after the modification, as in Fig. 1.10. Black arrows represent the links that are used to reach Seattle both before and after the topological change. Red arrows are initial links towards this destination that are no longer used after the link (CHIC, KANS) becomes unavailable, whereas green ones are new links, only used after the change. The transient routing loop appears on this graph as a cycle between the routers at Chicago and Atlanta.

A graph of the shortest paths towards a given node from all other nodes is denoted as a Reverse Shortest Path Tree (RSPT), or a Reverse Shortest Path DAG (RSPDAG) if we consider the use of ECMP. In a networking context, such a graph represents the paths that are used by every router to reach a particular destination. By definition, this graph contains no cycle in a stable situation, when the shortest paths from all routers are calculated based on consistent network maps. However, this is not the case during a convergence period, and the paths used by each router depends on whether this router has already updated its forwarding decision or not. When a router detects a change in the network, it starts recalculating its shortest paths and, at the same time, sends to its neighbors link-state messages containing the updated information. These neighbors will read the content of the message, see that it contains new information and forward it to their own neighbors (except the one they received it from). Signalization messages are thus flooded within the network, gradually spreading this new information to every router. These messages do not necessarily follow the shortest paths calculated by routers, but are still subject to propagation delays and may be slowed down due to congestion issues. Last but not least, the time required for a router to process new link-state information and modify its forwarding decisions may vary, depending on its internal processing capabilities and its current load. Even though routers closer to a topological change are more likely to update first, it is thus commonly assumed that routers update order is not controlled. When evaluating the routing states in a network during a convergence period, we have to consider that each router may either be in its initial state, based on the network map before the topological change, or in its final state, having taken the new link-state information into account. In order to detect potential routing inconsistencies that could occur for a given destination, one should thus compute the shortest paths towards this destination that all routers may use before and after the change. We represent this transient routing state as a merged RSPT, or a merged RSPDAG if we consider the use of ECMP. Each cycle appearing on this graph denotes a potential transient routing loop that could occur during network convergence.

3.2 Evaluation of routing loops on a real ISP network

We recently started a collaboration with RENATER [REN], the French ISP for education and research.The first objectives of our collaboration are to measure the frequency and the impact of transient loops on a real ISP network, while long-term goals include the evaluation of the benefits of our solutions for preventing these loops.

The national network infrastructure of RENATER includes 72 routers providing Internet access to most academic institutions in France. Some of these institutions participates in the PlanetLab project[1] and, as such, maintain server-class machines, called *PlanetLab nodes*, that we can use to perform network measurements. Yet these nodes do not provide a sufficient coverage of the network to efficiently detect transient routing loops. Besides, this platform suffers from several virtualization limitations preventing from a fined-grained networking control. To start our collaboration, we thus deployed 10 Raspberry Pi devices to complement the PlanetLab nodes already in place (Fig. 1.11).

[1] http://www.planet-lab.eu/

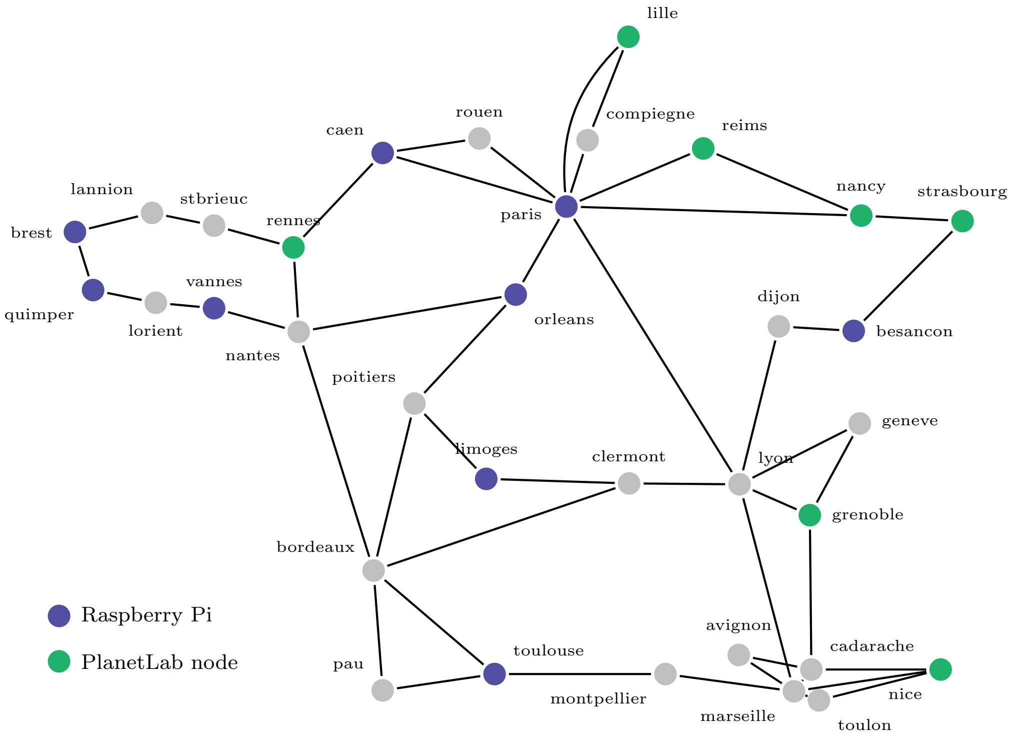
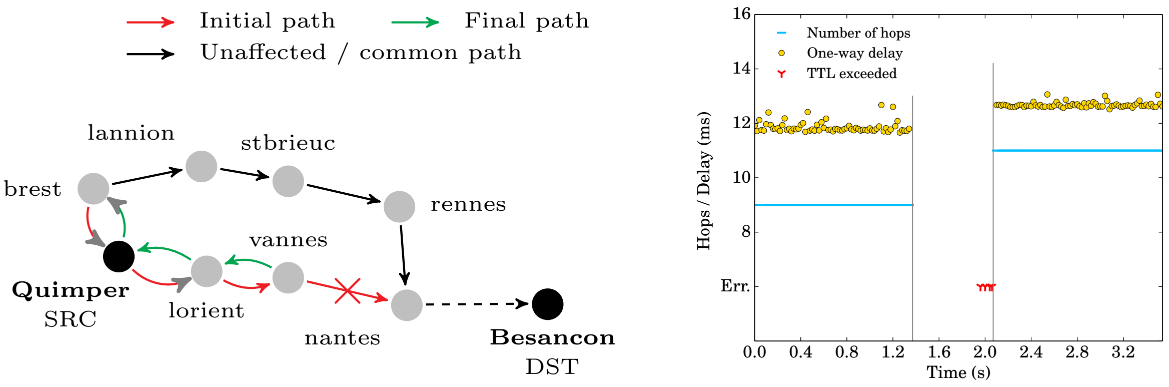


Figure 1.11: Measurement infrastructure on RENATER national network

These devices are directly connected to the routers in order to allow for more accurate measurements. In addition, we deploy an IS-IS-capable equipment in the network to establish an adjacency relationship with a RENATER router located in Paris in order to record every topological modification that has an impact on the routing protocol. This *listener* is based on a program developed by Richard Mortier for a similar study on the Sprint network [HMMD02].

As a preliminary campaign, we performed a first series of active measurements using only our 10 Raspberry Pi vantage points for a period of 21 days, from June, 6^{th} to June, 27^{th} 2014. Over this period, thanks to the listener, we detected 1371 topological modifications in the network, represented by unexpected LSAs modifying the LSDB of the listener. In average, more than 63 logical events have thus occurred every day (or 2.6 every hour). This frequency may seem very large, but does not necessarily reflects the physical events. For example, if a physical link is removed between two routers, the adjacency breaks down in both direction and triggering the transmission of two LSAs, one from each router. This is considered as two logical events, even though only one physical event occurred. In addition, note that routing events are not uniformly distributed across RENATER routers but rather follow a power law distribution according to their locations.

In the meantime, our 10 vantage points were exchanging ICMP messages at high frequency in order to provide accurate results as for the occurrence and duration of transient disruptions. Each vantage point was configured to transmit a message to every other with a period of 10 ms, while storing time and TTL information about all incoming and outgoing message. Note that each probe injected in the network is analyzed in a directed fashion, i.e., we do not rely on round trip measurements. This allows us to measure the variations in terms of number of hops on a given path and the one-way delay between the vantage points provided that the NTP synchronization provided by the RENATER server is sufficiently accurate. Each direction can be studied independently to finely understand the behavior on each path. This process allowed us to detect numerous transient loops in spite of the short duration of our experimentation. We present in the following two of the most interesting and typical cases we extracted from our data.



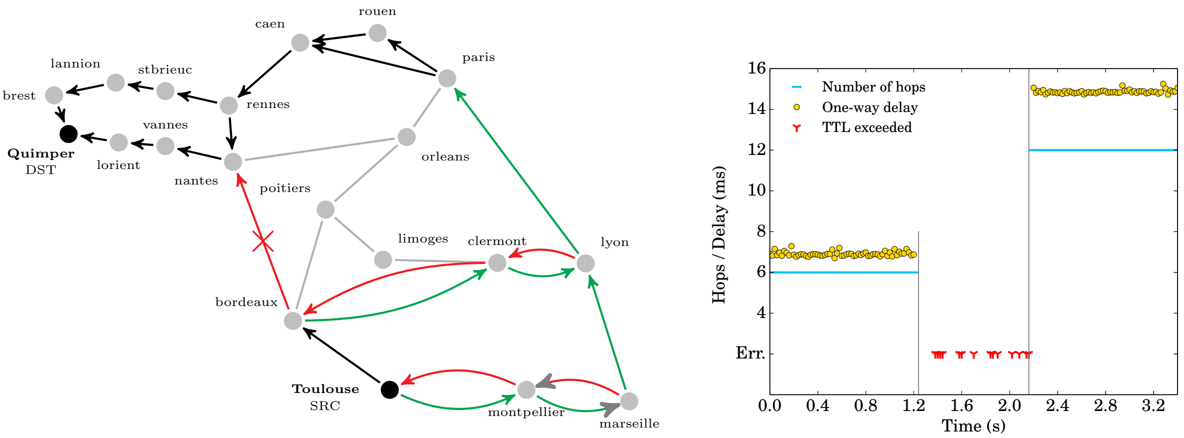
(A) Shortest paths to Besancon

(B) Traffic disruptions between Quimper and Besancon

FIGURE 1.12: Loops from Quimper to Besancon after the removal of link (Vannes, Nantes)

On Fig 1.12 we represent a transient loop that disrupted the traffic flows from Quimper to Besancon after the link between Vannes and Nantes went down. Initially, the traffic from our vantage point at Quimper was sent through Lorient, Vannes and Nantes (Fig. 1.12a), traversing a total of 7 routers in about 12 ms on its path to Besancon (Fig. 1.12b). However, when the router at Vannes lost its adjacency with Nantes, the traffic delivery was interrupted for about 720 ms, before a new path, through Brest, Lannion, Saint-Brieuc and Rennes was available. During the first 600 ms of the interruption, the packets were lost without our vantage points being notified of any event through ICMP error messages. Though it is not possible with limited information to be positive about the exact course of events, we can still make a conjecture. This period is likely to depict the time before a new route to Besancon was available on the router at Vannes. Yet, due to the low rate limit on *Destination Unreachable* ICMP messages (one message per 500 ms by default on Cisco routers), we cannot distinguish the failure detection delay from the computing time of the new route. After this period of silent losses, *TTL exceeded* ICMP error messages were received successively from the routers at Lorient, Quimper and Brest. Note that experiments we conduct on a small Cisco platform shows us that the rate limit on *TTL exceeded* ICMP error messages is about 20 messages every 500 ms.

The exact interfaces where the messages were received from are represented by gray arrows on Fig. 1.12a. This indicates at least two transient loops. A first one must have occurred between the routers at Lorient and Quimper, the former being up-to-date and forwarding the packets along its new path to Besancon, while the latter was not. Once the router at Quimper had updated its FIB, the loop was shifted to Brest, and packet losses continued until this last router had removed Quimper from its list of next-hops for Besancon. Then, the traffic delivery returns to normal, with a slightly larger delay caused by the longer path. To summarize, this first example shows two consecutive loops whose cumulated duration is greater than or equal to 120 *ms* although the traffic is disrupted during 720 *ms*.



(A) Shortest paths to Quimper (B) Traffic disruptions between Toulouse and Quimper

FIGURE 1.13: Loops from Toulouse to Quimper after the removal of link (Bordeaux, Nantes)

Fig. 1.13 shows a more complex routing loop scenario that occurred between Toulouse and Quimper. In this case, the traffic flow initially reaching Quimper through Bordeaux, Nantes, Vannes and Lorient, was disrupted after the removal of the adjacency from Bordeaux to Nantes. After this link failure, one can observe that there exists in theory up to 8 equal cost paths between the source and the destination. The silent loss period here barely exceed 100 *ms* (Fig. 1.13b), which tends to indicate that the router at Bordeaux detected the change and spread the information to the rest of the network within a very short delay. We can deduce from the collected *TTL exceeded* messages, received only from Montpellier and Marseille (Fig 1.13a), that the routers at Toulouse and Montpellier quickly updated their routes. The router at Marseille was however not so fast, and started forwarding its traffic along the new route almost one second later. Then, the traffic can reach the destination again, but the hop distance and the delay from Toulouse to Quimper are twice as large as before.

Besides, we can observe that the router at Toulouse only forwards the packets generated by our vantage point to Montpellier, despite the presence of an equal-cost alternate path through Bordeaux. Otherwise, some packets would have eventually reached Quimper or triggered *TTL exceeded messages* from Bordeaux, Clermont-Ferrand or Lyon.

These results not only indicate that transient loops actually occur, but also that they can cause non-negligible disruptions in traffic delivery, and jeopardize the respect of Service Level Agreements (SLAs). Generally speaking, we can conclude that a routing event such as a failure leads to significant disruptions whose durations are on the order of one second. Depending on the nature of the event (planned or not, adding or removing a component) and detection/processing delays, the shares between the black hole period and the loop one vary. For example, if the operator forces the modified component to flood a pseudo-infinte weight before a link removal (a common best practice), then the disruption duration is only made of the loop period. In the next-section, we present some of the solutions that have been developed to tackle disruption problems due to changes in the network.

4 Towards loop-free convergence

Over the past decade, several solutions [SB10a] have been proposed to address the problem of transient forwarding loops occurring during the convergence of link-state routing protocols. These solutions aim at ensuring loop-free convergence from an initial topology to a target one, assuming that both are entirely available during the whole process. Therefore, they are well suited to handle maintenance events such as weight reconfigurations or scheduled link or router state modifications. In case of failure, they could theoretically be combined with fast-reroute mechanisms (see section 2.2) to provide *smooth* transition from repair paths to the new topology. In such case, the failed component is still considered to be available, and traffic that should be sent through is forwarded along a repair path instead. However, the practical interactions between transient loop prevention and fast-reroute techniques are yet to be evaluated.

The solutions we describe in this section are extensions to routing and forwarding procedures that could be used for link and node-wide modifications. Some of them also apply to the more general case of a Shared Risk Link Group (SRLG), i.e. a set of seemingly independent links at the routing layer, but sharing properties at the physical or link layer.

4.1 Mitigating the effects of transient loops

Loop mitigation techniques aim at reducing the potentiality and effects of transient loops occurring in the network after a topological modification. They cannot guarantee that no loop would appear, but may reduce negative impacts on the traffic passing through the network. Since transient loops may only arise during a convergence period, speeding up the network convergence would be a straightforward solution to reduce the duration of loops, and thus mitigate their impact. Another possibility is to rely on a dedicated mechanism, such as Path Locking via Safe Neighbors (PLSN) [Zin05] or *local delay* [LDFF14].

Path Locking via Safe Neighbors (PLSN)

The PLSN method proposed in [Zin05] defines a *general safety condition* for each neighbor N of a given router S that has just been notified of a topological change in the network. In a symmetrically weighted network, N is a *safe* neighbor of S for a destination D if and only if both following statements hold:

1. N is a loop-free neighbor of S *before* the change, and

2. N is a downstream neighbor of S *after* the change

The first statement is similar to the LFA condition defined by equation 1.1 and simply ensures that the neighbor N does not use S to reach D in the initial topology. The second statement is more restrictive, as it requires that the distance from N to D in the new topology be strictly lower than the one from S. This is to ensure that S and N do not consider each other as safe neighbors. In an asymmetric network, N is a safe neighbor S only if this second criterion also holds before the change.

Each router classifies the destinations based on the safety degree of its neighbors into three categories:

- Type A: Destinations for which switching to the new primary next-hops cannot lead to a transient loop. Such destinations may either be completely unaffected by the change (type A1) or be reached through *safe* next-hops with respect to the above criteria.
- Type B: Destinations for which the new primary next-hops are *not safe*, but at least one other neighbor meets the safety condition.
- Type C: All remaining destinations.

The FIB is then updated according to the category of each destination. The routes for type A destinations are immediately modified to use the new next-hops, and FIB entries corresponding to type B destinations are updated to temporarily forward the traffic through their safe neighbors. In the meantime, traffic towards type C destination continues to be sent through the initial next-hops. These entries are only updated once all routers have changed their routes for type A and B destinations, using respectively new and temporary next-hops. Finally, the entries corresponding to type B destinations are updated to send the traffic to their new next-hops.

PLSN applies identically for any kind of topological modification, be it a single link reconfiguration, a router shutdown or startup, or even an SRLG failure. By ordering FIB entries updates and forcing routers to use safe, though not necessarily optimal, next-hops, this technique makes transient routing loops less likely to occur during network convergences. However, some loops may still arise when routers update their entries for type B and C destinations.

Local convergence delay

Litkowski et al. recently proposed in [LDFF14] a new method for transient loop mitigation in case of single link reconfigurations. This method is based on the assumption that most of the transient loops arising during network convergence are local to the modified component. Such local loops may arise after a link is shut down in one direction if the link source node updates its FIB, and starts using its new next-hops, before its neighbors do. This is likely to happen, as the source node notices the modification first. On the contrary, a local loop could occur when a new link is added if the neighbors update their FIBs before the link source node.

The solution presented in [LDFF14] aims at preventing these loops by introducing a convergence delay between the router detecting the event and the rest of the network. For link "down" events, the local node normally advertises its neighbors of the event, but a *positive delay* makes it wait before updating its own FIB. The neighbors would thus converge to the new forwarding state before the local node, so that no transient loop may occur when the local finally updates its routes. This behavior is reversed link "up" events. The local node immediately starts its convergence process, and delays the flooding of the new LSA (*negative delay*) to ensure that its neighbors do not converge first.

This transient loop mitigation technique is often referred to as *local delay*. Simulations show that implementing this method may prevent more than half the potential transient loops in case of link down events, and up to 80% in some topologies. *Local delay* can also be extended to prevent local loops for router-wide modifications and SRLG failures.

4.2 Preventing the effects of transient loops

The following methods completely negate the impact of transient routing loops by providing safe forwarding paths during the network convergence. These mechanisms can be classified into three categories: tunnel-based solutions, packet marking, and FIB update control.

Tunneling

Tunnel-based methods [SB10a] require that routers be notified about the topological modification before the convergence begins. Each router may thus determine which destinations are affected by the change and build tunnels to forward the corresponding traffic along static, safe paths instead. When the modification is actually advertised in the network, all routers may normally converge to their new routing state, without the traffic being disrupted. These methods do not prevent routing tables to be temporarily inconsistent but, since any packet that could potentially have been affected is carried within a safe tunnel, no transient loop may occur. Once all routers have converged, these loop preventing tunnels are removed and the normal forwarding resumes for all destinations. It is worth noting that the order in which tunnels are established and withdrawn is not important as long as they are in place for the duration of the convergence.

Different variations exist as for how safe tunnels are to be established. Nearside tunnels use normal routing to carry the traffic to the closest router adjacent of the failure, where it is forwarded through the modified link, or through a repair path in case of a failure. Each router hence only need to compute a single tunnel for all affected destinations. On the other hand, farside tunnels bring affected traffic to the far side of the modified component. These tunnels cannot use normal routing, since this would mean passing through the modified component, and must rely on repair paths provided by mechanisms such as Not-Via (see section 2.2). While this may appear as a drawback, using repair paths actually lighten the load on routers affected by the change, allowing for a more uniform distribution of the tunneled traffic. Besides, in the case of a node-wide modification, the decapsulation load is shared by the neighbors of the modified router rather than being held by this router alone. The last variation, denoted distributed tunnels, relies on repair paths computed by each router. The traffic headed to destinations that are affected by the change is tunneled along a repair path to an unaffected router, from where it is normally forwarded to the destination.

Packet marking

Packet marking [SB10a] is one of the most straightforward mechanisms to prevent transient loop. During the convergence period, packets transiting in the network are marked by the first router they cross and assigned to either the old or the new topology. Other routers are then forces to forwarded these packets along the topology they are marked for.This method however requires that a marking bit is available, for example in the Type of Service (TOS) field of IP packets, and that each router maintains two concurrent FIBs for the old and new forwarding states.

FIB update control

The methods described above do not really prevent transient loops, but rather their effects on the traffic. Should one consider only the normal forwarding state of the network, transient loops may still appear during the convergence. What prevents the packets from looping is specific processing performed by the routers, which does not, or not only, involve the normal FIB. Such special treatment however demands more computing capabilities from the routers and may slow down traffic delivery. It would thus be more efficient to only rely on normal forwarding, and control the update process to ensure that no transient loop could arise.

The ordered FIB (oFIB) approach, originally presented by Francois and Bonaventure in [FB05] and later developed in [FB07, SBP+13], relies on an analysis of the network topology to provide a FIB update ordering that provably prevents transient loops. In the case of *down* event, that is a link or router being shut down, or a link weight being increased, oFIB ensures that any router R updates its FIB only after all routers sending traffic via R and the modified component did. Providing that all routers affected by the change are oFIB-capable, packets sent by a router that has not yet updated its FIB are thus forwarded to the destination along initial paths only, and cannot loop. Also, the traffic sent by updated routers may may either reach the destination using only final paths, or (in asymmetrically weighted networks) pass through a router that is not up-to-date at some point and be forwarded according to the initial topology the rest of the way. In both cases, packets will reach the destination without being caught in a loop. In the case of an *up* event, the condition is reversed, forcing a router R to update its FIB before any other router that may use R to reach the modified component.

This transient loop-free FIB update ordering is ensured the following way. Upon being notified by IGP signalization of a *down* even, a router R computes an RSPDAG rooted at the modified component and based on the initial topology. Router R then waits for a duration that is a multiple of its rank in the RSPDAG before updating its FIB, where the rank is defined as the maximum length of a branch heading to R. Similarly, if R is notified of an *up* event, it computes an RSPDAG routed at the modified component in the new topology and wait for a multiple of its rank. In this second case, the rank is equal to the maximum length of a path from R to the modified component.

Due to the use of conservative timers, the process described above can be quite slow. It is possible to accelerate the convergence by replacing timers with completion messages. According to the RSPDAG, each router R computes the list of neighbors it must wait for before updating its FIB. Upon receiving a completion message from one of these neighbors, R removes the neighbor from the list. Once the waiting list is empty, R updates its FIB and sends a completion message to the neighbors that are waiting for it.

Ships-in-the-Night

Ships-in-the-Night (SITN) [VVP+12] is a recent technique designed for network-wide migration of link-state IGPs. Compared to other loop prevention techniques, SITN allow for simultaneous modification of multiple components all around the network, and even the replacement of one IGP with another. The technique relies on two concurrent routing processes, corresponding to the initial and final IGP configurations, running at the same time on each router. A priority system is used to determine which route is to be installed in the FIB. Initially, the lowest possible priority is assigned to the final IGP configuration, ensuring that no route from this process is installed in the FIB. Once the final IGP configuration has converged on every router in the network, those are progressively migrated according to a pre-computed, loop-free order. Eventually, the initial IGP configuration can be removed from all routers with no impact on the network.

The authors prove that the problem of finding a router migration ordering that prevents transient loops is *NP*-complete in the general case. However, an efficient and correct heuristic algorithmic would consist in computing sufficient ordering constraints separately for each destination. These constraints could then be considered together in order to determine a global ordering.

5 Metric-increment approach

5.1 Presentation

In the previous section, we presented efficient solutions to deal with the problem of transient forwarding loops. However, they all rely on non-standard behavior of IGPs, thus requiring extensions to protocol specifications. As a result, none of these transient loop prevention mechanisms is currently available for network operators, nor is likely to be in the near future.

An alternative approach to prevent transient loops in the case of a single link modification, using only the core functionality of link-state routing protocols, has been proposed in [IIOY03] and [FSB07]. This approach, often referred to as *metric-increment*, is based on the idea of controlling FIB updates in an implicit fashion, as opposed oFIB explicit ordering. It relies on successive link weight reconfigurations to progressively adjust the attractiveness of the modified link. It is apparent enough that increasing (resp. decreasing) the weight configured on a link makes it more (resp. less) likely to be used to forward traffic, however careful weight tweaking allows for a fine control of routers update process. An interesting property of shortest path routing is that nodes further away from the link are more sensitive than closer ones to weight increments, while the opposite holds for weight decrements. It is thus possible to compute a sequence of link weight updates that forces some routers to update first, while others still follow the initial routing plan.

Practically, the approach requires to model link addition or withdrawal operations as IGP weight reconfigurations. A link removal is thus considered as an increment from its current weight to the maximum possible one, known as *MAX_METRIC*, in both directions. Assuming that the network is 2-edges-connected, this operation will result in the link not being used anymore to reach any destination in the network. It may then be safely removed with no impact on the routing decisions. Similarly, the addition of a new link could be done by first adding it in the network with a weight of *MAX_METRIC*, and then decreasing it to the one specified by the operator. The approach then consists in splitting this weight modification, which could cause transient routing loops, into a sequence of *safe* weight updates. Intermediate updates are computed such that no transient loop could appear as they are applied, provided that two consecutive updates are separated by a *sufficient* amount of time. In the case of a single link whose weight is to be modified in both directions, two different sequences are required. However, both sequences may be computed independently and applied at the same time.

Also note that loop-free sequences are reversible. An update sequence preventing transient loops during the convergence for the reconfiguration of a link weight from X to Y can be applied in reverse order for the reconfiguration from Y to X. We thus often focus on weight increment operations, without loss of generality.

As an illustration, consider a scheduled withdrawal of the link from Chicago to Kansas City on Internet2 network (Fig.1.14). In section 3.1, we showed that a direct removal, or a weight increment to *MAX_METRIC*, leads to a potential transient loop between Atlanta and Chicago. In particular, if the router at Chicago updates its FIB before the one at Atlanta, it will start forwarding its traffic towards Seattle on its link to Atlanta, while the router at Atlanta still reaches Seattle through Chicago.

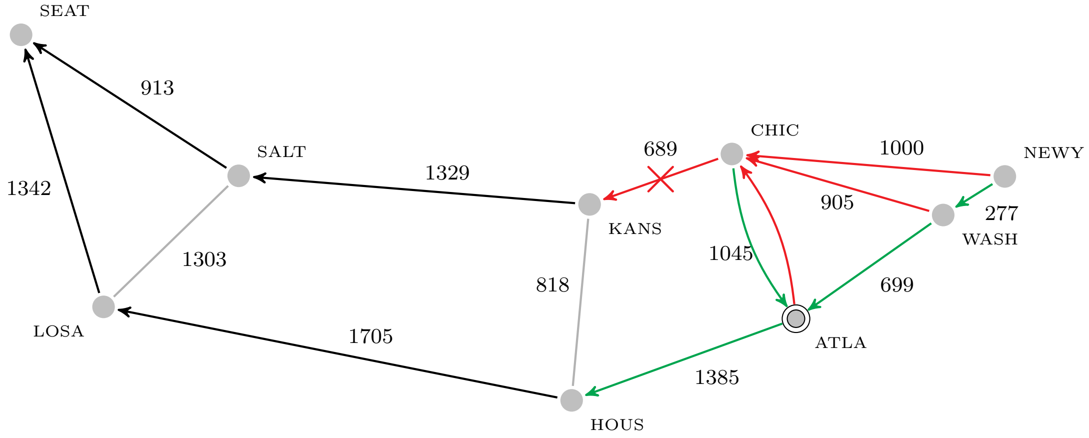


FIGURE 1.14: Merged RSPDAG towards Seattle for the removal of link (CHIC,KANS)

Using the metric-increment approach, this transient loop can be prevented by inserting an intermediate state before configuring the weight on link (CHIC, KANS) to *MAX_METRIC*.

On Fig. 1.15, we consider a first increment of 1000 that brings the weight on this link from 689 to 1689. From the perspective of the router at Atlanta, Seattle is now at a distance of 4976 through Chicago, which is more than the distance via Houston. The routing protocol running on the router reacts to this change by recomputing its shortest paths and replaces Chicago with Houston in the FIB entry for Seattle. The other routers at Chicago, New York City and Washington also recompute their routing tables, but the paths going through link (Chic, Kans) remain the shortest ones after the change, so that no modification is pushed to the FIB. The router at Atlanta being the only one to change its routing decision for Seattle, no transient loop can occur during the transition.

Then, when the link weight is configured to *MAX_METRIC*, the router at Chicago can safely reroute to Atlanta, as no transient loop could occur during the transition (see Fig. 1.16). The potential transient loop between Chicago and Atlanta has been prevented

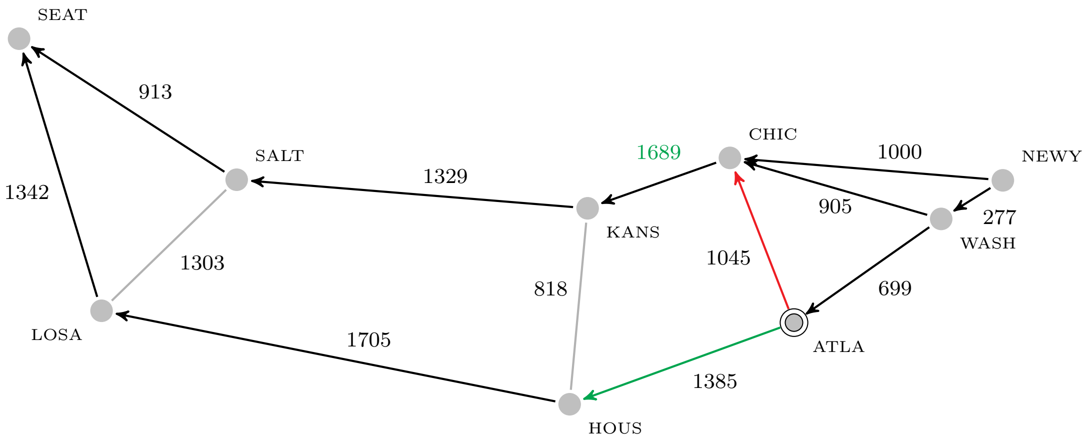


FIGURE 1.15: Merged RSPDAG for the weight increment from 689 to 1689 on (CHIC,KANS)

by making the router at Atlanta update its FIB at a previous step, so that it no longer sent its traffic toward Seattle via Chicago when the target weight was configured.

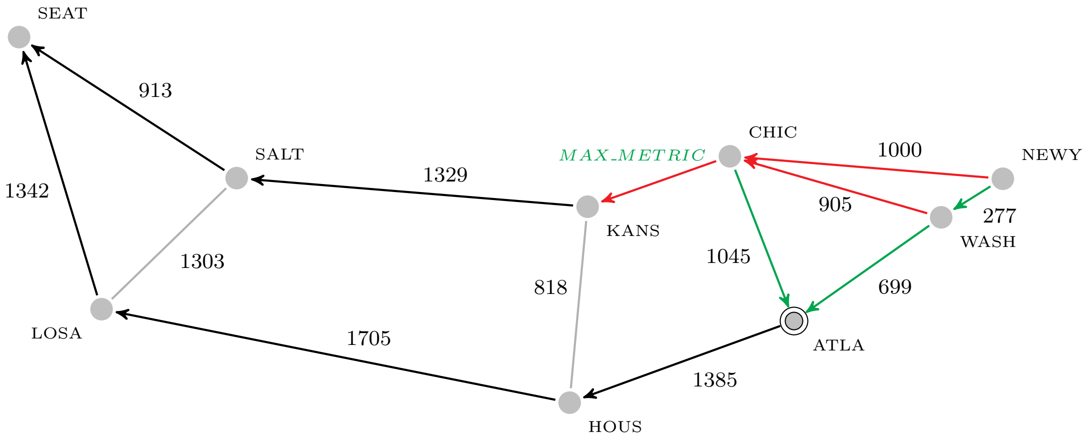


FIGURE 1.16: Merged RSPDAG for the weight increment from 1689 to *MAX_METRIC*

Table 1.5 shows the state of the routing table entries for Seattle at each step of the convergence. In the initial state (1.5a), the router at Chicago uses Kansas City as it next-hop towards Seattle, while the ones at Atlanta, Washington and New York City go via Chicago. When the weight of link (CHIC,KANS) is first modified (1.5b), the next-hop at Atlanta is modified and the distance is updated accordingly. The weight increment is also reflected on the three other routers using the link, but it is not sufficient to modify their routing decision. Their next-hop is only modified when *MAX_METRIC* is configured on the link (1.5c).

5.2 Loop-free update sequences

In [FSB07], the authors prove that increasing a link weight by 1 cannot lead to a transient loop during the convergence. It is thus theoretically possible to *safely* perform any weight

Source	Next-hop	Distance	Next-hop	Distance	Next-hop	Distance
SEAT	–	0	–	0	–	0
LOSA	SEAT	1342	SEAT	1342	SEAT	1342
SALT	SEAT	913	SEAT	913	SEAT	913
HOUS	LOSA	3047	LOSA	3047	LOSA	3047
KANS	SALT	2242	SALT	2242	SALT	2242
CHIC	KANS	2931	KANS	3931	ATLA	5477
ATLA	CHIC	3976	HOUS	4432	HOUS	4432
WASH	CHIC	3836	CHIC	4836	ATLA	6176
NEWY	CHIC	3931	CHIC	4931	WASH	6453

(A) Initial state (B) Intermediate state (C) Final state

TABLE 1.5: Routing table entries of each router towards Seattle

modification as a succession of +1 or −1 operations, provided that two consecutive operations are separated by a sufficient amount of time. Such solution is however not realistic in practice, as it would require far too much time in most cases.

The authors thus present an algorithm to compute short loop-free weight update sequences for any single link removal operation. A sequence is called loop-free if no transient loop may arise during the convergence between two subsequent updates. A loop-free sequence relies on special weight values, denotes as *key metrics*, which represent the minimum weight to be configured on the modified link in order to force a node into using a *new path* that does not pass through the modified link for a given destination. Such key metrics are computed by comparing the shortest path distance between a node and the destination in the initial and final topologies. Formally, given a modified link (a, b) and a destination d, the key metric associated to node x is equal to

$$w_0(a, b) + C'(x, d) - C(x, d)$$

where $w_0(a, b)$ represents the initial weight configured on link (a, b), and $C(x, d)$ and $C'(x, d)$ respectively represent the shortest path distance between x and d in the initial and final topologies.

Bringing together in a sorted sequence the key metrics of all nodes in the topology is however not sufficient to ensure a loop-free convergence. In addition, a *reroute metric sequence* includes *intermediate metrics* that are equal to key metrics plus one. While key metrics are the lowest weights such that a node starts using a new path, intermediate ones represent the minimum weights such that the node stops using a path through the modified link. It has been proven that a *reroute metric sequence* computed for a given destination is always loop-free for this destination. However, such sequence may contain *unnecessary* metrics that are not required to prevent transient loops, and can be pruned from the sequence using a trial-and-error approach. Eventually, it is possible to obtain for each destination an *optimal reroute metric sequence* containing the least number of intermediate update to ensure a loop-free convergence.

In order to provide loop-freeness for every destination, *optimal reroute metric sequences* are merged into a *global metric sequence*. Since it can be proven that inserting intermediate updates in a loop-free sequence preserves the loop-freeness property, *global metric sequences*, defined as the sorted union of per-destination sequences, provably prevent transient loops for all destinations in the network. Global sequences may contain redundant values and can be reduced as well using the same procedure as for per-destination sequences. Finally, it is possible to obtain short *global metric sequences* that allow for a transient loop free convergence of the network.

5.3 Limitations

The metric-increment approach makes it possible to prevent transient loops for any single link modification. However, the algorithms presented in [FSB07] come with certain limitations. Among those, the time required to compute loop-free sequences grows rapidly with the size of the topology. This is mostly due to the reduction algorithm, which requires testing each intermediate metric in order to prune redundant ones. During the first reduction phase that is performed for every per-destination *reroute metric sequence*, an SPDAG representing the intermediate forwarding state is to be calculated for each element of the sequence and compared with the final state. As for the global reduction phase, each remaining increment is to be tested that way for every destination. Hence, the more destinations there is in the network, the more time is required to compute a sequence.

Also, while *optimized reroute metric sequences* are proved of minimal length, there is no guarantee that the global sequence, after the reduction stage, shares this property. Ironically enough, the reason why global sequences may not be of minimal length is because of the per-destination reduction. From a pure algorithmic point of view, this first sequence optimization is in fact unnecessary, as redundant metrics would be pruned by the global sequence reduction anyway. It is performed in practice to reduce the total time required to compute a sequence, metric pruning coming for cheaper when it only needs to be tested for one destination. It would thus be possible to ensure global minimality at the expense of longer computing times.

We developed this solution in [CMP$^+$14], providing a new algorithm to compute minimal global sequences. This algorithm is based on the notion of loop-free metric intervals, representing necessary and sufficient conditions to prevent transient loops. Even though enumerating each potential transient loop to obtain the associated interval is impractical in large networks, we devised efficient methods to extract enough information about these intervals and compute minimal sequences in reasonable time.

Another possible limitation of this approach is the risk of negative interactions with BGP. In [VVCB13], Vanbever et al. discuss the impact of graceful IGP operations on certain BGP configuration. Since BGP decisions are partly based on IGP routing, each IGP reconfiguration forces BGP enabled routers to recompute their best paths and possibly update their forwarding table entries. The authors show that, on networks using multiple layers of BGP route reflectors, such changes may also trigger BGP forwarding loops. Fortunately, sufficient conditions can be met to ensure that no such loop could occur. In particular, network-wide packet encapsulation, which is widely used in transit networks, provably prevents IGP reconfigurations from triggering BGP anomalies.

6 Conclusion

In this chapter, we presented the general context of our work. We first described the basis of IP routing and gradually focused on the problem of transient routing loops in network running link-state protocols. Based on measurement we conducted on a real ISP network, we showed that such loops could cause significant traffic disruptions. Several solutions have been proposed in the past to prevent the occurrence or the effects of transient loops, yet all but one require extensions the protocol specifications, hindering a practical deployment. On the other hand, the metric-increment approach entirely relies on basic functionalities of link-state routing, making it both practical and, in a sense, incrementally deployable on any ISP network. However, current algorithms are limited to reconfigurations of a single link. We generalize this approach in the next chapter, detailing how to efficiently compute minimal loop-free sequences for any router-wide operation.

Chapter 2

Algorithmic contributions

Contents

This chapter provides algorithms and formal proofs that demonstrate how our solutions are able to prevent any type of disruption during the convergence of link-state intra-domain routing protocols. These solutions are designed to help network administrators perform scheduled operations on their network, by preventing transient inconsistencies that impact the traffic. Our contributions extend the metric-increment approach described in the previous chapter with more efficient algorithms and additional use cases. In particular, our new algorithms provides minimal reconfiguration sequences for any modification on a single link, in both directions, but also on a whole router or a subset of its outgoing links. A single link operation may either consist in adding a new link to the network, removing or shutting down an existing one, or reconfiguring the Interior Gateway Protocol (IGP) weight associated to the link. As for router-wide operations, our approach supports the addition or withdrawal of a whole router, as well as any positive weight increment or decrement on a subset of its outgoing links.

It is considered best current practice that the removal of an entire router is only performed after having configured the weight on each transit link to *MAX_METRIC*. This results in the router not being used anymore as a transit node, while stub networks remain normally reachable. On networks running Intermediate System to Intermediate System (IS-IS) protocol, the same behavior can be obtained without modifying link weights, by setting the *overload bit* [McP02] in outgoing Link-State Advertisements (LSAs). This method avoids transient traffic black-holing that usually occur when a router is being abruptly shut down. Similarly, routers should be started up with the overload bit set, or *MAX_METRIC* configured on transit links, in order to prevent forwarding paths from being modified until the router is completely operational. Based on the assumption that these guidelines are respected, our algorithms compute sequences of vectorial weight updates preventing transient inconsistencies that could occur during the convergence from the initial to the final routing state.
Vectorial updates represent weight reconfigurations to be applied simultaneously on several outgoing links of the router. As a consequence, it is a requirement for our solution that routers software supports the simultaneous weight reconfiguration of multiple outgoing links, and effectively advertises such modification in a single LSA. This behavior varies among router operating systems as it depends on Open Shortest Path First (OSPF) and IS-IS implementations. To the best of our knowledge, simultaneous modifications are currently supported on Juniper's *Junos OS* and Cisco's *IOS XR*.

For the sake of simplicity, we focus in this chapter on the case of a router being shut down. Router additions can be simply performed by reversing the shutdown process, while operations on a subset of links only require to ignore unmodified links. Since we aim at rerouting the traffic out of the router, we do not consider reconfigurations that would make it, or a subset of its links, more attractive than they initially were.

Intermediate weight updates may thus not decrease a link weight below its initial value. To emphasize this aspect, we refer to weight reconfigurations as *increments*, even though we allow for weight decrements that satisfy the above property.

1 Weight increment basics

Using the weight increment approach, a first, naive way to safely shut down a router is to consider the problem as a set of single-link shutdown operations. One single-link increment sequence could be computed and applied for each outgoing link of the router. Such technique does not however benefit from the opportunity of simultaneously modifying the weight on multiple links. Moreover, since increasing the weight on one outgoing link of a router can make the traffic be rerouted through another outgoing link of the same router, the order in which link weights are increased may have an impact on the total number of intermediate steps. In other words, the weight increment sequence to be performed on one link could depend on the previously applied sequences. Our experiments show that this approach leads to very long sequences that are not realistic for practical use. We thus devised solutions specifically tuned for node-wide operations.

We present here the basics of our node-wide approach, explaining how it can be used to prevent transient forwarding loops in the case of a router shutdown operation, and discussing the main properties it relies on. We also show that, while it is necessary to achieve feasible sequence lengths, aiming for short update sequences through non-uniform weight modifications may lead to a different kind of transient disruptions.

1.1 Distance increments and uniform sequences

We consider the problem of shutting down a router in the network without incurring transient forwarding loops during the convergence. Using our approach, safely performing such operation requires to progressively increase the distance of passing through this router, in order to make it less and less attractive as a transit node. Contrary to links, routers have no internal weights, so that it is not possible to directly vary the distance via a given router. However, the same effect can be obtained by uniformly modifying the weights configured on each outgoing link of this router. For example, let us consider that the weights on all outgoing links of a router are increased by an arbitrary value u. The costs of all paths in the network are hence increased by this value for every occurrence of such a modified link. Since it is a property of shortest paths that each router can be traversed no more than once, and the weights are only increased in one direction, the distance of all shortest paths through this router is thus increased by exactly u. In this section, we rely on such uniform increments to explain how progressively increasing the distance through a router enables to prevent transient loops.

Definitions and Notations

In link-state IGPs, forwarding paths are computed based on a weighted graph $G = (N, E, w)$, such that N is the set of routers, E is the set of IGP adjacencies between routers, and $w : E \rightarrow \mathbb{N}$ maps each oriented link to its integer weight as defined by the IGP configuration. Note that we consider adjacency relationships to be symmetrical, i.e. if A is adjacent to B then B is adjacent to A, but a different weight may be associated with each direction. Also, point-to-multipoint adjacencies are represented as a collection of point-to-point links between adjacent devices.

$P(x, d)$ denotes the set of shortest paths linking node x to node d while $C(x, d)$ refers to the cost of paths in $P(x, d)$. $RSPDAG(d)$ is the Reverse Shortest Path DAG (RSPDAG) rooted at d, which contains the shortest paths towards d from all other nodes in N. Our theoretical framework is based on Directed Acyclic Graphs (DAGs), rather than trees, to support Equal-Cost Multi-Path (ECMP) routing. That is, simultaneously using for a single destination multiple shortest paths having the same cost. Although, for IP networks, destinations are prefixes in practice, which would correspond to the edges of our graph, we choose instead to consider the nodes in N as the destinations. Such representation is more intuitive, especially for people who are not familiar with IP routing, and remains realistic if we consider routing paths towards loopback addresses configured on each router. Besides, most of the traffic passing through an Internet Service Provider (ISP) network is actually headed to destinations outside of this network.

When the *distance* through a node in G is modified, we respectively denote the RSPDAGs of d before and after the change as $RSPDAG(d)$ and $RSPDAG'(d)$. We say that a change is *loopfree* for destination d if no forwarding loops can occur during the convergence triggered by the change, whatever the order of router updates. Such forwarding loops can be detected by merging $RSPDAG(d)$ with $RSPDAG'(d)$. If the merged graph $RSPDAG(d) \bigcup RSPDAG'(d)$ contains cycles, then forwarding loops may occur. More generally, we say that a change is *loopfree* if it is *loopfree* for all destinations in the network.

Considering a given destination $d \in N$, let $G_r(d) \subset RSPDAG(d) \subset G$ be the subgraph impacted by the distance modification on router r. Each node in $G_r(d)$ is affected by the change on r, either because its shortest paths towards d are modified to avoid r or because their cost is increased. Note that $G_r(d)$ forms an RSPDAG rooted at r, which contains all the paths in $RSPDAG(d)$ ending at r.

Since we often consider a single destination d, we simplify our notation when there is no possible ambiguity. We use $P(x)$ and $P'(x)$ to respectively refer to $P(x, d)$ and $P'(x, d)$, the set of shortest paths used by node x to reach destination d after the change. Similarly,

$G(N, E, w)$	Directed weighted graph
$G_r(d)$	Subgraph impacted by a distance increment through router r for destination d
λ	Smallest distance increment that can be applied
$RSPDAG(d)$	RSPDAG rooted at d before the change
$P(x, d)$, $P(x)$	Set of paths from x to d in $RSPDAG(d)$
$C(x, d)$, $C(x)$	Cost of the paths in $P(x, d)$
$RSPDAG'(d)$	RSPDAG rooted at d after the change
$P'(x, d)$, $P'(x)$	Set of paths from x to d in $RSPDAG'(d)$
$C'(x, d)$, $C'(x)$	Cost of the paths in $P'(x, d)$

Table 2.1: General notations

$C(x)$ and $C'(x)$ are simplified notations to denote their respective costs. Table 2.1 summarizes all the notations described above.

Illustration on a gadget

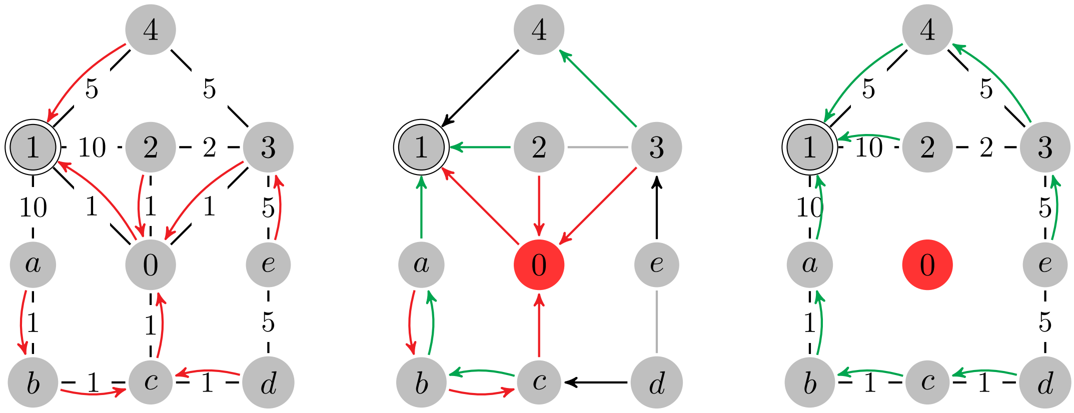
To illustrate how intermediate distance increments can prevent transient forwarding loops in the case of a router shutdown, let us consider the network represented in Fig. 2.1. For the sake of clarity, the nodes in this graph are labeled with a mixed set of digits and letters. Node 0 represents the router to be shutdown, nodes 1 to 4 are destinations for which transient loop occur during the operation, and nodes a to e are source nodes that could be involved in these loops. On the left-hand side figure (2.1a), we represent the initial forwarding paths towards destination 1, namely $RSPDAG(1)$. The target forwarding paths, that are used after router 0 is shut down, are represented on the right-hand side figure (2.1c). The central figure (2.1b) represents the merging of the initial and target paths. This figure shows that, if the shutdown operation is performed directly, whether or not a MAX_METRIC or overload bit mechanism is used, transient loops can potentially occur on links (a, b) and (b, c).

Let us now show that there exists a sequence of uniform distance increments for node 0 that enables to shut it down without incurring any transient loop. Consider that, at the first step, the distance through 0 is only increased by 7, i.e. the IGP weight configured on each outgoing link of 0 is increased from 1 to 8. After the modification, the distance for reaching 1 by traversing 0, which initially was equal to 2 whatever the source and destination (1 to go to 0 and 1 more to get out), becomes of 9. The cost to enter 0 is still 1, but exiting the router now has a cost of 8. Such increment may lead some of the routers whose shortest paths go through 0 to reconsider their routing decision. In particular, the distance from router a to 1 through 0 is now equal to 11, which is larger than the cost through the direct link $(a, 1)$. Router a thus updates its Forwarding

Information Base (FIB) and replaces b with 1 as its next-hop towards 1. The other routers also receive the LSA, but their initial paths for this destination remain the most attractive. The modifications this first increment brings to the forwarding paths towards 1 are shown on Fig 2.2a.

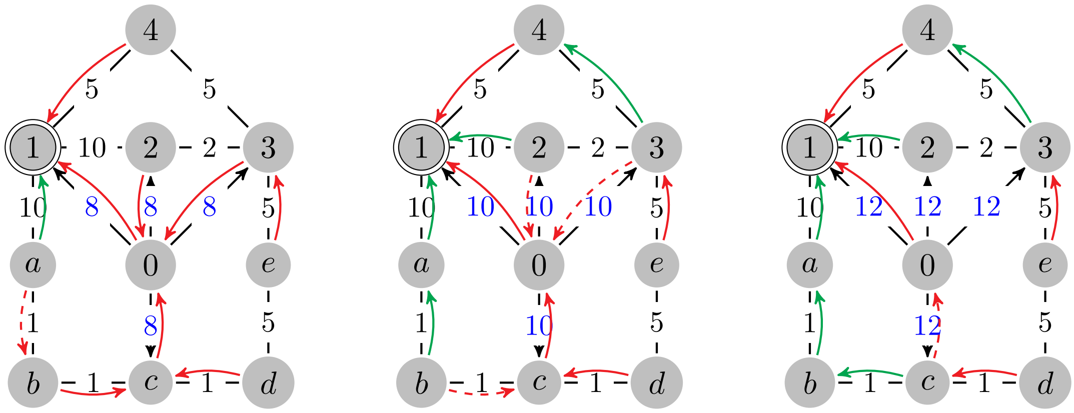
At the second step, we increase the distance by 9 compared to the initial state (plus two compared to the first step), making the cost for traversing 0 equal to 11. This change has no effect on the routing decisions of node a, for it was not using 0 anymore, but routers 2, 3 and b now have better paths not via 0 to reach 1, as represented in Fig 2.2b. Node 2 can use its direct link to 1 for a cost of 10, instead of 11 if through 0. Node 3 switches 0 with 4 in its FIB entry, and b reroutes its traffic for destination 1 through node a. Finally, a third increment of 11 makes node c reroute through b for a distance of 12, compared to 13 via 0. After this last step, no more node but itself uses 0 to reach destination 1, and the router can be safely shutdown.

Let us now focus on the potential transient loops that we mentioned before. According



(A) *RSPDAG*(1) (B) *RSPDAG*(1) ⋃ *RSPDAG'*(1) (C) *RSPDAG'*(1)

FIGURE 2.1: Forwarding paths towards destinations 1 before and after the removal of node 0.



(A) Uniform increment of 7 (B) Uniform increment of 9 (C) Uniform increment of 11

FIGURE 2.2: Progressive increment of the distance through node 0

to the merged RSPDAG, one could have occurred on the link (a, b) in the case router b rerouted before a. However, if the distance through 0 is progressively incremented as we advised, router a stops forwarding its traffic via b at the first step, while b only starts using a at the second step. Consequently, if a sufficient delay separates the execution of these two steps to allow for router a to converge, this transient loop cannot occur. The same holds for the other loop on link (b, c), whose involved routers respectively reroute at the second and third step. The sequence $\{7, 9, 11\}$ thus prevents all possible transient loops that could arise for destination 1 when router 0 is being shut down. In the following, we refer to such a sequence as a *loopfree uniform distance increment sequence*. In practice, that last increment of 11 can be replaced with MAX_METRIC, or any other technique that would prevent 0 from being used for transit.

Just as for single link operations, similar sequences are to be computed for each destination in the network that is affected by the node shutdown operation. In our example, other affected destinations include routers 2, 3 and 4, whose associated loopfree distance increment sequences are represented on Fig. 2.3. Destinations 2 and 3 each require a single intermediate increment, which is respectively equal to 10 and 8, while the distance for destination 4 has to be successively increased by 7 and 8. In order to ensure global loopfree convergence, it is thus sufficient to merge together these destination-oriented sequences. Considering the network represented in Fig 2.1, the sequence $\{7, 8, 9, 10, 12\}$ thus prevents any transient forwarding loop that could occur due to the removal of router 0.

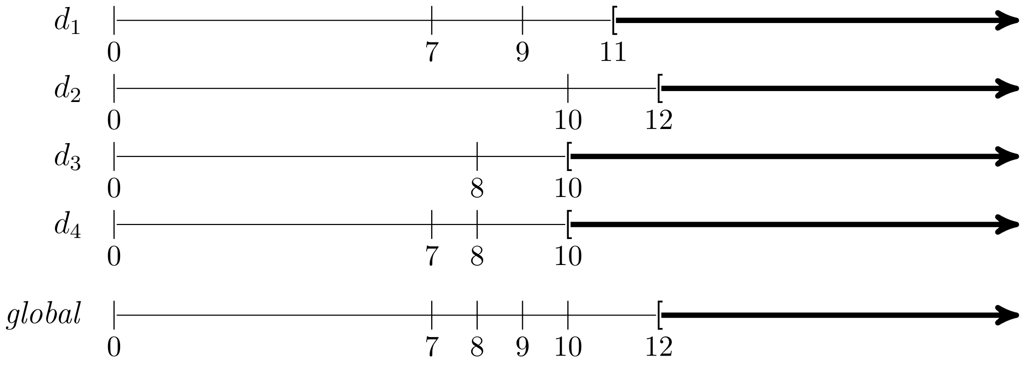


FIGURE 2.3: Destination-oriented and global distance increment sequences

Existence of loopfree distance increment sequences

To prove that there always exists a global *loopfree increment sequence* for any router shutdown operation, we demonstrate that uniformly increasing the distance through a node by λ, where λ represents the smallest distance increment that can be applied on a router, never causes transient loops. Hence, progressively increasing this distance, at worst by repeated increments of λ, can make a router not being used for transit anymore, without incurring transient loops.

Note that, since distance increments correspond to weight increments applied on the outgoing links of the router, we have $\lambda = 1$ in practice.

Theorem 1.1. *In a stable network, incrementing the distance through a router by λ leads to a loop-free convergence process.*

Proof. Let us assume by contradiction that a forwarding loop occurs, for a destination d, between nodes $n_1, n_2, \ldots, n_i, \ldots, n_p = n_1$ during the convergence following an increment of the distance through router r by λ. Consistently with the notations previously introduced, we respectively denotes as $C(n_i)$ and $C'(n_i)$, the shortest path distance from n_i to d before and after the change. We also note $\mathscr{C}(n_i)$ the distance according to which n_i is forwarding packets for destination d at the time a packet is forwarded by n_i along the loop. We will show that this loop cannot possibly occur, for it would require $\mathscr{C}(n_1) > \mathscr{C}(n_1)$.

The cost of the shortest paths to d can either remain the same after the distance through r has been increased, meaning that the paths from n_i to d do not include r anymore, or it can be increased by λ, if the shortest paths still include r. Formally, we have $C'(n_i) = C(n_i)$ or $C'(n_i) = C(n_i) + \lambda$, so that $C'(n_i) \geq \mathscr{C}(n_i)$. Besides, either node n_i is not yet up-to-date at the time it forwards a packet along the loop, and $\mathscr{C}(n_i) = C(n_i)$, or it is and $\mathscr{C}(n_i) = C'(n_i)$. We then have the two following properties:

(P1): If n_i is updated, then it forwards its packets towards neighbors that lie on its post-convergence paths towards d. From properties of shortest paths, this means that the post-convergence distance of these neighbors is strictly lower than the post-convergence distance of n_i. We thus have $C'(n_i) > C'(n_{i+1})$ when n_i is updated, which gives $\mathscr{C}(n_i) > \mathscr{C}(n_{i+1})$.

(P2): If n_i is not updated, n_i is forwarding to neighbors lying along its old shortest paths towards d, which gives $C(n_i) > C(n_{i+1})$, as per properties of shortest paths. Hence $C(n_i) \geq C(n_{i+1}) + \lambda$. By definition of $C(n_{i+1})$, $C(n_{i+1}) + \lambda \geq C'(n_{i+1})$ and, by definition of $\mathscr{C}(n_{i+1})$, we have $C'(n_{i+1}) \geq \mathscr{C}(n_{i+1})$. This gives $\mathscr{C}(n_i) \geq \mathscr{C}(n_{i+1})$, when n_i is not updated.

From $(P1)$ and $(P2)$, we know that $\mathscr{C}(n_i) \geq \mathscr{C}(n_{i+1})$. Note that $\mathscr{C}(n_i) = \mathscr{C}(n_{i+1})$ only when n_i is not updated while n_{i+1} is updated and has increased its distance towards d. If $\mathscr{C}(n_i) = \mathscr{C}(n_{i+1})$ we then have $\mathscr{C}(n_{i+1}) > \mathscr{C}(n_{i+2})$. Besides, if n_{i+1} is updated we have $(P1)$ for $i+1$, thus a strict inequality. Therefore we have $\mathscr{C}(n_i) > \mathscr{C}(n_{i+k})$, $\forall k \geq$

2, $\forall i+k \leq p$ which is in contradiction with the initial loop statement. Also, no transient loop can occur for $p = 2$ and $k = 1$, since a node cannot select itself as its next hop. □

When a router has to be shut down, the length of paths traversing it can thus be progressively increased by λ until it becomes sufficiently large to make it unused. However, such a technique can be inefficient, as a large number of increments would have to be issued when a wide range of metrics is used across the network. For example, the weights assigned to the links of the European Research Network GEANT [gea] are taken from the interval $[1, 20000]$. In theory, the original specification of IS-IS [ISO02] allows for link weights up to $2^6 - 1$ (63) by default, but this limit can be increased to $2^{24} - 1$ $(16, 777, 215)$ on networks supporting *wide metrics* [SL04, Par04]. In OSPF [Moy98], only one type of metric exists, which supports link weights as large as $2^{16} - 1$ $(65, 535)$. In order to overcome this limitation, we propose to perform larger distance increments when they are known to provide loopfree convergence.

Computing short distance increment sequences

Before introducing the methodology, we first discuss some properties among the nodes using each other to reach a destination d, before and after the application of a distance increment. For each path $(n_1, \ldots, n_i, \ldots, n_j, \ldots, n_m)$ in $G_r(d)$, $\forall i < j$, n_i is an *upstream* node of n_j and reciprocally n_j is a *downstream* node of n_i.

These properties are based on a pivot increment, denoted $\Delta^r_d(x)$, that we define as follows:

$$\forall x \in N, \;\; \Delta^r_d(x) = C'(x) - C(x)$$

Note that this notation depends on the destination and the considered distance increment operation, however, for the sake of clarity, we ignore those cumbersome indexes when there is no possible ambiguity. $\Delta^r_d(x)$ is the minimum distance increment to be performed on r, such that there exists a shortest path from x to d that does not include r. A distance increment of $\Delta(x)$ triggers an intermediate change in the forwarding plane of x for d. It forces node x into an *ECMP transient state* where it uses both its initial and final paths towards d. For example, in Fig. 2.1, nodes 2, 3 and b enter an ECMP transient state when the distance through 0 is increased by 8. They use respectively the outgoing edges $(2, 0)$ and $(2, 1)$, $(3, 0)$ and $(3, 1)$, and (b, c) and (b, a). If the distance is increased by any larger value, e.g. $\Delta(x) + \lambda$, node x is in a final state and its shortest paths towards d no longer include r. Note that $\Delta(x) = 0$ for all nodes not in $G(d)$, but a node $x \in G(d)$ may also verify $\Delta(x) = 0$. A node x verifies $\Delta(x) = 0$ if at least one of its initial shortest paths to d does not include r. Either x does not use r at all to reach d, or some of its equal-cost shortest paths do not contain r. This is the case for node 4 in Fig. 2.1.

Let us now introduce three fundamental properties related to delta differences. We call an *PRE* (respectively *POST*) edge an edge (x, y) that starts a subset of paths in $P(x, d)$ but not in $P'(x, d)$ (respectively starts $P'(x, d)$ but not $P(x, d)$). When an edge (x, y) starts a subset of paths $\mathscr{P} \in P(x, d)$, we have $\mathscr{P} = (x, y) \circ P(y, d)$ (it is the same for P'). A *COMMON* edge starts both a path in $P(x, d)$ and one in $P'(x, d)$ (the two paths may differ further). Let (x, y) be an edge in G, we have the three following properties:

Property 1.1. *If* (x, y) *is an* ***PRE****-edge then* $\Delta(x) < \Delta(y)$*.*

Proof. On the one hand, if node x is not updated and continues to use its PRE neighbor y, we have by definition $C(x) = C(y) + w(x, y)$. On the other hand, if node x does not use y anymore once updated we have $C'(x) < C'(y) + w(x, y)$.
Thus, $\Delta(x) = C'(x) - C(x) = C'(x) - (C(y) + w(x, y)) < C'(y) + w(x, y) - C(y) - w(x, y) = \Delta(y)$. □

Property 1.2. *If* (x, y) *is a* ***POST****-edge then* $\Delta(x) > \Delta(y)$*.*

Proof. When node x updates its path towards d and decides to go through a POST neighbor y, we have by definition $C'(x) = w(x, y) + C'(y)$. We also have $C(x) < C(y) + w(x, y)$: x was not using y as a next hop towards d before the update.
Thus, $\Delta(x) = C'(x) - C(x) = w(x, y) + C'(y) - C(x) > C'(y) - (C(y) + w(x, y)) + w(x, y) = \Delta(y)$. □

Property 1.3. *If* (x, y) *is a* ***COMMON****-edge then* $\Delta(x) = \Delta(y)$*.*

Proof. When node x updates its path towards d and still uses its PRE neighbor y towards d, we have by definition $C'(x) = w(x, y) + C'(y)$. We also have $C(x) = C(y) + w(x, y)$: x was using y as a next hop towards d before the update.
Thus, $\Delta(x) = C'(x) - C(x) = w(x, y) + C'(y) - C(x) = C'(y) - (C(y) + w(x, y)) + w(x, y) = \Delta(y)$. □

Fig. 2.1 illustrates such properties. Along the PRE-path $a \to b \to c$, we have $\Delta(a) = 6 < \Delta(b) = 8 < \Delta(c) = 10$. Reciprocally, we notice $\Delta(c) < \Delta(b) < \Delta(a)$ along the POST forwarding path towards d. We also have $\Delta(3) > \Delta(4) = 0$ on the POST-edge $(3, 4)$, since node 3 is not in $G_r(d)$. Finally, on a common edge such as (d, c), we have $\Delta(d) = \Delta(c)$.

A first interesting consequence of such properties is that delta values are increasing along paths in $G_r(d)$ (PRE and COMMON edges). Hence, sorting all delta values for nodes in $G_r(d)$ yields a list of strictly increasing weight increments for each pair (r, d).

Formally, we denote as $\Delta_i + \lambda$ sequence, the sorted union of $\Delta(x) + \lambda$ for each node x in $G_r(d)$. From an edge perspective, such a sequence ensures that each PRE-edge (x, y) connecting an upstream node x to a downstream one y stops being used before a POST, or transient (containing at least one POST-edge), sub-path re-links y to x, thus creating a cycle. This sequence makes it possible to update upstream nodes strictly before downstream ones, or in the same convergence period if (i) they share a common delta value, or (ii) if their delta values only differ by λ ($\Delta(y) = \Delta(x) + \lambda$).

Δ_i and Δ_{i+1} being two consecutive delta values in the sequence, a routing change occurs only once between $\Delta_i + \lambda$ and Δ_{i+1}. Precisely, POST-edges start being used as the distance increment reaches Δ_{i+1}, but PRE ones used for $\Delta_i + \lambda$ and Δ_{i+1} are the same. Thus, if a loop occurs during the transition from $\Delta_i + \lambda$ to $\Delta_{i+1} + \lambda$, it also occurs during the transition from Δ_{i+1} to $\Delta_{i+1} + \lambda$. From Theorem 1.1, we know that no loop can occur while incrementing the cost of a path by λ, so that the transition from Δ_{i+1} to $\Delta_{i+1} + \lambda$ is loop free, as well as from $\Delta_i + \lambda$ to $\Delta_{i+1} + \lambda$. Therefore, using the $\Delta_i + \lambda$ sequence for a pair (l, d) provides a loop free convergence towards d, provided that the network completely converges between two consecutive increments. Such properties on delta values are also the basis of the algorithms we propose in the next sections. Since removing one edge from an elementary cycle is a sufficient condition to avoid it, we can deduce a necessary condition to avoid a transient loop. At least one node in the cycle has to be fully updated, no longer using its PRE edges, before the distance through r is increased by a value greater than or equal to the maximum delta value of all nodes in the cycle.

We showed in this section that a loop-free sequence of uniform distance increments exists for any router shutdown operation. However, the $\Delta_i + \lambda$ sequence may contain elements that are not necessary to avoid transient loops. Such unnecessary elements may result from the union of $\Delta_i + \lambda$ value of distinct sub-shortest paths upstream of the modified router. Intermediate increments for different destinations could also be merged within the limits of distance increment intervals. Finally, performing non-uniform increments could open more combinations and allow for even shorter sequences.

1.2 Towards non-uniform multi-link increments

Applying a sequence of uniform distance increments practically means that the IGP weights configured on the outgoing links of the router have to be increased by the same value at each step. This constraint prevent from computing minimal sequences, and also makes it more difficult to finely tune the final link weights.

This is not a problem if the router, or a subset of its links, has to be shut down. As

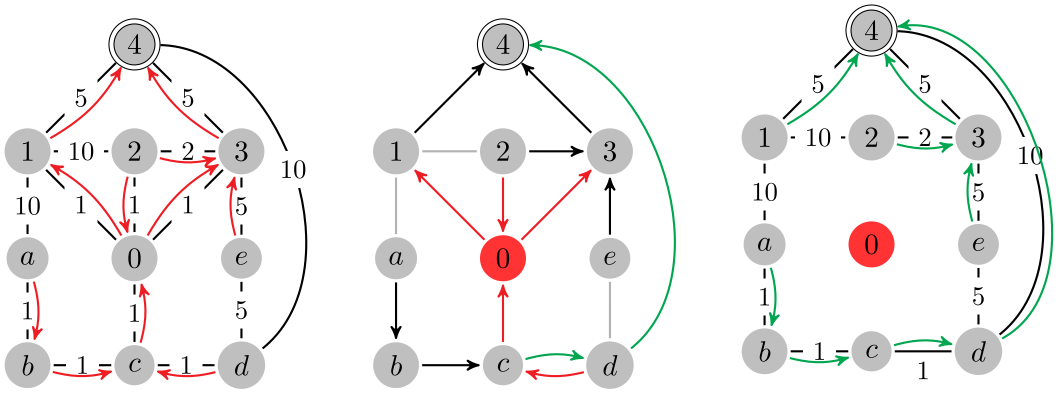


FIGURE 2.4: paths towards destination 4

long as the modified components are not used for transit anymore, their actual IGP weights do not really matter. However, if the weights have to be set to specific values, it could be necessary to compute and successively apply multiple distance increment sequences. Each sequence would increase the weights configured on all but the links that are already set at their final value, by the smallest required increment minus the sum of those previously performed. For example, consider a node having three outgoing links with the same initial weight of 1. If the weights on these links have to be set to 10, 16 and 25, respectively, three successive increment sequences would have to be applied. This first one would increase the weights on all links by 9, setting them all to 10. The second sequence would increase the weights on the second and third links by 6, so that the weight on the tree links would respectively be equal to 10, 16 and 16. Finally, the last sequence would increase the weight on the third link to 25.
Most of all, uniformity represents a non necessary constraint added to the problem that prevents from achieving minimality in terms of sequence length, even without a specific final case.

In order to compute minimal weight increment sequences we thus have to consider simultaneous and non-uniform multi-link increments. Computing such minimal sequences is challenging for two main reasons. First, all the destinations in the network must be taken into account, as our goal is to minimize the number of steps across them. This problem also exists when aiming for minimality considering distance increments. However the solution space is scalar in such case, while it is k-dimensional when considering non-uniform multi-link increments on a node of degree k. Second, applying several weight increments in a single LSA may lead to the use of next-hops that do not correspond to either initial nor final ones. It is then not sufficient to only rely on those two *end point states* to capture every intermediate next-hop change that may occur in the network while the sequence is applied. Considering the example on Fig. 2.4, the updated node 0 may transiently use node 2 as one of its next hops towards 4 during the convergence if the weights on links $(0, 1)$, $(0, 2)$ and $(0, 3)$ are respectively set to 4, 2 and 4, which are the minimal values to avoid the loop between c and d. Those *intermediate next-hops* possibly lead to additional transient loops. In the initial state given in Fig. 2.4, the shortest paths from 2 to destination 4 include 0 as an ECMP next hop. Hence, when applying the increment suggested above, a transient loop can occur between 0 and 2, which depends on the values in the LSA sent to avoid the initial loop. We call such a loop an *intermediate forwarding loop*, for it is triggered by intermediate forwarding changes. Note that loops are not a necessary consequence of intermediate next-hops. They only occur in specific circumstances that we describe in section 3.

In the next section, we put this last problem aside and present an algorithm to compute minimal weight increment sequences, preventing non-intermediate transient loops for any node removal operation. We address the problem of intermediate next-hops and loops in section 3, proposing several solutions to prevent either all intermediate next-hops or only the loops they may induce. We also show that, in practice, intermediate forwarding loops can be prevented even without modifying the weight increment sequences, by using additionnal local mechanisms.

2 Computing minimal weight increment sequences

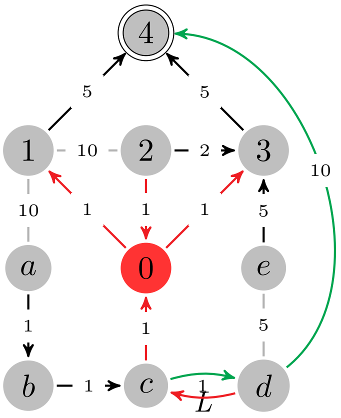
This section aims at providing a theoretical framework for non-uniform weight increment sequences. We model a weight increment as a vector v, having $k = |v|$ components. For any weight increment v, a given component $v[i]$ corresponds to the weight increment applied to the i-th outgoing link of the modified router. For simplicity, we often refer to the component of a vector using the identifier of the neighbor at the far end of the link. We also define a partial order relationship between vectors the same size. Thus, we say that two vectors v_1 and v_2 of size $k > 0$ are equal, i.e., $v_1 = v_2$, if $\forall\, i \in [\![1, k]\!],\ v_1[i] = v_2[i]$. Similarly, $>$, $\geq$, $<$, $\leq$ relationships, hold on vectors if they hold on all the corresponding components. In addition, given two vectors v_1 and v_2 (such that $|v_1| = |v_2| = k$), we say that v_1 is *positively greater* than v_2, and note $v_1 >^+ v_2$, if

$$\forall\, i \in [\![1, k]\!], \qquad \begin{cases} v_1[i] > v_2[i] & (\text{if } \ v_2[i] \in \mathbb{N}) \\ v_1[i] \geq 0 & (\text{if } v_2[i] \in \mathbb{Z}_{<0}) \end{cases}$$

This new relationship is later used to define the positive intermediate vectors that constitute a weight increment sequence. It reflects our assumption that the weight of any link outgoing from r is always greater than or equal to its initial weight. That is, since we aim at offloading traffic from the router to be removed, we do not consider sequences including weight decrements with respect to the initial state, as this would make r more attractive. Nevertheless, we admit negative components in weight increments, e.g., if following positive increments. Note that the conditional statements in parenthesis are implicit in the inequalities, and merely provided for information.

2.1 Defining necessary constraints for loop avoidance

We now define the concept of loop-constraint to formalize the property a weight increment sequence must satisfy to provably avoid transient loops. More precisely, we define a *loop-constraint*, or simply *constraint*, as the weight increment interval associated to a single loop. For any given transient loop L, a loop-constraint l is a pair of vectors $l := (\underline{l}, \overline{l})$. Vectors $\underline{l}$ and $\overline{l}$ have one component per outgoing link of the modified router r (i.e., $|\underline{l}| = |\overline{l}| = k$, where k is the degree of router r), and respectively represent the lower and upper bounds of the constraint associated with L. To compute the actual bounds of loop-constraints, we rely on a vectorial variant of the *delta* values previously introduced, which we refer to as *delta vectors*. Given a router $x \neq r$ and a destination d, we denote as $\Delta_d(x)$ the vector of weight increments such that the shortest paths from x to d include both the initial and final paths (as computed in G and G', resp.). Let $C'(x, d)$ be the cost of the shortest paths from x to d in G', l_i be the i-th link outgoing



$$\Delta_4(c) = \begin{pmatrix} 11-(1+6) \\ 11-(1+8) \\ 11-(1+6) \\ 11-(1+12) \end{pmatrix} = \begin{pmatrix} 4 \\ 2 \\ 4 \\ -2 \end{pmatrix}$$

$$\Delta_4(d) = \begin{pmatrix} 10-(2+6) \\ 10-(2+8) \\ 10-(2+6) \\ 10-(2+12) \end{pmatrix} = \begin{pmatrix} 2 \\ 0 \\ 2 \\ -4 \end{pmatrix}$$

$$l = (\ \underline{l} = \Delta_4(d),\ \bar{l} = \Delta_4(c)\ )$$

Figure 2.5: Delta and constraint vectors calculation

from 0, and $C(x, l_i, d)$ be the cost of the shortest path from x to r plus the cost of the shortest *simple* path from r to d via l_i in G. *Delta vectors* are formally defined as

$$\Delta_d^r(x)[i] = C'(x, d) - C(x, l_i, d)$$

Then, the loop-constraint l associated to a loop L to a destination d is defined as

$$l := (\underline{l} := \min_{\forall x \in L}(\Delta_d^r(x)), \bar{l} := \max_{\forall x \in L}(\Delta_d^r(x)))$$

Note that, for a given destination d, the set of vectors $\Delta_d(x) \forall x \in N$ is totally ordered and can be assimilated to scalars. Indeed, for any router x, we have $C(x, l_i, d) = C(0, l_i, d) - C(0, d) + C(x, d)$. In other words, for a given destination the offsets between the *delta vector* components are the same for every router.

By definition of delta vectors, the vector v_x verifying $\forall\ i \in [\![1, k]\!],\ v_x[i] = \max(\Delta_d(x)[i] + 1, 0)$ is the smallest set of increments to be configured on the outgoing interfaces of router r, such that router x switches to its final state and no longer uses r to reach d. Hence, in order to satisfy a loop-constraint l such that $\underline{l} = \Delta_d(z)$ and $\bar{l} = \Delta_d(y)$, an intermediate vector v must be *positively greater* than $\Delta_d(z)$, but not greater than or equal to $\Delta_d(y)$.

An illustration of delta and constraint vectors calculation is provided on Fig. 2.5. In this example, $\Delta_4(c) = (4\ 2\ 4\ -2)$ and $\Delta_4(d) = (2\ 0\ 2\ -4)$, where components respectively map to links $(0, 1)$, $(0, 2)$, $(0, 3)$, and $(0, c)$. Since $C'(c, 4) = 11$ and $C(c, (0, 1), 4) = 7$, we have $\Delta_4(c)[1] = 4$. This value indicates that adding 4 to the weight configured on link $(0, 1)$ makes the path from c to 4 through $(0, 1)$ as long as the final ones. Similar calculations are performed for every other component of $\Delta_4(c)$ and $\Delta_4(d)$. According to those calculations, forwarding paths from c (resp., d) are ensured not to include 0 if weight increments greater than $\Delta_4(c)$ (resp., $\Delta_4(d)$) are applied to the outgoing links of

0. Besides, the constraint l associated to the loop L between c and d is formalized as $l = (\Delta_4(d), \Delta_4(c))$.

By definition of l, applying weight increments positively greater than $\underline{l}$ (resp. $\bar{l}$) will cause the shortest paths from at least one router (resp. all the routers) in L not to traverse r anymore. In the previous example, applying a weight increment positively greater than $\underline{l} = \Delta_4(d)$ will cause d, but not necessarily c, to switch to its final shortest paths. Both c and d are guaranteed to switch to their respective final paths when the weight increments is positively greater than $\bar{l} = \Delta_4(c)$. To provably avoid a transient loop, we must then force weight increments changing only to forwarding paths of d, e.g. a relative increase of (3 1 3 0), before applying the final weights.

To formally state the problem of finding such intermediate weight increments, we introduce the following terminology. We say that a weight increment v ***meets*** a constraint $(\underline{l}, \bar{l})$ if $v >^+ \underline{l}$ and $\exists\, i \in [\![1, k]\!] \mid v[i] < \bar{l}[i]$. We also say that a weight increment v ***precedes*** a constraint l if $\exists\, i \in [\![1, k]\!] \mid v[i] \leq \underline{l}[i]$, and that v ***follows*** l if $\forall\, i \in [\![1, k]\!] \mid v[i] \geq \bar{l}[i]$. Given a constraint l and a sequence of weight increments $\{v_0, \dots, v_n\}$, where $v_0 = \vec{0}$ is the initial state of node r and $v_n = \vec{\infty}$ represents the final routing state of r (after *MAX_METRIC* or an *overload bit* has been configured), a pair of consecutive vectors v_i and v_{i+1} constitutes an *unsafe transition* if either i) v_i precedes $\underline{l}$ and v_{i+1} follows $\bar{l}$; or ii) v_i follows $\bar{l}$ and v_{i+1} precedes $\underline{l}$. Trivially, a pair of consecutive vectors is said to form a ***safe transition*** with respect to a given constraint if it is not unsafe.

In the previous example, setting router 0 directly in its final state, which we represent as the sequence $\{\vec{0}, \vec{\infty}\}$, is an *unsafe transition* with respect to constraint $l = \{(2\ 0\ 2\ 0), (4\ 2\ 4\ 0)\}$. On the contrary, both transitions in sequence $\{\vec{0}, (3\ 1\ 3\ 0), \vec{\infty}\}$ are safe with respect to l, since the intermediate vector (3 1 3 0) *meets* l. Loop-constraints and intermediate vectors can be graphically represented as in Fig. 2.6. For the sake of clarity, we show only two dimensions that correspond to links $(0, 1)$ and $(0, 2)$ (the first two indices in the vectors). A constraint is represented as a colored L-shape, whose arms are the intervals of possible vector values. Any vector whose representation is within the L-shape *meets* the constraint, while vectors outside the L-shape either *precede* or *follow* the constraint. In this example, v_1 and v_2 precede l for their representations are respectively on the left and below the L-shape corresponding to constraint l. On the opposite, v_6 follows l because it is represented on the right above the constraint. Only v_3, v_4 and v_5 are within the L-shape and actually meet l.

From the definition of delta vectors and loop constraints, we can deduce the following properties. We refer to an arbitrary loop L considering a given destination d, its corresponding constraint l, and a given weight increment v applied to links outgoing from 0.

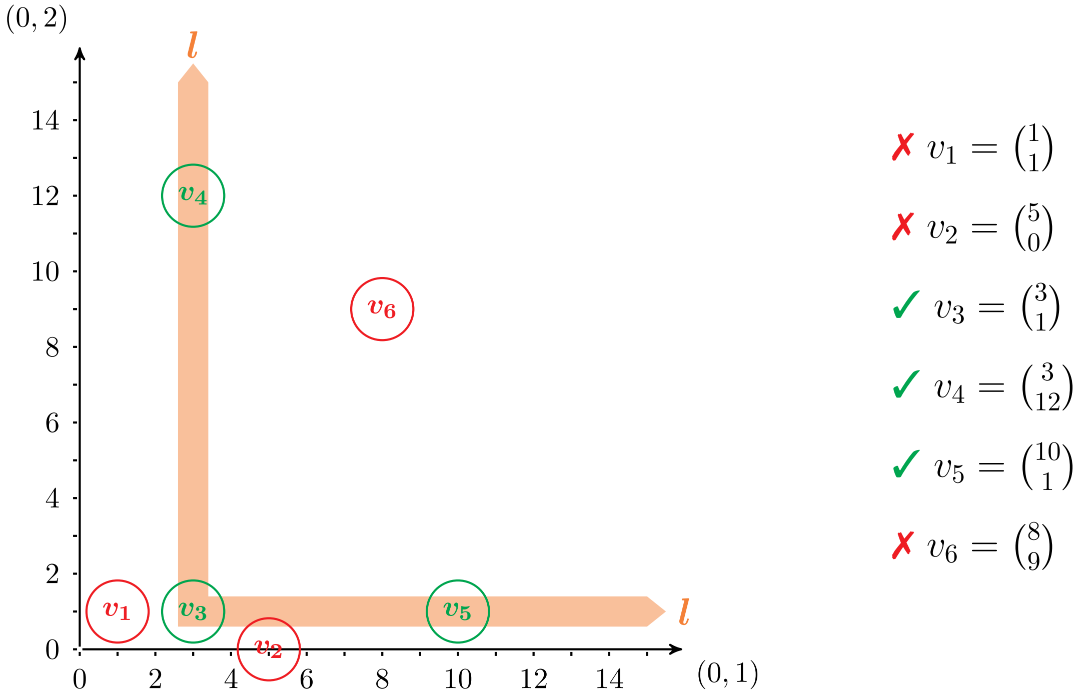


FIGURE 2.6: Graphical representation of constraints and vectors

For any router x, we denote its successors in $RSPDAG(d, G)$ and $RSPDAG(d, G')$ as its PRE and POST next-hops to d, respectively.

Property 2.1. *Given a constraint l, $\forall i \in [\![1, |l|]\!]$, $\bar{l}[i] \geq \underline{l}[i] + 2$.*

Property 2.2. *If v meets l, there exist at least two routers $x, z \in L$ such that (i) x uses its POST next-hops $y_1, \ldots, y_n$ to d, with $y_1, \ldots, y_n \notin L$ because $\Delta_d(x) = \underline{l}$; and (ii) z uses its PRE next-hops $w_1, \ldots, w_n$ to d, with $w_1, \ldots, w_n \notin L$ because $\Delta_d(z) = \bar{l}$.*

Property 2.3. *All routers $x \in L$ use their respective (i) PRE next-hops to d if v precedes l, and (ii) their POST next-hops to d if v follows l.*

We leverage these properties to prove that loop-constraints are necessary and sufficient conditions to prevent transient loops.

Theorem 2.1. *A weight sequence s avoids a loop L if and only if s contains only safe transitions with respect to the constraint corresponding to L.*

Proof. Let $l = (\underline{l}, \bar{l})$ and d respectively be the loop constraint and the destination associated to loop L. We prove the statement in two steps.

- if s includes an unsafe transition $(v_i\ v_{i+1})$ for l, then s does not prevent L. Indeed, by definition of unsafe transition, we have two cases: (i) v_i precedes $\underline{l}$ and v_{i+1} follows $\bar{l}$, and (ii) v_i follows $\underline{l}$ and v_{i+1} precedes $\bar{l}$. All the routers in L will switch from their PRE to their POST next-hops to d in the first case, and from their POST to their PRE next-hops in the second case. In both cases, the transition from v_i to v_{i+1} can cause L to occur by definition of transient loop.

- if s only includes safe transitions for l, then s prevents L. Indeed, by definition of safe transition, for each pair of weight increments v_i and v_j, where v_i precedes l, v_j follows l and $j > i$, there must exist a vector v_k such that $i < k < j$ and v_k meets l. By Property 2.2, this means that each time routers in L switch from their PRE to their POST next-hops, there is an intermediate step (corresponding to v_k) in which some routers switch before others in such a way that the possible loop is prevented. A symmetric argument can be applied to the case in which v_i follows l and v_j precedes l.

The two cases prove the statement. □

Theorem 2.1 implies that, for each constraint $(\underline{l}, \overline{l})$, at least one vector must meet the constraint for each transition from weight increments smaller than $\underline{l}$ to those greater than $\overline{l}$, and vice versa. *Always increasing* sequences thus seem a natural candidate for targeting minimality, as each constraint would have to be met only once. Note that we define as *always increasing* any sequence $s = \{v_0, \ldots, v_m\}$ verifying $\forall\, i \in [\![1, m]\!]$, $v_{i-1} \leq v_i$. A simplified version of Theorem 2.1 holds for always increasing sequences.

Theorem 2.2. *An always increasing weight sequence s avoids a loop L if and only if s contains at least one vector meeting the constraint corresponding to L.*

Proof. Let $l = (\underline{l}, \overline{l})$ be the constraint corresponding to any loop L. By definition of always increasing sequence, s is a concatenation of three subsequences, $s = l\ m\ h$, where l is composed by vectors preceding $\underline{l}$, m contains vectors meeting l, and h includes vectors following $\overline{l}$. By hypothesis, m cannot be empty. Thus, s does not contain unsafe transitions for l. The statement then follows by Theorem 2.1. □

Constraint extraction

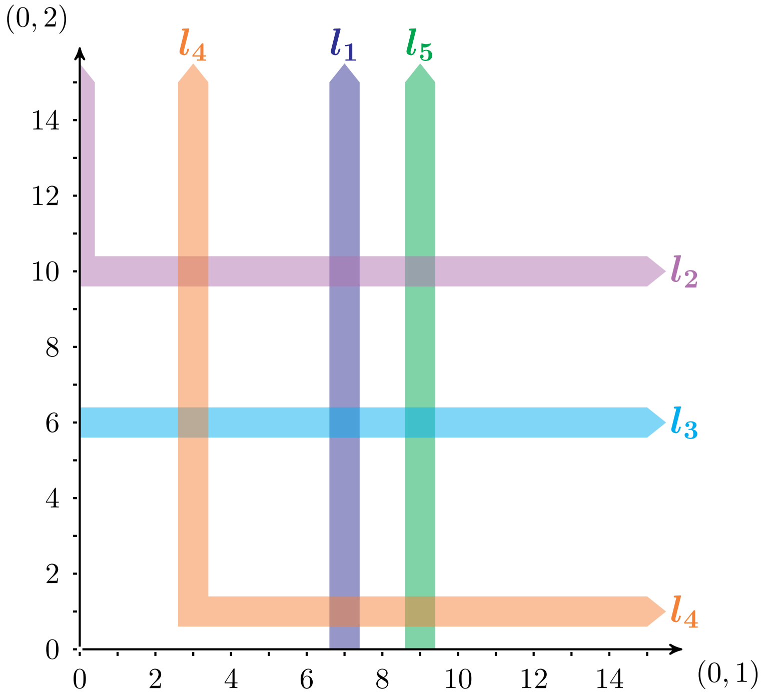
In practice, there exists several methods to retrieve the list of all loop-constraints associated to a given operation. The most intuitive one is to enumerate, for every destination, the elementary circuits in the merged RSPDAG, and extract the minimum and maximum delta vectors among the nodes in each circuit. However, the worst-case time complexity of the best enumeration algorithm to our knowledge [Joh75] is in $O((|N| + |E|)(c + 1))$, where c is the number of elementary circuits. Since there can be up to $\sum_{i=1}^{|N|-1} \binom{|N|}{|N|-i+1}(|N|-i)!$ circuits in a complete directed graph, using this algorithm might lead to very long computing times on large networks. Fortunately, it is not necessary in our situation to accurately enumerate each cycle in order to retrieve the associated constraint. Another solution consists in computing a transitive closure [Dij59, Flo62] of the merged RSPDAG for each destination, and detecting cycles as a path existing between two nodes in both directions. This method needs to be slightly tuned in practice, in order to extract the minimum and maximum delta values, yet it yields the best results in terms of computing time to our knowledge. For a given destination, the actual number of constraints is indeed limited by the combinations of two different delta values, which are at most equal to $|N| \times (|N|-1)/2$. Besides, in the case of overlapping constraints, only the most restrictive one has to be considered. This further reduces the maximum number of constraints to be considered for one destination to $|N|$.

On Fig. 2.7, we show the constraints associated to each transient loop that could possibly occur when router 0 is shut down. Constraints l_1 and l_5 correspond to two potential transient loops for destination 1, while l_2, l_3 and l_4 respectively map loops for destinations 2, 3 and 4. Since we do not allow for negative increments with respect to the initial weights, the graphical representation appears truncated for constraints whose first two components include negative values. Constraints l_1 and l_5, which have positive values only on their first components, are thus represented as vertical strips on the figure. This means that they can only be met by a vector whose first value is between the bounds of the constraint, regardless of the other components. One the contrary, l_3 is depicted by a horizontal strip for it may not be satisfied on the first index.

In this context, we study the problem of finding minimal safe sequences with respect to all constraints. In particular, we present algorithms to compute always increasing sequences, that are provably minimal and safe. This also implies that restricting to always increasing sequences does not limit our ability to optimally solve the safe router update problem. That is, for every router shutdown operation, there exists at least one minimal safe sequence which is always increasing.

2.2 A greedy backward algorithm for computing minimal sequences

In order to minimize the operational impact of our approach, it is necessary to minimize the number of intermediate weight increments to be performed. We designed an algorithm that aims at computing a sequence of vectors satisfying both the safety and the minimality properties. That is, (i) no transient loop could appear in the network between successive updates or at either end of the sequence, and (ii) there exists no



$$l_1 = \left(\begin{pmatrix}\mathbf{6}\\\mathbf{-4}\\-4\\-6\end{pmatrix}, \begin{pmatrix}\mathbf{8}\\\mathbf{-2}\\-2\\-4\end{pmatrix}\right)$$

$$l_2 = \left(\begin{pmatrix}\mathbf{-1}\\\mathbf{9}\\7\\-4\end{pmatrix}, \begin{pmatrix}\mathbf{1}\\\mathbf{11}\\9\\-2\end{pmatrix}\right)$$

$$l_3 = \left(\begin{pmatrix}\mathbf{-3}\\\mathbf{5}\\7\\-4\end{pmatrix}, \begin{pmatrix}\mathbf{-1}\\\mathbf{7}\\9\\-2\end{pmatrix}\right)$$

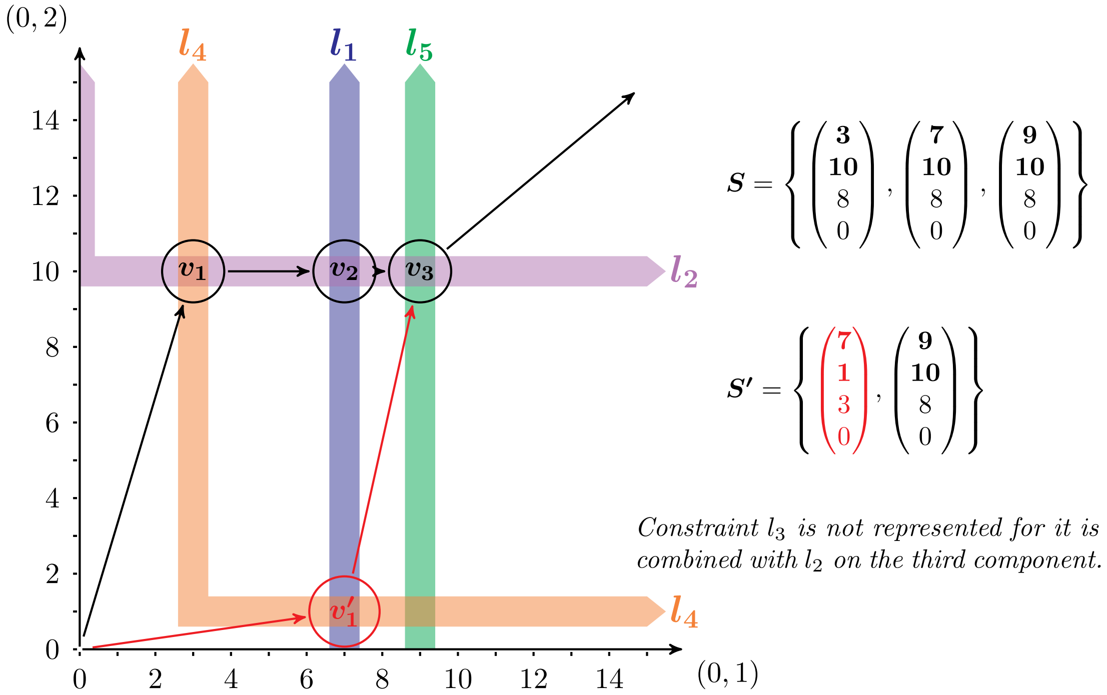
$$l_4 = \left(\begin{pmatrix}\mathbf{2}\\\mathbf{0}\\2\\-4\end{pmatrix}, \begin{pmatrix}\mathbf{4}\\\mathbf{2}\\4\\-2\end{pmatrix}\right)$$

$$l_5 = \left(\begin{pmatrix}\mathbf{8}\\\mathbf{-2}\\-2\\-4\end{pmatrix}, \begin{pmatrix}\mathbf{10}\\\mathbf{0}\\0\\-2\end{pmatrix}\right)$$

FIGURE 2.7: Loop-constraints for all destinations affected by the removal of router 0

shorter sequence satisfying this property. However, while the first property is easily met once all the constraints have been extracted, the second one is more challenging due to the partial order relationship among vectors. Since a vector only needs to be lower than the upper bound of a constraint on one component, a forward based greedy algorithm (similar to the one presented in [FSB07] and [CMP+14]) would have to decide, at each step, which components are to remain below the upper bound of the lowest constraint and which are to be increased, in order to meet multiple constraints. Thus, several combinations of constraints could be possible at each step, which does not necessarily lead to sequences of the same length.

Let us consider the constraints represented on Fig 2.7. Based on the graphical representation and the vector values, we notice that l_4 is to be met first. This constraint can be combined together with either l_2 and l_3 (that is combined with l_2 on the third component), or l_1, but not the three of them. A greedy option would thus be to meet three constraints, l_2, l_3 and l_4 with a single intermediate vector, denoted v_1 on the figure. This leaves two constraints, l_1 and l_5, unsatisfied. Since they cannot combine, two additional vectors, v_2 and v_3, are required to complete the sequence. Another option would consist in combining only two constraints, l_1 and l_4, at the first step. While this may appear less interesting, it would enable the three remaining ones to be met together at a second step, making the final sequence $\{v_1', v_3\}$ one vector shorter than the greedy option. We illustrate the sequence calculation process on Fig. 2.8, graphically showing how intermediate vectors meet each loop-constraint and form safe sequences.



$$\boldsymbol{S} = \left\{ \begin{pmatrix} \mathbf{3} \\ \mathbf{10} \\ 8 \\ 0 \end{pmatrix}, \begin{pmatrix} \mathbf{7} \\ \mathbf{10} \\ 8 \\ 0 \end{pmatrix}, \begin{pmatrix} \mathbf{9} \\ \mathbf{10} \\ 8 \\ 0 \end{pmatrix} \right\}$$

$$\boldsymbol{S'} = \left\{ \begin{pmatrix} \mathbf{7} \\ \mathbf{1} \\ 3 \\ 0 \end{pmatrix}, \begin{pmatrix} \mathbf{9} \\ \mathbf{10} \\ 8 \\ 0 \end{pmatrix} \right\}$$

Constraint l_3 is not represented for it is combined with l_2 on the third component.

Figure 2.8: Sequence calculation on a forward mode

In this example, there is only one choice to be made, and one can easily compute and compare both possibilities. However, on larger graphs such case could appear at several steps, with possibly more than two options each time. A brute-force algorithm could definitely find the minimal solution by exploring each constraint combination and returning one of the shortest resulting sequences. Yet the solution would come at a significant cost, as the number of possibilities can be combinatorial. This minimization problem comes from the flexibility of the upper bound of the constraints. A typical forward mode algorithm would in fact increase the vector values as much as possible, considering predefined constraints. Contrary to scalar increments, though, it is unclear which values are to be increased, and by how much, and which are not.

On the other hand, the lower bounds of loop-constraints are strict. That is, an intermediate vector can meet a constraint only if all of its components are larger than the lower bound. It is sufficient that one component of the vector be below the lower bound for the vector to precede the constraint. Hence, vector values can be greedily pushed towards this bound without implying any decision process. In other words, when looking at the constraints on a backward mode, there is only one possible lowest vector that does not precede any of them. Each component of this vector is equal to the lowest possible value meeting the constraint with the largest lower bound. The vector is computed such that decrementing any component necessarily leads to at least one constraint being preceded. Besides, we can show that this minimal vector meets at least one constraint. By Property 2.1, this vector is indeed lower on at least one component than the upper bound

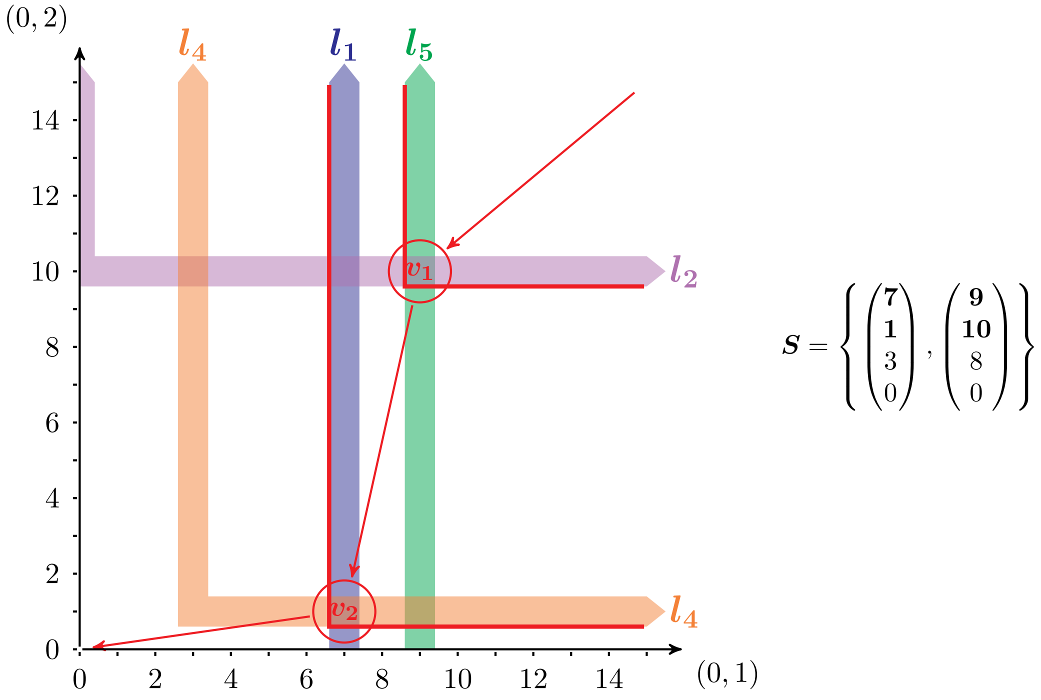


FIGURE 2.9: Sequence calculation on a backward mode

of each constraint whose lower bound was used to compute the vector. Thus, since the vector is positively greater than the lower bounds of these constraints and lower than their upper bounds on at least one component, it necessarily satisfies these constraints. The same process can then be repeated considering only the constraints that the first vector left unsatisfied.

As an example, on Fig. 2.7, the constraints having the largest upper bounds are c_5 on the first component, c_2 on the second and third, and c_3 on the third. The lowest possible vector is thus equal to (9 10 8 0), which satisfies all three constraints. Then, by repeating the same process with unsatisfied constraints, we obtain a second vector equal to (7 1 3 0), which satisfies all the remaining constraints. Considering that constraints c_1 and c_5 cannot possibly be satisfied by the same intermediate vector, the resulting sequence $\{(7\ 1\ 3\ 0), (9\ 10\ 8\ 0)\}$ is of minimal length. On Fig. 2.9, we illustrate this backward calculation process, representing for each step the *constraint precedence limit* as a red line.

The algorithm derived from this principle is called Greedy Backward Algorithm (GBA) and presented in Alg. 1. From a set of loop-constraints $\mathcal{L}$, which we assume have been extracted beforehand, it consists in finding, at each iteration, the lowest possible vector that is positively greater than the lower bound of every constraint $l \in \mathcal{L}$. Formally, each of its components is calculated as the maximum value among the lower bounds of unsatisfied constraints plus one (ll. 4-9). Each constraint this vector satisfies is then removed from the set (ll. 10-16), and the process is repeated until there is no more unsatisfied constraints (l. 3).

Algorithm 1 GBA Core

```
1:  function GreedyBackwardAlgorithm($\mathcal{L}$)
2:      $sequence \leftarrow \varnothing$
3:      while $\mathcal{L}$ not $\varnothing$ do
4:          $vector \leftarrow \vec{0}$                                  } Compute the current vector
5:          for $l \in \mathcal{L}$ do                                   }
6:              for $i \in 1 \dots k$ do                                 }
7:                  $vector[i] \leftarrow \mathbf{max}(vector[i], \underline{l}[i] + 1)$  }
8:              end for                                                  }
9:          end for                                                      }
10:         for $l \in \mathcal{L}$ do                                   } Remove satisfied constraints
11:             for $i \in 1 \dots k$ do                                 }
12:                 if $vector[i] \leq \bar{l}[i]$ then                  }
13:                     $\mathcal{L}$.remove ($l$)                       }
14:                 end if                                               }
15:             end for
16:         end for
17:         $sequence$.append ($vector$)          ▷ Add the new vector to the sequence
18:     end while
19:     return $sequence$                         ▷ Return a minimal loop-free sequence
20: end function
```

Safety and minimality

We now prove that GBA computes weight sequences that prevent convergence loops. And, most of all, that these weight sequences are of minimal length. Formally, we show that GBA optimally solves the following problem.

Problem 2.1. *Minimal Loop-free Problem (MLP): Given a set $\mathcal{L}$ of loop-constraints, compute a minimal weight increment sequence that contains no unsafe transition for any constraint in $\mathcal{L}$.*

In our proofs, we use the term *before* iteration j to denote all previous iterations considering a backward sequence building. We also say that a constraint is *unsatisfied* (resp. *satisfied*) at an iteration j if it is not met (resp. it is met) by any vector computed by GBA before j. Our proofs leverage the following properties of GBA that hold by definition of the algorithm.

Property 2.4. *At each iteration j, GBA computes a vector v such that $v >^+ \underline{l}$ for all the constraints $l = (\underline{l}, \bar{l})$ still unsatisfied before j.*

Property 2.5. *At each iteration j, GBA computes a vector v such that, for each component i, a constraint $l = (\underline{l}, \overline{l})$ still unsatisfied before j exists that meets $v[i] = \max(\underline{l}[i] + 1,\ 0)$.*

Property 2.6. *GBA computes always increasing sequences.*

Property 2.7. *GBA stops as soon as all the constraints are met.*

Properties 2.4 and 2.5 are ensured by the greedy vector computation. Property 2.6 is the result of both vector computation and constraint removal. Property 2.7 derives from the constraint removal mechanism.

These properties are used to show that, given any initial set of constraints, GBA always terminates and produces a sequence of minimal length.

Lemma 2.1. *At each iteration, GBA computes a vector v that meets at least one constraint not met before.*

Proof. Consider any iteration j of GBA. Let $\mathcal{L}$ be the set of unsatisfied constraints at the beginning of iteration j, and let v the vector computed by the algorithm during the iteration j. By Properties 2.4 and 2.5, there must exist at least one constraint $l \in \mathcal{L}$ such that $l = (\underline{l}, \overline{l})$, $v >^+ \underline{l}$ and $\exists i | v[i] = \underline{l}[i] + 1$. By Property 2.1, then $\exists i | v[i] < \overline{l}[i]$. This means that l is met by v, hence the statement. □

We now leverage Lemma 2.1 to prove that GBA always terminates in a finite number of iterations, bounded by $|\mathcal{L}| < |N|^2$.

Theorem 2.3. *For any MLP instance $I =< \mathcal{L} >$, GBA always terminates in $O(|\mathcal{L}|)$ iterations.*

Proof. By Property 2.7, GBA stops when all the initial constraints $\mathcal{L}$ are met. Hence, the statement directly follows by Lemma 2.1. □

We also show that the sequences computed by GBA are guaranteed to be safe.

Theorem 2.4. *The weight sequences computed by GBA prevent all transient loops.*

Proof. Let I be any MLP instance, and let s be the sequence computed by GBA on I. By Property 2.6, s is an always increasing sequence. By Lemma 2.1, each constraint is met by at least one vector in s. Thus, the statement follows by Theorem 2.2. □

Eventually, we prove the minimality of the sequences computed by GBA.

Lemma 2.2. *Let $I =< \mathcal{L} >$ be any MLP instance, $s = (s_1 \dots s_n)$ be any sequence solving I, and $g = (g_1 \dots g_m)$ be the sequence computed by GBA on I, with possibly $n \neq m$. The last vectors of the sequences verify $s_n \geq g_m$.*

Proof. Assume by contradiction that $s_n[i] < g_m[i]$ for a given component i. By Property 2.5, there must exist at least one constraint $l = (\underline{l}, \bar{l}) \in \mathcal{L}$, such that $g_m[i] = \underline{l}[i] + 1$. This implies that $s_n[i] \leq \underline{l}[i]$, hence l is not met by s_n. This means that s contains an unsafe transition for l, since s_n is the last weight increment in s and the final weight assignment is greater or equal than $\bar{l}$. Theorem 2.1 implies that s does not solve I, contradicting the hypothesis and yielding the statement. □

Lemma 2.3. *Let $I =< \mathcal{L} >$ be any MLP instance, $s = (s_1 \dots s_n)$ be any sequence solving I, and $g = (g_1 \dots g_m)$ be the sequence computed by GBA on I, with possibly $n \neq m$. All the loop-constraints met by s_n are also met by g_m.*

Proof. Assume by contradiction that a constraint l is met by s_n, but not by g_m. By Property 2.4, $g_m >^+ \underline{l}$. Hence, for g_m not to meet l, it must be $g_m >^+ \bar{l}$. Also, in order for s_n to meet l, it must exist a component i such that $s_n[i] < \bar{l}[i]$. As a result, $s_n[i] < \bar{l}[i] < g_m[i]$. This contradicts Lemma 2.2, hence yielding the statement. □

Theorem 2.5. *GBA computes a minimal sequence solving any given MLP instance.*

Proof. Consider an MLP instance $I =< \mathcal{L} >$. Let $s = (s_1 \dots s_n)$ and $g = (g_1 \dots g_m)$, with $n \leq m$, be respectively any minimal solution of I and the sequence computed by GBA on I. If $m = 1$, n must be equal to 1 as well, and the statement directly follows. Otherwise, we know by Lemma 2.3 that if g_m meets a set $\mathcal{L}_1 \subseteq \mathcal{L}$ of loop constraints, then s_n meets a subset of constraints in $\mathcal{L}_1$. Consider now the sequences $(s_1 \dots s_{n-1})$ and $(g_1 \dots g_{m-1})$. Again by Lemma 2.3, g_{m-1} meets at least the same set of constraints as s_{n-1}. This implies that the sequence $(g_{m-1}\ g_m)$ meets at least the same constraints met by $(s_{n-1}\ s_n)$. By iterating the same argument, we can show that $(g_{m-n+1} \dots g_m)$ meets at least the same set of constraints as $(s_1 \dots s_n)$. Thus, by definition of s, $(g_{m-n+1} \dots g_m)$ meets all the constraints in $\mathcal{L}$. Also, Property 2.7 ensures that GBA stops at g_{m-n+1}. Hence, it must be $m - n = 0$ and $|g| = |s|$, yielding the statement. □

Theorem 2.4 and 2.5 respectively prove the safety and minimality of the weight sequences computed by GBA. These imply that, among the set of minimal positive weight increment sequences, there exists at least one that is always increasing. In other words, restricting to always increasing sequences does not limit our ability to optimally solve the node shutdown problem, as long as only positive increments are considered. Considering negative increments (within the limit of positive IGP weights), in particular at the beginning of a sequence, would enable to re-balance the traffic between the outgoing links of the router to be shut down. This could allow for additional constraint combinations, thus leading to shorter sequences. However, it might not be realistic from a practical networking perspective, as it would imply rerouting traffic through links that may not be provisioned enough, and possibly even attract more traffic into a router that we want to shut down. Besides, we show in the next section that modifying the set of next-hops used by the modified router may lead to another kind of inconsistencies.

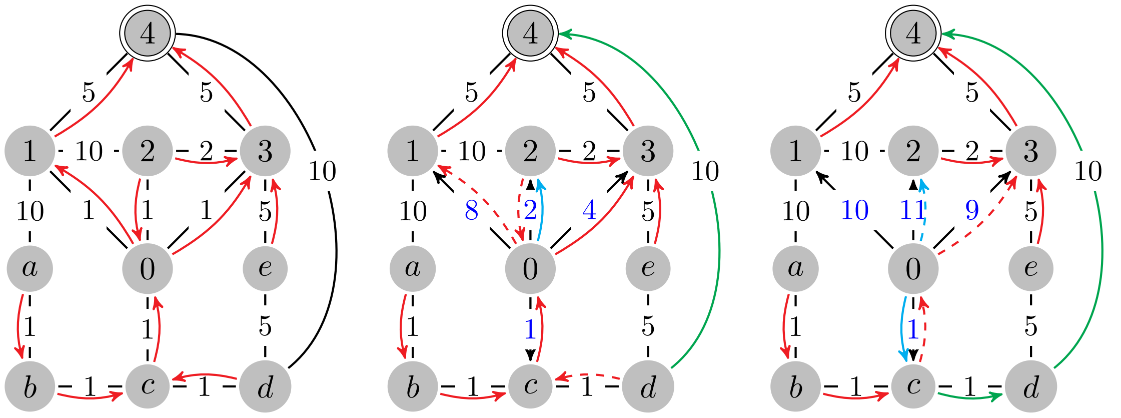
3 Preventing disruptions caused by intermediate updates

Although permitting to further reduce the length of a sequence, the opportunity of simultaneously applying non-uniform weight updates to several links comes with a new kind of transient inconsistencies to deal with. Contrary to uniform updates, which modify the distance towards each destination without affecting the attractiveness of any outgoing link of node 0, heterogeneous ones may change its set of next-hops with each new update. Such phenomenon leads to traffic flows being repeatedly disturbed as successive diversions are applied to their routes. Indeed, route diversions increase the probability of out-of-order packet delivery, in addition to delay and time-to-live (TTL) variations, which may have a negative impact on connection based transport layer protocols. Moreover, these inconsistencies may also cause transient forwarding loops involving edges that are used neither in the initial nor in the final routing graph; hence not appearing when merging the RSPDAGs.

Let us illustrate this problem with a simple example. On the network represented on Fig. 2.10, we consider the shutdown operation of router 0 and the associated sequence computed by GBA.

$$S_{GBA} = \left\{ \begin{pmatrix} 7 \\ 1 \\ 3 \\ 0 \end{pmatrix}, \begin{pmatrix} 9 \\ 10 \\ 8 \\ 0 \end{pmatrix} \right\}$$

In order to prevent a potential transient loop between routers c and d, the first vector of this sequence forces d to shift to its final path, while c still follows its initial one. However, this weight increment also has an impact on the routing decisions of node 0, forcing it to update its shortest paths to 4 (Fig. 2.10b). More precisely, 0 starts using



(A) Initial paths towards 4 (B) Weight increment of (7 1 3 0) (C) Weight increment of (9 10 8 0)

FIGURE 2.10: Illustration of intermediate inconsistencies for destination 4.

nodes 2 and 3 instead of 1 and 3 as next-hops, and forwarding traffic on path (0 2 3 4), which it does not use either in G nor in G'. Note that, contrary to final paths that are expected to be used after the modification, such an intermediate path may not be sufficiently provisioned to carry new flows, hence leading to congestion issues. In this example, node 3 may act as a bottleneck on the paths used by 0 to 4, which are no longer disjoint. Even worse, a transient loop can occur between 0 and 2, since 2 was initially using 0, as highlighted by the red arrow from 2 to 0. Such a loop can also appear during the intermediate convergence resulting from the application of second vector (Fig. 2.10c). Indeed, the weight increment makes router 0 use c as a next-hop towards 4, while the edge $(c, 0)$ was used in the previous state to reach the same destination.

We define as *intermediate forwarding change* such a modification in the set of forwarding paths used by 0, which does not coincide with its initial or final set of paths. In certain cases, intermediate forwarding changes may cause *intermediate transient loops*, as the ones presented on Fig. 2.10. Those loops always include router r and depend on the shortest paths on intermediate forwarding graphs obtained by applying non-uniform weight increments. As such, they do not correspond to cycles in the merged graph $RSPDAG(d) \cup RSPDAG'(d)$. Due to their nature, intermediate transient loops induce two complications. First, they are not considered by GBA, as shown by the example in Fig. 2.10. Second, they map to *dynamic constraints* depending on the increment sequence itself, as opposed to GBA constraints that can be computed through a static analysis on the initial and final RSPDAG.

In this section, we propose several solutions to deal with intermediate forwarding changes and transient loops. In Sec. 3.1, we describe our Adjusted Greedy Backward Algorithm (AGBA) that computes provably minimal sequences preventing all kinds of intermediate disruptions. In Sec. 3.2, we present a heuristic algorithm, called Dynamic Greedy Backward Heuristic (DGBH), which only focuses on intermediate transient loop prevention in order to provide shorter sequences. Finally, in Sec. 3.3, we discuss a technical alternative to DGBH, preventing intermediate transient loops while using standard GBA sequences.

3.1 Algorithmic solution to prevent intermediate forwarding changes

Since the root cause of intermediate next-hops leading to loops and new forwarding paths is induced by local changes on node r, a sufficient condition to avoid any intermediate edge consists in enforcing that r *maintains* its initial next-hops throughout the IGP convergence.

As a preamble to the description of our intermediate change prevention mechanism, let us introduce some new notations and definitions. We denote the component of a vector v associated to a link (r, x) as $v[x]$.

Definition 3.1. A node s is called **initial successor** of r to d if (r, s) is the first edge of a path in $RSPDAG(d, G)$. We denote the set of initial successors of r to d as $S^*(d)$.

Intuitively, initial successors are next-hops used by r to reach d in G. In the example in Fig. 2.10, nodes 1 and 3 are initial successors of r for destination 4, while 2 and c are not.

Definition 3.2. Let d be a destination and x a neighbor of router r. We define the *offset value* of x towards d as

$$\mathit{offset}^r_d(x) = C(r, x, d) - C(r, d)$$

where $C(r, x, d)$ represents the cost of the shortest elementary path in G from r to d via link (r, x).

This offset value reflects the *attractiveness* of a neighbor compared to the initial next-hops of r to d. Note that, if $x \in S^*(d)$ then $C(r, x, d) = C(r, d)$ and $\mathit{offset}_d(x) = 0$. The purpose of such an offset is to retrieve the *distance increment* towards d represented by a vector component. Indeed, the less a neighbor is attractive the lower the corresponding component of a vector would have to be, in order to increase the distance from r to d by any given value.

Definition 3.3. Let d be a destination, s^* be an initial successor of r to d, and v be a weight increment. We define the *intermediate forwarding Change Prevention Conditions (CPCs)* as the set of inequalities

$$v[s] = v[s^*]$$

$$v[x] > v[s^*] - offset_d(x)$$

for each initial successor $s \in S^*(d)$ of r, and for each other neighbor x of r such that $x \notin S^*(d)$.

As an illustration, consider again Fig. 2.10 and let $s^* = 1$. The CPCs for destination 4 consists of inequalities $v[1] < v[2]+2$ and $v[1] = v[3]$. Observe that CPCs are formulated with respect to a single initial successor (i.e., 1 in the example above). However, the correctness of the CPCs does not depend on the considered initial successor.

Moreover, for each neighbor $x \notin S^*(d)$, it must be $C(r, d) < C(r, x, d)$ by definition of initial successors. Hence, $offset_d(x) > 0$, and the following property holds.

Algorithm 2 AGBA-1 : Minimal Constraints Initialization

```
function AGBA_INIT($n, D, offset$)
    $M \leftarrow \varnothing$                                        ▷ Minimal constraints matrix
    for $d$ in $N$ do
        $S^* \leftarrow \varnothing$                                  ▷ Initial next-hops of router $n$
        for $x$ in $n$.succ() do
            $offset[d][x] \leftarrow w(n, x) + C'(x, d) - C(n, d)$
            if $offset[d][x] = 0$ then
                $S^*.append(x)$
            end if
        end for
        for $s$ in $S^*$ do
            for $x$ in $n$.succ() do
                $M[s][x] \leftarrow$ min $(M[s][x], offset[d][x])$
            end for
        end for
    end for
end function
```

Algorithm 3 AGBA-2 : Greedy Vector Adjustment

```
function AGBA_ADJUST($M, n, gv$)
    $indexes \leftarrow n$.succ()
    while $indexes \neq \emptyset$ do
        $p \leftarrow$ pop_max_index $(indexes, gv)$
        for $x$ in $n$.succ() do
            if $M[p][x] = 0$ then
                $gv[x] \leftarrow gv[p]$
            else if $gv[x] \leq gv[p] - M[p][x]$ then
                $gv[x] \leftarrow gv[p] - M[p][x] + 1$
            end if
        end for
    end while
    return $gv$
end function
```

Property 3.1. *Any CPC inequality can be written as* $v[s^*] \leq v[x] + m$*, with* $m \geq 0$*.*

Intuitively, CPCs impose that, for a given destination, paths via initial successors of r should be shorter than any other paths via a non initial successor. That is, we aim to control weight increments such that no intermediate forwarding change occur. Hence, verifying CPCs for a destination d guarantees that the shortest paths from r to d remain the same. This implies the following theorem.

Theorem 3.4. *If a weight increment* v *satisfies the CPCs for all destinations, no forwarding change occurs when* v *is applied.*

Proof. Assume by contradiction that a forwarding change occurs for a destination d when v is applied, even if v verifies all the CPCs for d. By definition of forwarding change, a node $\bar{x}$ must exist such that one of its shortest paths to d after the application of v is not included either in the initial nor in the final ones. Since only the weights of the links outgoing from r are changed by v, the paths from $\bar{x}$ to r are the same as the initial ones. This means that r must also have changed its shortest paths to d.

By definition of CPCs, all the paths from r to d via initial successors have the same length after the application of v. Thus, for a forwarding change on r to occur, there must exist a path $(r\ x \dots d)$ shorter than or equal to the shortest paths from r to d via any initial successor s^*. Since only the weights of the links outgoing from r are changed by v, this means that $v[s^*] + C(r,d) \geq v[x] + C(r,x,d)$, i.e., $v[s^*] \geq v[x] - offset_d(x)$. This inequality contradicts the hypothesis that all CPCs are verified by v, thus proving the statement. □

Since intermediate transient loops cannot occur in the absence of forwarding changes, the following corollary holds.

Corollary 3.5. *If a weight increment v satisfies the CPCs for all destinations, no intermediate transient loop occurs.*

We now present a variation of GBA, called AGBA, that guarantees prevention of intermediate edges by enforcing accommodation of CPCs for all network destinations. More precisely, AGBA solves the following problem.

Problem 3.1. *Minimal intermediate Change-free and Loop-free Problem (MCLP): Given a set $\mathcal{L}$ of loop-constraints and a set $\mathcal{A}$ of CPCs, compute a minimal weight increment sequence that contains no unsafe transition for any constraint in $\mathcal{L}$, and no weight increment that violate any condition in $\mathcal{A}$.*

Provided that all the loop-constraints and the CPCs are correctly computed, solving an MCLP instance implies preventing all possible convergence loops and forwarding changes in the corresponding network as per Theorems 2.1 and 3.4.

To solve the MCLP problem AGBA post-processes each weight increment gv as computed by GBA. To this end, AGBA adds two main algorithmic steps to each iteration of GBA. One in its initialization, the other within the main loop iteration to adjust the greedy vector.

First, AGBA computes every offset values and optimizes them across all destinations, as shown in Alg. 2. In particular, for each destination, it computes all the offsets and identifies the initial successors (ll. 5–8 in Alg. 2). Moreover, for each pair initial successor and neighbor of r, it only keeps the smallest offset (ll. 9-11 in Alg. 2), as it corresponds to the most constraining CPCs.

Second, AGBA modifies the greedy vector gv as computed by GBA, applying the following operations. 1) *vector sorting*, in which the components of gv are considered from the biggest to the smallest one (this corresponds to consider all the CPCs in decreasing order). The goal is to retrieve the up to date pivot component p (line 4 in Alg. 3); and 2) *vector adjusting*, in which the current component of gv is modified to satisfy all the sorted CPCs. AGBA enforces the CPCs by imposing:

$$v[s] = m_d$$

$$v[x] = m_d - offset_d(x) + 1$$

where $s \in S^*(d)$, $x \notin S^*(d)$, and $m_d = max_{\forall s \in S^*(d)}(v[s])$. That is, given a weight increment, AGBA calculates the maximum component corresponding to an initial successor, which we call *pivot component*, and imposes that all the other components of the vector must enforce the CPCs with respect to such a pivot component. We formally define this particular component as follows.

Definition 3.6. Given a vector v and a set of CPCs, we denote as *pivot component* the largest component of v appearing in the left hand side of any unsatisfied CPC inequality.

Consider again the example in Fig. 2.10. The pivot component of the shown weight increment v is $v[1]$ and $m_4 = 7$. AGBA imposes that $v[1] = v[3] = 7$, $v[2] = 6$ and $v[c] = 2$. Eventually, the complete sequence computed by AGBA on the network in the figure is

$$S_{AGBA} = \left\{ \begin{pmatrix} 3 \\ 2 \\ 3 \\ 3 \end{pmatrix}, \begin{pmatrix} 7 \\ 6 \\ 7 \\ 7 \end{pmatrix}, \begin{pmatrix} 8 \\ 7 \\ 8 \\ 8 \end{pmatrix}, \begin{pmatrix} 9 \\ 10 \\ 9 \\ 9 \end{pmatrix} \right\}$$

As shown on Fig 2.11, this sequence has no impact on the routing decisions of node 0. Remember that weight increments only apply on the outgoing links of 0. In particular, the cost of edge $(0, c)$ is not modified by the sequence. For destination 4, the first vector only makes node d reroute through link $(d, 4)$ (Fig. 2.11b), while node c alone updates its next-hop with the second one (Fig. 2.11c). The two subsequent vectors have no impact at all on the shortest paths towards 4.

This sequence is two vectors longer than the one produced by GBA, yet one shorter than a minimal uniform sequence $\{3, 7, 8, 9, 10\}$ for the same operation. Although it may appear a large sequence length overhead in this particular case, our experiments show that sequence stretching remains marginal in many realistic situations (see Ch. 3).

Correctness and optimality proofs

Intuitively, AGBA is correct and optimal because CPCs are static conditions, in the same vein as transient loop-constraints. The greedy behavior of GBA is then still ensured with respect to an additional kind of static constraints, i.e., the "minimal" resolution of a linear inequality system.

In our proofs, we use the term *before* iteration j to denote all iterations that are lower than j. We also say that a constraint is *unsatisfied* (resp., *satisfied*) at an iteration j if it is not met (resp., it is met) by any vector computed by AGBA before j. For simplicity, we restrict to the case of a single *pivot component* per vector. However, lemmas and theorems can be easily generalized to multiple pivot components.

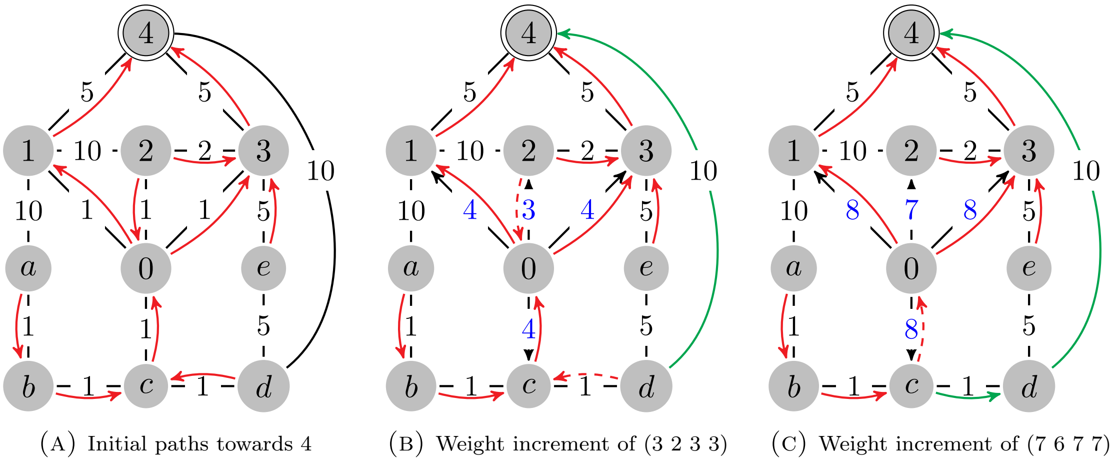
AGBA being a variation of the GBA, it inherits of the same properties, but for some minor modifications. Besides, the proofs of two statements related to minimality are also exactly the same as GBA.

Property 3.2. *At each iteration j, AGBA computes a vector v such that $v >^+ \underline{l}$ for all the constraints $l = (\underline{l}, \bar{l})$ still unsatisfied before j.*

Property 3.3. *Given any AGBA iteration j, let v be the vector that AGBA computes before the adjusting phase. For each component i of v, a constraint $l = (\underline{l}, \bar{l})$ still unsatisfied before j exists such that $v[i] = \max(\underline{l}[i] + 1, 0)$.*

Property 3.4. *AGBA computes always increasing sequences.*

Property 3.5. *AGBA stops as soon as all the constraints are met.*



(A) Initial paths towards 4 (B) Weight increment of (3 2 3 3) (C) Weight increment of (7 6 7 7)

FIGURE 2.11: Illustration of AGBA sequence for destination 4.

Properties 3.2 and 3.3 are ensured by the greedy vector computation, plus the fact that AGBA only increases some components of the vector during the adjusting phase. Property 3.4 is the result of both vector computation and constraint removal. Properties 3.5 derives from the constraint removal mechanism.

First of all, we show that AGBA always terminates.

Lemma 3.1. *For any AGBA iteration, the pivot component of the computed vector remains bigger than any component appearing in the left side of any CPC inequality during the adjusting phase.*

Proof. The statement hold at the beginning of the adjusting phase by definition of pivot component.

Now, assume by contradiction that the statement holds until a given step s during the adjusting phase, but not after s. That is, at step s AGBA computes a vector in which at least one component m is bigger than the pivot component j, and m appears in the left side of some CPCinequalities. Let w and z be the vectors computed by AGBA respectively before and after step s. By hypothesis, $\forall i \; w[i] \leq w[j]$ while $z[m] > z[j]$.

This hypothesis implies that AGBA has increased the m-th component of w at step s. By definition, AGBA increases a component only if it appears in the right side of an CPC inequality. Thus, the inequality considered by AGBA in s must be $v[l] \leq v[m] + k$, with $k \geq 0$ and $l \neq m$. To accommodate this inequality, by definition, AGBA enforces $z[l] = z[m] + k$, that is, $z[l] \geq z[m]$ since $k \geq 0$. All the other components are left unmodified, hence $z[l] = w[l]$ and $z[j] = w[j]$.

We have two cases. If $l = j$, then $z[l] \geq z[m]$ implies $z[j] \geq z[m]$, which contradicts the definition of z. Otherwise, if $l \neq j$, then it must be $w[l] = z[l] \geq z[m] > z[j] = w[j]$, i.e., $w[l] > w[j]$, which contradicts the definition of w. In both cases, we contradict the hypothesis, which yields the statement. □

Lemma 3.2. *The pivot component is never modified by AGBA during the adjusting phase.*

Proof. Let p be any vector before the adjusting phase, and let $p[j]$ be its pivot component. We now show that AGBA never modifies $p[j]$.

In the adjusting phase, AGBA iterates once on the sorted set of CPC inequalities, considering one inequality at the time and increasing some components of the vector if needed. Consider any step s in this iteration. Let w be the vector at the beginning of s. The following cases apply to the CPC inequality that AGBA considers at s.

- $v[j]$ does not appear in the inequality, hence it is not modified, by definition of AGBA.

- $v[j]$ appear in the left side of the inequality, which has the form $v[j] \leq v[i] + k$, with $i \neq j$ and $k \geq 0$. If the inequality is satisfied, AGBA will not modify any component of w. Otherwise, by definition, AGBA will only increase the value of $w[i]$ while not modifying $w[j]$.

- $v[j]$ appear in the left side of the inequality, which has the form $v[l] \leq v[j]+q$, with $l \neq j$ and $q \geq 0$. By Lemma 3.1, it must be $w[l] < w[j]$. Hence, the inequality is already satisfied by w, and by definition, AGBA does not modify any component of w.

In all the cases, AGBA does not modify $w[j]$. The statement follows by applying the same argument to all the steps performed by AGBA during the adjusting phase. □

Lemma 3.2 implies that at least one constraint is satisfied by AGBA at each step.

Lemma 3.3. *At each iteration, AGBA computes a vector v that meets at least one constraint not met before.*

Proof. Consider any AGBA iteration i. Let $\mathcal{L}$ be the set of unsatisfied constraints at the beginning of i, let v be the computed vector before the adjusting phase, and let $v[j]$ be its pivot component. By Property 3.2 and 3.3, one constraint $l = (\underline{l}, \bar{l}) \in \mathcal{L}$ must exist such that $v >^{+} \underline{l}$ and $v[j] = \underline{l}[j] + 1$. By Property 2.1, it must also be $v[j] < \bar{l}[j]$, that is, l is met by v. The statement follows by noting that $v[j]$ is unmodified by AGBA in the adjusting phase, by Lemma 3.2. □

We now leverage Lemma 3.3 to prove that AGBA always terminates in a finite number of iterations.

Theorem 3.7. *For any MCLP instance $I =< \mathcal{L}, \mathcal{A} >$, AGBA always terminates in $O(|\mathcal{L}|)$ iterations.*

Proof. By Property 3.5, AGBA stops when all the initial constraints $\mathcal{L}$ are met. Hence, the statement directly follows by Lemma 3.3. □

We now show that the sequences computed by AGBA are guaranteed to be safe and to avoid intermediate edges.

Lemma 3.4. *In AGBA, adjusting a vector according to a given CPC does not invalidate previously satisfied CPCs.*

Proof. Assume by contradiction that AGBA invalidates a previously satisfied CPC inequality (1) $v[l] \leq v[m] + k$ to satisfy another CPC inequality (2) $v[i] \leq v[j] + q$. By definition of CPC, $k, q \geq 0$. Let w and z be the vectors computed during the adjusting phase immediately before and immediately after considering (2), respectively. Our assumption translates to having $w[l] \leq w[m] + k$, $z[i] \leq z[j] + q$, and $z[l] > z[m] + k$. One of the following cases must hold.

- w is already compliant with (2). Then, by definition, AGBA does not modify any component of the current vector w, hence $z[l] = w[l]$ and $z[m] = w[m]$. By definition of w, this means that it must be $z[l] \leq z[m] + k$, which contradicts the assumption.
- w is not compliant with (2) and $j \neq l$. By definition, AGBA only increases the j-th component of w, i.e., $z[j] > w[j]$ but $z[l] = w[l]$ and $z[m] = w[m]$. By definition of w, this implies that $z[l] \leq z[m] + k$, which contradicts the assumption.
- w is not compliant with (2) and $j = l$. Since (1) has been considered by AGBA before (2), then it must be $w[l] = w[j] > w[i]$ by definition of the sorting phase in AGBA. This means that (2) has been already satisfied by w, contradicting the hypothesis of this case.

All cases lead to a contradiction, yielding the statement. □

Theorem 3.8. *The weight sequences computed by AGBA prevent both transient loops and intermediate edges.*

Proof. Let I be any MCLP instance, and let s be the sequence computed by AGBA on I. By Property 3.4, s is an always increasing sequence. By Lemma 3.3, each constraint is met by at least one vector in s. Thus, Theorem 2.2 ensures the prevention of transient loops. Moreover, by definition of the AGBA adjusting phase and by Lemma 3.4, all the CPCs inequalities are satisfied by each weight increment in s. Hence, Theorem 3.4 guarantees the prevention of intermediate edges. □

Finally, we prove the minimality of the sequences computed by AGBA.

Lemma 3.5. *Let I be any MCLP instance, $s = (s_1 \dots s_n)$ be any sequence solving I, and $g = (g_1 \dots g_m)$ be the sequence computed by AGBA on I, with possibly $n \neq m$. The last vectors of the sequences verify $s_n \geq g_m$.*

Proof. Let $I =< \mathcal{L}, \mathcal{A} >$. Assume by contradiction that $s_n[i] < g_m[i]$ for a given component i. We have two cases.

- *AGBA did not modify the i-th component in the adjusting phase* of its first iteration. Then, by Property 3.3, there must exist at least one constraint $l = (\underline{l}, \overline{l}) \in \mathcal{L}$ such that $g_m[i] = \underline{l}[i] + 1$. This implies that $s_n[i] < \underline{l}[i] + 1$, hence l is not met by s_n.

- *AGBA modified the i-th component in the adjusting phase* of its first iteration. Then, by definition of adjusting phase, $\mathcal{A}$ must include an CPC inequality $v[j] \leq v[i] + y$. Since i-th component was actually adjusted by hypothesis, it must be $y > 0$, and AGBA enforced $g_m[j] = g_m[i] + y$, i.e., $g_m[j] < g_m[i]$. Moreover, for s to prevent intermediate edges, $s_n[j] \leq s_n[i] + y$. Since $s_n[i] < g_m[i]$ by hypothesis, it must be $s_n[j] < g_m[i] + y$, hence $s_n[j] < g_m[j]$.

In the second case, we can iterate the argument above starting from j-th component. Each time the second case applies, we end up with a component of g_m strictly bigger than the previously considered one. Thus, the second case can hold until we reach the biggest component of g_m that appears in the left side of any CPC inequality. By Lemma 3.1, this component is the pivot component. Thus, Lemma 3.2 ensures that the first case eventually applies.

Hence, at least one constraint l is not met by s_n. This means that s contains an unsafe transition for l, since s_n is the last weight increment in s and the final weight assignment is greater or equal than $\overline{l}$. Theorem 2.1 implies that s does not solve I, contradicting the hypothesis and yielding the statement. □

Lemma 3.6. *Let I be any MCLP instance, $s = (s_1 \dots s_n)$ be any sequence solving I, and $g = (g_1 \dots g_m)$ be the sequence computed by AGBA on I, with possibly $n \neq m$. All the loop constraints met by s_n are also met by g_m.*

Proof. Same as Lemma 2.3, using Property 3.2 and Lemma 3.5. □

Theorem 3.9. *AGBA computes a minimal sequence solving any given MCLP instance.*

Proof. Same as Theorem 2.5, using Lemma 3.6 and Property 3.5. □

Note that AGBA is minimal even if the MCLP instance allows relative weight modification in $\mathbb{Z}$ while GBA ensures minimality for the MLP only if globally relative weight modifications stays in $\mathbb{N}$.

3.2 Algorithmic solution to prevent intermediate transient loops

AGBA enforces strong consistency guarantees during IGP convergence at the cost of increasing the sequence length. In this section, we explore a trade-off between routing consistency and sequence length. In particular, we investigate the opportunity of relaxing the CPCs, allowing for intermediate next-hop changes as long as they do not lead to transient loops.

As opposed to CPCs that only depend on the initial routing state, our new requirements are defined for a transition, which may either be between two intermediate vectors, or between a vector and an extremity state. In both cases, the conditions to be satisfied depend on the intermediate vector themselves. Intuitively, let us consider a weight increment sequence $s = \{s_0, \ldots, s_{i-1}, s_i, \ldots, s_m\}$, such that the application of vector s_i triggers an intermediate next-hop change that would lead to an intermediate transient loop, with respect to the routing state induced by vector s_{i-1}. On one hand, it is possible to modify s_i to prevent the intermediate forwarding change, so that no loop occur when applying the vector. This is the most intuitive solution, as it follows in the same spirit as AGBA. However some constraints initially satisfied at step s_i with GBA can be *shifted* to next iterations $\leq i-1$ such that it may possibly increase the sequence length. On the other hand, we could also deal with the intermediate loop as a normal transient one, i.e. by encoding it as a static loop-constraint. In practice, it will then consist in forcing s_{i-1} to be larger than the lower bound of the associated constraint.

In order to explore this new compromise, we first need to characterize intermediate transient loops. By definition, such a loop may occur due to an intermediate forwarding change, that is, when an intermediate weight increment makes node r consider as next-hop a neighbor that is not part of its initial or final set of next-hops. However, a transition making node r use an intermediate next-hop does not necessarily lead to an intermediate transient loop. Indeed, some neighbors of r may be considered *safe*, allowing node r to use them as next-hops without triggering intermediate transient loops. For instance, neighbors that do not reach d through node r before the change cannot be involved in any transient loop for this destination, thus being *safe* according to this definition. Since all other neighbors initially reach destination d through r, their safety status may depend on the transition, in particular their state before and after the transition. We thus define the circumstances leading GBA to result in intermediate transient loops as follows.

Definition 3.10. Let d be a destination and, v_1, v_2 be two weight increment vectors such that $v_1 < v_2$. An intermediate transient loop can occur for destination d during a transition between v_1 and v_2, if v_2 makes node r consider as next-hop towards d a neighbor that is not in its final state after v_1.

Let us denote as p a neighbor of r meeting the condition described in the definition above. If p is not in its final state after v_1, there necessarily exists in this state a shortest path $P_1 = (p, \dots, r, \dots d)$. Besides, if v_2 makes r consider p as a next-hop towards d, a shortest path $P_2 = (r, p, \dots, d)$ exits in this second state. Hence, a transient loop $(p, \dots, r, p)$ may indeed occur during the transition.

Let us now introduce some new notations and definitions for this problem.

Definition 3.11. A node p is called **initial predecessor** of r to d if the edge (p, r) is in $RSPDAG(d, G)$. We denote the set of initial predecessors of r as $P_r^*(d)$.

In addition, we extend the distance notations $C(x, d)$ and $C(x, l_i, d)$ to consider intermediate graphs resulting from the application of a weight increment v, using respectively $C_v(x, d)$ and $C_v(x, l_i, d)$.

Definition 3.12. Let d be a destination and v_1, v_2 be two weight increments such that $v_1 < v_2$. We define the Dynamic Loop Constraints (DLCs) as follows:

$$\forall\ p^* \in P^*(d) \mid C_{v_2}(r, p^*, d) = C_{v_2}(r, d), \quad v_1 >^+ \Delta_d(p^*)$$

In order to prevent intermediate transient loops, such DLCs are to be satisfied by each vector in the sequence. Should we consider that the larger vector, v_2, cannot be modified, a DLC could be compared to a static loop-constraint $l = (\Delta_d(p^*), v_2)$.
In the following, we denote the set of dynamic loop-constrants related to a given vector as $\mathcal{C}$ and individual constraints as c. Besides, we define as follows the *minimal satisfaction* of a CPC.

Definition 3.13. If a vector v_1 verifies $v_1 >^+ \Delta_d(p^*)$ and $\exists\ i \mid v_1[i] = \Delta_d(p^*)[i] + 1$, we say that the vector $v_1 < v_2$ minimally satisfies the dynamic loop-constraint $c = (\Delta_d(p^*), v_2)$.

Minimal CPC satisfication represents a sufficient property for the lower bound of a dynamic constraint to no longer appear at any subsequent iteration of a GBA-based algorithm. That is, if a vector v_1 minimally satisifies a CPC $c = (\Delta_d(p^*), v_2)$, no intermediate transient loop involving both p^* and r could occur for any transition between vectors lower than v_1. Indeed, no lower weight increment could make r consider p^* as a next-hop towards d: p^* is no longer to be considered as a lower bound of a DLC.

More generally, DLCs impose that any initial predecessor of r, for a given destination d, be in its final state at least one step before being used by r on a shortest path towards d. These constraints could be satisfied by increasing the values in v_1 as mentioned above, but also by adjusting the values in v_2 to ensure to no initial predecessor of r that is not in its final state after v_1 is used on a shortest path. One way or the other, verifying the DLCs for two weight increment vectors guarantees that no intermediate transient loop could occur during the transition. This implies the following theorem.

Theorem 3.14. *If two weight increments v_1 and v_2, such that $v_1 < v_2$ satisfy the DLCs for all destinations, then no intermediate transient loop may occur when v_2 is applied after v_1.*

Proof. Assume by contradiction that an intermediate transient loop may occur for a destination d when v_2 is applied, even though v_1 and v_2 satisfy the DLCs. By definition 3.10, an initial predecessor p^* of r must exist such that r is on a shortest path from p^* to d before the convergence, while the opposite holds after the change. This means that p^* is not in its final state before v_2 is applied, that is, after v_1.
By definition of DLCs, if a shortest path from r to d include p^* after v_2 is applied, then $v_1 >^+ \Delta_d(p^*)$. Thus, by definition of delta vectors, p^* is in its final state after v_1. This contradicts the hypothesis, and yield the statement. □

Although we can show that any sequence computed using AGBA satisfies the DLCs, such sequence is not necessarily of minimal length with respect to these conditions. We thus define as follows the problem of optimally satisfying the DLCs.

Problem 3.2. *Minimal Intermediate Loop-free Problem (MILP): Given a set $\mathcal{L}$ of loop-constraints, compute a minimal weight increment safe sequence that does not result in any intermediate transient loop on* 0.

Finding such a minimal sequence, which satisfies conditions based of the sequence itself, seems a complex problem. An algorithm similar to AGBA would have to adjust the values of intermediate vectors in order to prevent intermediate transient loops for every transition. The difference compared to AGBA is that DLCs depend on the state before the transition, which can be induced by a previous vector, rather than an initial, fixed, state. It is thus possible to satisfy these conditions by modifying either the current vector, the previous one, or both. A choice has to be made, which could impact the length of the resulting sequence. Modifying the values of an intermediate weight vector, in order to prevent intermediate transient loops for a given transition, might lead to more intermediate transient loops to deal with at the next transition or many *static constraints* being left unsatisfied. Even a choice that appears wise at one step may impact the sequence in such a way that more intermediate steps would be required in the end to satisfy all constraints.

Algorithm 4 DGBH Greedy Vector Adjustment

```
function DGBH_ADJUST($r, d, v_1, v_2$)
    for $i$ in $r$.succ() do
        if $i$.is_final ($v1$) then
            $v_2\_incr =$ min ($v_2\_incr, v2[i] + offset[d][j]$)
        end if
    end for
    for $i$ in $r$.succ() do
        if not $i$.is_final ($v1$) and $v_2[j] + offset[d][j] \leq v_2\_incr$ then
            $v_1\_adj =$ max ($v_1\_adj, \Delta[d][i] + 1$)
        end if
    end for
    if $v_1\_adj > 0$ then
        for $i$ in $r$.succ() do
            $v_1[i] =$ max ($v_1[i], v_1\_adj - offset[d][i]$)
        end for
    end if
end function
```

Exploring every possibility to satisfy the DLCs would no doubt permit to solve the MILP. Such a combinatorial exploration however comes with a theoretical complexity too high to be practical for real life usage. Besides, from a practical point of view the minimality of a sequence matters less than its actual length. For example, knowing that a 50-vectors long sequence is minimal does not make it usable. On the other hand, we could use a sequence containing 4 intermediate vectors, even if it is one vector longer than a minimal one. In this perspective, we designed a heuristic to efficiently compute short weight increment sequences, provably preventing both normal and intermediate transient loops.

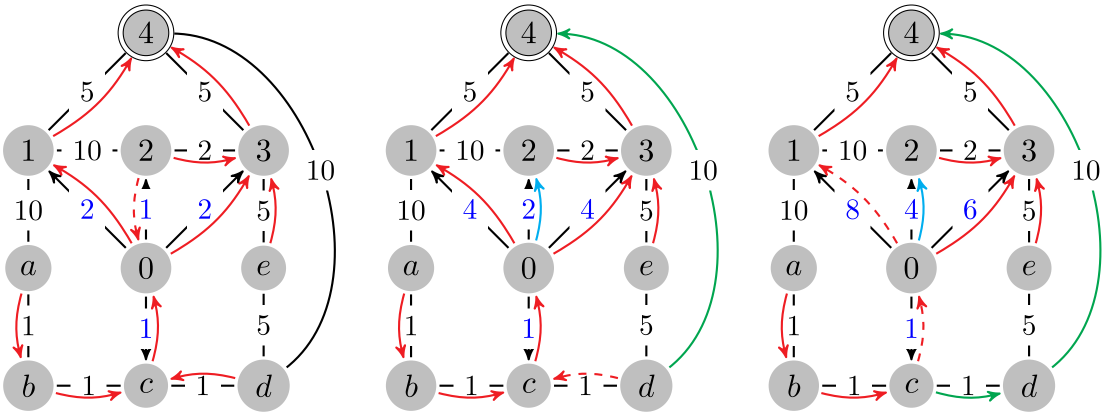
In order to deal with the DLCs, our greedy heuristic, called DGBH, potentially adds dynamic loop constraints at each iteration. In practice, it simply extends GBA to retrieve the dynamic constraints associated to cycles including node r before each greedy vector is computed. This way it computes the lower bounds of last constraints related to intermediate loops. Note that those additional operations neither require any extra information nor dedicated computation process, keeping a time efficiency similar to the original GBA.

Function DGBH_ADJUST, presented in Alg 4, ensures that no intermediate transient loop appears, for destination d, during the transition between two greedy vectors, v_1 to v_2 ($v_1 < v_2$), computed by GBA. Such intermediate transient loops are prevented by considering v_2 as an invariant and modifying the values in v_1 to satisfy the DLC for d. This function first extracts from v_2 the *lowest distance increment* from r to d via any neighbor that is in its *final state* after v_1, formally $min(\{v_2[i]+offset_d(i) : i \in P_r^{v_1}(d)\})$

(lines 2–6).

For practicality purposes, we rely on function $n.\textbf{is\_final}(v)$ to check whether a node n is in its final state after vector v. This could easily be done for every node having a strictly positive delta value by comparing this value with the *distance increment* for v: $min(\{v[i] + offset_d(i) : i \in S_r^v(d)\})$. Such a node is in its final state if the *distance increment* of v is strictly greater than the delta value associated to n. However, this statement does not necessarily hold for nodes having a delta value equal to 0. Although we know for sure that such a node n has at least one shortest path to d not going through r in the initial state, it is unsure whether or not it also has one via r. If it has, then a *distance increment* of at least one is required for it to be in its *final state* and the previous statement holds. On the contrary, if no shortest path from this node to d goes through r, the node is already considered in its *final state*, even before any weight increment is performed.

The second step of DGBH_ADJUST function aims at detecting a potential intermediate transient loop (lines 7–11) by comparing the *distance increment* to d, via neighbors that are not in their *final state* after v_1, with the *lowest distance increment*, via any other neighbor, previously computed. This test makes it possible to retrieve the maximum delta value, plus one, among the neighbors that do not satisfy the DLC, hence involved in potential intermediate transient loops occurring during the transition from v_1 to v_2. The resulting value represent the lowest possible *distance increment* towards d such that every *unsafe* neighbor of r, not satisfying the DLC, is in its final state and does not use r to reach d. The final step thus consists in increasing the values in v_1 so that they carry at least this *distance increment* (lines 12–16). Eventually, every neighbor that would have been involved in a potential intermediate transient loop is forced in its *final state* by v_1, so that no intermediate transient loop may appear during the transition from v_1 to v_2. Note that some of the static constraints previously held by v_1 might be left unsatisfied as its values are increased, and additional intermediate vectors could be required. In particular, if $v_1 = \vec{0}$, the initial weight assignment, a modification of the vector would actually mean a new intermediate vector. In practice, this function is meant to be added in the main GBA loop, as would AGBA_ADJUST, in order to adjust on-the-fly the latest vector computed by GBA. At the first iteration, v_1 would thus be the first computed intermediate vector (or the last to be applied), while v_2 would be the final weight assignment. It is also worth mentioning that, since DLCs of all destinations cannot be aggregated beforehand, this adjustment function is to be called for every destination.



(A) Weight increment of (1 0 1 0)

(B) Weight increment of (3 1 3 0)

(C) Weight increment of (7 3 5 0)

FIGURE 2.12: Illustration of DGBH sequence for destination 4.

Let us illustrate these principles for the removal operation of router 0 from the network on Fig. 2.10, whose associated DGBH sequence is

$$S_{DGBH} = \left\{ \begin{pmatrix} 1 \\ 0 \\ 1 \\ 0 \end{pmatrix}, \begin{pmatrix} 3 \\ 1 \\ 3 \\ 0 \end{pmatrix}, \begin{pmatrix} 7 \\ 3 \\ 5 \\ 0 \end{pmatrix}, \begin{pmatrix} 9 \\ 10 \\ 8 \\ 0 \end{pmatrix} \right\}$$

For destination 4, the first vector in this sequence, (1 0 1 0), prevents the intermediate loop between 0 and 2 by forcing node 2 to stop using 0 as a next-hop (Fig 2.12a). Hence, the intermediate forwarding change triggered by the second vector cannot lead to a loop (Fig 2.12b). The convergence to the third vector is also safe with respect to intermediate transient loops, because the weight increments on links $(0, 2)$ and $(0, 3)$ are not sufficient for node 0 to reroute towards c (Fig 2.12c).

Proof of correctness

We now prove the correctness of DGBH as a heuristic for the MILP problem. That is, we show that our algorithm produces a sequence of intermediate weight increments preventing both transient loops and intermediate transient loops, which is not necessarily of minimal length.

Our proofs leverage the following properties of DGBH that hold by definition of the algorithm.

Property 3.6. *At each iteration j, DGBH computes a vector v such that $v >^{+} \underline{l}$ for all the constraints $l = (\underline{l}, \bar{l})$ still unsatisfied before j.*

Property 3.7. *Given any iteration j, let v be the vector that DGBH computes before the adjusting phase. For each component i of v, a constraint $l = (\underline{l}, \overline{l})$ still unsatisfied before j exists such that $v[i] = \max(\underline{l}[i] + 1, 0)$.*

Property 3.8. *DGBH computes always increasing sequences.*

Property 3.9. *DGBH stops as soon as all the static constraints are met and the set of dynamic constraint is empty between the initial state and the first vector.*

Property 3.10. *Given any iteration j, let v be the vector that DGBH computes before the adjusting phase, and $v' > v$ be the modified version of v after the adjustment phase. v'* minimally satisfies *at least one DLC.*

Properties 3.6 to 3.9 are ensured by the greedy behavior of DGBH, the vector computation and constraints removal, plus the fact that DGBH only increases the vector during the adjusting phase. Property 3.10 derives from lines 9 and 14 of Alg. 4. It allow us to prove the progression of DGBH on DLCs.

Let us now prove that DGBH terminates and produces sequences that satisfy the CPCs.

Lemma 3.7. *At each iteration, DGBH computes a vector v that meets at least one static or dynamic constraint not met before.*

Proof. Consider any DGBH iteration j. Let $\mathcal{L}$ be the set of unsatisfied static constraints at the beginning of j, let $\mathcal{C}$ be the set of DLCs at j, and let v be the vector computed after the adjusting phase. We differentiate two cases.

- If $\mathcal{C}$ is empty, v is not modified during the adjustment phase. By Properties 3.6 and 3.7, one constraint $l = (\underline{l}, \overline{l}) \in \mathcal{L}$ must exist such that $v >^+ \underline{l}$ and $\exists\ i \mid v[i] = \underline{l}[i] + 1$. By Property 2.1, it must also be $v[i] < \overline{l}[i]$, that is, l is met by v.
- If $\mathcal{C}$ is not empty, v minimally satisfies at least on dynamic constraint by Property 3.10.

These two cases prove the statement. □

We now leverage Lemma 3.7 to prove that DGBH always terminates in a finite number of iterations.

Theorem 3.15. *For any MILP instance $I =< \mathcal{L} >$, DGBH always terminates in $O(|\mathcal{L}| + k \times |N|)$ iterations.*

Proof. By Property 3.9, DGBH stops when all the static constraints in $\mathcal{L}$ are met and there is no dynamic constraints between the initial state and the first weight increment. Since there is at most as many dynamic constraints to be met as the number of delta vectors for all neighbors of r, the statement directly follows by Lemma 3.7. □

Theorem 3.16. *The weight sequences computed by DGBH prevent both normal and intermediate transient loops.*

Proof. Let I be any MILP instance, and let s be the sequence computed by DGBH on I. By Property 3.8, s is an always increasing sequence. By Lemma 3.7, each constraint is met by at least one vector in s. Thus, Theorem 2.2 ensures the prevention of transient loops. Moreover, by definition of the DGBH adjusting phase all the DLCs are satisfied by each weight increment in s. Hence, Theorem 3.14 guarantees the prevention of intermediate transient loops. □

While sequences computed by DGBH are correct, they are not guaranteed to be minimal. In fact, considering again the removal of 0 on Fig. 2.10, a sequence of the same length as GBA's exists that does not incur intermediate transient loops. This sequence is thus an optimal solution for the MILP.

$$S_{MILP} = \left\{ \begin{pmatrix} 7 \\ 2 \\ 3 \\ 0 \end{pmatrix}, \begin{pmatrix} 9 \\ 10 \\ 8 \\ 3 \end{pmatrix} \right\}$$

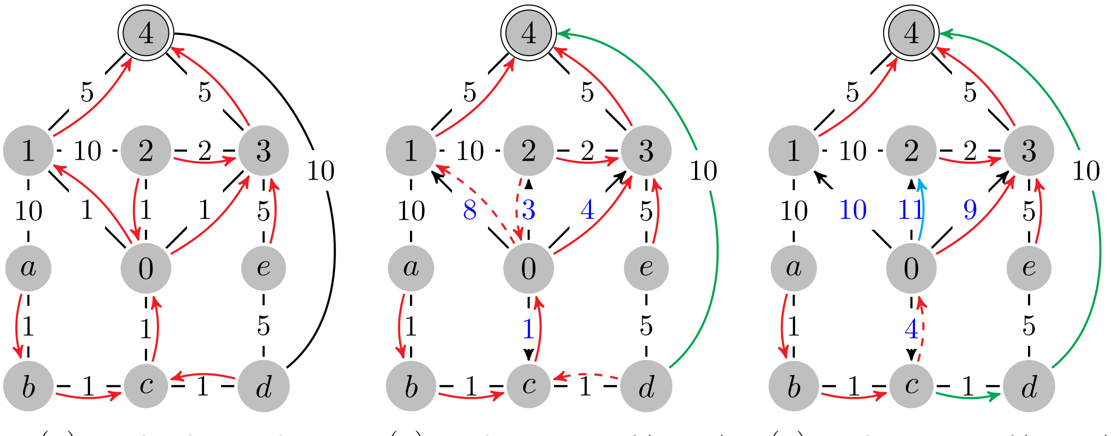
Compared to the solution produced by GBA, the first vector of this sequence differs only on the second component, which is one weight unit larger. Albeit slight, this difference prevents node 0 from using 2 as a next-hop towards destination 4, hence the loop that could have occurred between the two nodes (Fig. 2.13b). Similarly, the second vector is the same as the original GBA sequence, but for an increment of 3 on the last component that avoids the use of edge $(0, c)$ (Fig. 2.13c).

Generally speaking, in order to target minimality from a GBA-based perspective, two strategies can be adopted to prevent intermediate loops, namely, 1) modify the current greedy vector to avoid the intermediate change at r; OR 2) add a dynamic constraint to the computation of the next greedy vector, to force another predecessor node participating in the loop to not use r before it switches. Unfortunately, none of the two strategies always leads to minimal sequences when applied independently. While the presence of alternative strategies at each step seems to force a combinatorial space exploration, the theoretical problem of efficiently solving MILP is left open. However, our evaluations show that a heuristic based only on the second strategy, i.e., DGBH, computes sequences as short as GBA in the vast majority of our experiments performed on real-world IGP networks.

3.3 Technical workaround for intermediate transient loops

It is worth noticing that, even though transient loops might affect a large amount of traffic flows transiting in the network, such inconsistencies are always due to route changes performed by the node whose weights are being modified. Besides, local modifications are already required on this node in order to apply the update sequence. We thus consider a technical solution limited to this node to be a reasonable possibility. Our technical workaround relies on the local convergence delay mechanism [LDFF14], which is currently under standardization process at the IETF but already available in latest Cisco IOS (XR) releases. This mechanism introduces a delay between the convergence of the local router and the rest of the network. The delay is positive in case of a weight increment, so that Link-State Packets (LSPs) are flooded normally while the local shortest paths computation is slightly delayed. As a result, the local router will converge after its neighbors, preventing it from being implied in any transient forwarding loops. On the other hand, a negative delay is used for weight decrement operations, causing LSPs to be flooded after a short time, while shortest paths are computed normally on the local router. When enabled on a router, this mechanism is triggered on every local weight reconfiguration.

In our situation, *local delay* would be triggered for every intermediate update, effectively preventing local transient loops with no impact on the update sequence computed by GBA. However, the router undergoing the modification would still converge according



(A) Initial paths towards 4 (B) Weight increment of (7 2 3 0) (C) Weight increment of (9 10 8 3)

FIGURE 2.13: Illustration of a sequence optimally solving the MILP for destination 4.

to the last intermediate update before the next one is processed. As a result, intermediate route diversions, together with their other negative effects, are not prevented. Intermediate forwarding changes could theoretically be avoided by extending the desynchronization period between the Routing Information Base (RIB) and FIB to cover the application of the whole sequence. Such solution cannot be used in practice, though, because preventing the forwarding plane from reacting to concurrent topological modifications could be much more harmful for the network.

Table 2.2 shows a comparison of the sequence lengths achieved by the different variation of GBA on our running example, with their respective guarantees in terms of intermediate disruption avoidance. Note that *minimal link-by-link* and *minimal intermediate loop free* sequences are only provided for informative purposes, since we do not have, at the present time, an efficient algorithms for either of them. All of the methods in this table are safe with respect to both static and intermediate transient loops, but only uniform and AGBA sequences ensure the absence of intermediate forwarding changes. Since repeated modifications of the packet forwarding paths may have a negative impact on the control mechanisms implemented at the transport layer, increment sequences meeting the CPCs should be preferred whenever possible. As such, AGBA is the best possible choice, for it provides provably minimal sequences for problem 3.1. However, should these sequences appear too long, DGBH could be used instead as a decent compromise between routing stability and sequence length. Although this heuristic often yields sequences of the same length as GBA in practice, it may occur that longer sequences are produced in some rare cases. A feature such as *local-delay* could then serve as a technical workaround to prevent intermediate transient loops. If temporary routing instability is not a problem, and this feature is available, combining *local-delay* with sequences computed by GBA appears the most efficient solution to prevent any potential transient loops that could arise during network convergence.

	Sequence length for the removal of 0	Intermediate disruptions avoidance	
		Transient loops	Forwarding instabilities
Minimal link-by-link	5	✓	✗
Minimal uniform	5	✓	✓
AGBA	4	✓	✓
DGBH	4	✓	✗
Minimal int. loop free	2	✓	✗
GBA with local delay	2	✓	✗

Table 2.2: Sequence lengths for the removal of 0 on Fig. 2.10 and intermediate disruption avoidance levels

4 Towards an efficient implementation

In this section, we describe an efficient implementation of our main algorithm, GBA. Although the actual code that we are using for our experiments is written in *C* for performance purposes[1], we show here an higher level representation based on *Python*-like syntax, which is easier to read and explain.

Our implementation mainly relies on one core structure, which is specific to each destination and contains the merging of the initial and the final RSPDAG. The primary goal of this structure, that we often denote as *mdag* for merged DAG, is to exhibit potential transient loops that could occur in case of an abrupt modification. Besides, it also enables to manipulate the forwarding paths resulting from any intermediate weight increment performed on the modified router. The properties of *delta values* indeed makes it unnecessary to re-calculate any intermediate RSPDAG, for they are only a combination of *PRE* and *POST* edges that belong to either one of the two extremity RSPDAGs. For a given intermediate weight increment, a node simply uses its initial next-hops if the *distance increment* is lower than or equal to the delta value of this node, and its final ones if the increment is greater than or equal to the delta value. This reduces the complexity of constructing an intermediate RSPDAGs from $O(|N| \times log(|N|) + |E|)$ to $O(|N|)$ Note that we use scalar delta values for every such destination-oriented operation, and rely on *offsets vectors* (Def. 3.2) for global calculations. Aside from improving the readability of the algorithms, this minor enhancement decreases both the memory consumption and the computing time. All this information is computed for each destination during the initialization stage (Function 1).

4.1 Constraint extraction and removal

Unlike the solutions proposed in Section 2.1, our implementation does not rely on a preliminary extraction of all loop-constraints. Instead, relevant constraint values are calculated *on-the-fly* at each iteration of GBA. That is, for each destination, only the largest lower bound among the unsatisfied constraints is captured. Due to the total order existing among *delta vectors*, and thus constraints, for a given destination, it is indeed sufficient for the greedy vector calculation to consider only one lower bound per destination. In addition, the removal of satisfied constraints is made implicit by the evolving behavior of our *mdag* structure. This graph is designed to constantly reflect the potential forwarding inconsistencies that could occur during a convergence from the initial routing state to the one induced by the last computed vector. As the sequence calculation progresses, the graph is thus pruned of all nodes that may no longer be

[1] http://sourceforge.net/projects/metric-incr/

```
def mdag_init (G, r, d):
  # Compute PRE and POST DAGs
  DAG_PRE, SR = G.shortest_paths (d, r)
  G_POST = G.copy().remove_node (r)
  DAG_POST = G_POST.shortest_paths (d, SR)

  # Compute Delta values
  mdag.Delta = [ DAG_POST.shortest_path_len (s,d) - \
      DAG_PRE.shortest_path_len (s,d) for s in SR ]

  # Compute offset values
  mdag.offset = [ G.weight (r,s) + DAG_POST.shortest_path_len (s,d) - \
      DAG_PRE.shortest_path_len (r,d) for s in G.successors (r) ]

  # Merge PRE and POST DAGs
  mdag = compose (DAG_PRE, DAG_POST, SR)

  # Initialize 'roots' list
  mdag.roots = DAG_PRE.predecessors (r)

  # Initialize 'swallow_list'
  swallow_list = [ u for u in mdag.nodes () \
      if not mdag.predecessors (u) or not mdag.successors (u) ]

  # Compute the first increment for this destination
  mdag.popMax (swallow_list)

  # Ignore the destination if there is no possible loop
  if next_increment == 0:
    mdag = None

  # Return the newly created DAG
  return mdag
```

Function 1: Initialization

involved in a transient loop. We denote this process as *graph swallowing*, for it consists in progressively ignoring safe parts of the graph, that may no longer contain any transient loop, in order to focus on the portion containing unresolved inconsistencies. If, at any iteration, the graph is completely *swallowed*, i.e. all remaining nodes are removed, this indicates that the current sequence is loop-free for this destination. Obviously, if this happens before any vector has been computed, then no transient loop could occur for the destination. The *graph swallowing* thus serves both as an efficient cycle detection mechanism, with a time complexity in $O(|E|)$, a constraint extraction technique, and an implicit method to remove satisfied constraints. Indeed, cycles corresponding to satisfied constraints simply do not appear at the next iteration.

Graph swallowing for cycle detection

The first, most intuitive usage for the graph swallowing technique is to detect the presence of cycles in a directed graph. As such, the algorithm consists in progressively removing from the graph the nodes that cannot possibly be involved in a cycle. It starts from the list of nodes that have no predecessors in the graph, a.k.a. *root nodes*, and thus may be involved in no cycle. Each node in this *swallow list* is successively removed from the graph, or *swallowed*, along with all its outgoing edges. That may cause some of its successors to become *roots* themselves, for their last predecessor has just been removed. Since these nodes hence cannot be involved in a cycle either, they are added in the *swallow list*. Recursively, every such node is thus removed from the graph. The algorithm ends when the *swallow list* is empty. At that time, if all nodes have been swallowed, and the graph is empty as well, then it contained no cycle. Otherwise, there is at least one cycle that is located in the remaining part of the graph. Indeed, it means that every remaining node has at least one predecessor, which is only possible in case of a cycle.

This algorithm has a time complexity in $O(|E|)$ plus the cost of retrieving the initial *swallow list*. In the case of a merged RSPDAG, the list can be calculated in $O(1)$, for it contains only the destination node.

Note that this process can also be performed in the other direction. That is, starting from *leaves* and progressively swallowing nodes that have no more successors. It is even possible, with no extra cost, to swallow from both directions at the same time, as in Function 2, to better pinpoint the area of the graph where inconsistencies linger.

Graph swallowing on intermediate graphs

Another usage of the graph swallowing technique is to monitor the effects of intermediate vectors on a merged RSPDAG. More precisely, to perform the cycle detection considering a transition from the initial state to the one induced by a given intermediate vector, without having to actually compute the associated intermediate graph. Once again, this technique relies on the properties of delta values, which imply that, if the distance increment produced by the intermediate vector towards the destination is lower than the delta value of a node, then this node is still in its initial routing state. It thus cannot possibly be involved in a transient loop when considering a convergence from the initial state.

Starting from the result of a previous swallowing operation, which has been stopped early because of a cycle, the algorithm simply consists in resuming the swallowing process after having removed every node whose delta value is larger than the distance increment of the vector. This causes any cycle involving one of these nodes to be disregarded. Hence, only transient loops that could occur during a convergence from the initial routing state to the one produced by the vector are considered.

Graph swallowing for extracting constraints

The principle of GBA states that, at each iteration, the greedy vector to be computed must be positively greater than the lower bound of all remaining constraints. It is thus necessary to retrieve, for every destination, the largest lower bound among the remaining loop-constraints. That is, the smallest delta such that the convergence from any larger vector to the last computed vector is loop-free.

This smallest delta can be extracted from a reduced version of a merged RSPDAG considering the convergence between the initial and current state, which can be computed as described in the previous section. The process consists in repetitively swallowing the graph after having removed the nodes with the smallest delta, until the graph can be completely emptied. The largest lower bound is equal to the delta vector of the nodes removed at the last step. In practice, this delta is extracted as a scalar and denoted *next increment*.

```
def swallow (mdag, swallow_list):
  # While there is nodes to swallow
  while swallow_list:
    # Pop the first node in the list
    LF = swallow_list.pop (0)

    # Remove its outgoing links
    for s in mdag.successors (LF):
      mdag.predecessors (s).remove (LF)
      if not mdag.predecessors (s):
        # Add new leaves at the end of the list
        swallow_list.append (s)

    # Remove its incoming links
    for p in mdag.predecessors (LF):
      mdag.successors (p).remove (LF)
      if not mdag.successors (p):
        # Add new roots at the end of the list
        swallow_list.append (p)

    # Update max Delta
    mdag.max_delta = max (mdag.max_delta, mdag.Delta[LF])

    # Remove the current node from 'roots'
    mdag.roots.remove (LF)

    # Add its predecessors in the PRE graph to the 'roots' list
    mdag.roots.append (mdag.pred_old (LF))

    # Remove the current node from the DAG
    mdag.remove (LF)
```

Function 2: Graph swallowing

4.2 Algorithmic improvements

Affected destinations retrieval

When a topological modification on a single link or router is performed, only part of the flows passing through the network are impacted. The proportion of impacted flows depends on the centrality of the modified component. For many operations, it is thus possible that no flow towards a subset of destinations are impacted. No transient loop could occur for these destinations and they do not have to be considered in the process of computing a weight increment sequence. Using GBA, such destinations are quickly set aside as no cycle is detected in their merged RSPDAG. However, this requires computing two RSPDAGs and running a cycle detection algorithm, which represents a time complexity in $O(|E| + |N|\ log(|N|))$, for each destination. It would be interesting to decrease this cost by reducing the set of destinations given as input to GBA to those whose routes are actually affected.

On the one hand, in the case of a modification on a single link, these *affected destinations* can be easily identified. They are those the modified link source node reaches through this link, considering the initial routes in the case of a weight increment (or link removal) operation and the final routes in the case of a weight decrement (or link addition). More formally, let us denote as w_i and w_t respectively the initial and target weights of the modified link $l = (a, b)$. The set of affected destination D_l is equal to $N \cap G_l$, where G_l represents the set of nodes downstream of b in $RSPDAG(a)$ if $w_i < w_t$ and $RSPDAG'(a)$ if $w_i > w_t$.

```python
def affected_destinations_link (
    G,     # Initial network graph
    H,     # Final network graph
    lsrc,  # Link source node
    ldst   # Link destination node
    ):

  if G.edge[lsrc][ldst]['weight'] < H.edge[lsrc][ldst]['weight']:
    dag = G.shortest_paths(lsrc)
  else:
    dag = H.shortest_paths(lsrc)

  return dag.subgraph(ldst).nodes()
```

Function 3: Computing the set of affected destination for single link modifications

On the other hand, retrieving the set of destination affected by a node-wide modification is more complex. The only sure thing is that the source node (i.e. the one to be shut down or the source of all modified links) cannot be part of the set of *affected destinations*. Indeed, the weight of the links towards this node remain unchanged, and so does the routing graph. As for the remaining nodes, it depends on whether or not the modified node is to be considered as a source. If it is, e.g. in the case of arbitrary weight modifications on its outgoing links, then the set of affected destinations can be computed using a slight variation of the algorithm used for single link operations. We consider as affected each destination reached from the source node through any of the modified links. As a consequence, if all outgoing links of the node are modified, each other node in the network is an affected destination. However, in the case of a shutdown or startup operation, it can be reasonably assumed that the modified node is not a source while the weight update sequence is being applied. Only transiting traffic, coming from other nodes in the network and passing through the modified node is to be taken into account. Thus, affected destinations are those reached through the modified node from any source other than this node. Though such definition would require to compute a Shortest Path DAG (SPDAG) for every node, but the modified one, in the network, it can be reduced to the SPDAG of the predecessors of the updated node. Indeed, if the path from a given source to a destination goes through the modified node, then there is at least one predecessor of the updated node such that the path from this predecessor to the destination goes through the node. Formally, the set of affected destinations is the union of the nodes downstream of the modified node in the routing graphs of each predecessor of this node. Eventually, in order to handle the general case of an arbitrary subset of the outgoing links of a node being modified, we use a combination of the previous mechanisms as described in Function 4.

```
def affected_destinations (
  G,  # Network graph
  r,  # Source node
  ll  # List of outgoing links of ‘r’ to be modified
  ):

 D = set ()

 for pred in G.predecessors (r):
  dag = G.shortest_paths (pred)

  for link in ll:
   if link.head in dag.successors (r):
    D |= dag.subgraph(link.head).nodes()

 return D
```

Function 4: Computing the set of affected destination in the general case

Subgraph reductions

Aside from the set of destinations, it is also possible to improve several aspects of graph-related calculations. In particular, by considering only a subset of nodes at some points in the algorithm. The first improvement consists in reducing the set of source nodes to be considered, for each destination, to those initially reaching the destination through the modified component. Indeed, only these nodes may be involved in a transient loop. In practice, *affected sources* are marked as such during the initial RSPDAG calculation (Fn 1, line 3). The final and merged RSPDAGs are then computed considering only this subset of nodes (Function 1, lines 5 and 17).

The second improvement aims at enhancing the performances of the satisfied constraints removal. Rather than iterating over every remaining node in the *mdag* to check whether its delta value is larger than the distance increment of the current greedy vector, it is possible to consider only a subset of these nodes. More precisely, the predecessors in the initial RSPDAG of already removed nodes, for they hold the largest delta values among the remaining nodes. This set of potential *roots* is filled in the swallowing function (Function 2, line 28) and used to construct the list of nodes to removed in Function 7.

The third improvement consists in skipping the *next increment* extraction phase if the greedy vector computed at the current iteration of GBA has no impact at all on the *mdag* for a given destination. This situation may occur when the distance increment produced by this vector is larger than any delta value among the remaining nodes in the *mdag*. In that case, no constraint may have been satisfied for this destination, and the largest lower bound of a constraint is the same as the previous iteration. The largest delta value is extracted at the same time as the *next increment*, within the swallowing function (Function 2, line 22).

4.3 Sequence calculation

Let us now specifically describe each function involved in our efficient GBA implementation.

The core GBA function (Function 5) starts by computing the set of *affected destinations* as the nodes that are reached through 0 by at least one source (other than 0 itself). Indeed, if node 0 is not used by any source to reach a given destination, no transient loop could appear for that destination. Then, for each affected destination d, our algorithm computes $mdag(d)$, the merging of the initial and final forwarding graphs towards d. At this stage, the *popMax* function checks whether transient loops could appear and, if so, computes the *next increment*. If the returned increment is greater than 0, i.e., if there is at least one constraint, an offset value is then computed for each outgoing link of node 0. Otherwise, it means that no transient loop could possibly appear for this destination. Eventually, the $mdag(d)$ is added to the global *MDags* set.

Once the *MDags* set is computed, our algorithm enters the second phase. At each round of the global loop, a new greedy vector v is computed (and added to the sequence S) as the smallest one that is safe with respect to the *next increment* for all subgraphs in the *MDags* set. Then, for each destination d, the actual distance update m associated to this vector is computed, and function *upMax* is called in order to extract the next increment from each impacted *mdag*. This function modifies the graph, now considering v as the final weight assignment, and prunes all nodes that cannot be involved in any cycle. It then extracts the *next increment*, if any, and returns 0 otherwise. If there are no more constraints to be satisfied for this destination, it is removed from the *MDags* set. The main loop iterates this way until *MDags* is empty, meaning that all constraints are satisfied by the sequence S.

Function *swallow* (Function 2) iterates over nodes in the prepared *swallowing list*, removing from the *mdag* each node in the list and, recursively, all their neighbors that have either no successors or no predecessors when removing them from the current *mdag*. In addition, the function maintains a variable with the largest delta value among the removed nodes (l. 22) and a list of *roots* containing, for each removed node, its predecessors in the initial routing graph that are still in *mdag* (l. 28).

Function *popMax* (Function 6) starts by permanently pruning from the *mdag* all nodes that are not, or no longer, involved in any transient loop (l. 7). Then, this function *clones* this current state of the *mdag* (l. 10) and repeatedly performs *swallowing* operations on a temporary copy in order to extract the next increment. At each iteration, the *minimum delta value* among the remaining nodes in the copy is extracted (l. 15) and the associated nodes are added to the next *swallowing list* (l. 18). It iterates this way until no cycles appear. Eventually, the minimum delta value computed at the last step is stored as the *next increment* to be considered by GBA (l. 25).

Function *upMax* (Function 7) initializes the next *swallowing list* with every node in *roots* whose associated delta value is greater than m (l. 6–7). Besides, the predecessors in the initial routing graph of each node to be *swallowed* are added to the *root list* (l. 8). *upMax* iterates this way until all its elements have been treated.

```
def GBA (
    G,   # Network graph
    r,   # Source node
    ll,  # List of outgoing links of ‘r’ to be modified
    wl   # List of weights to be applied
    ):

  Seq = []       # Vector sequence
  MDags = []  # Set of Pre-Post graphs

  # Initialization
  for d in D:
    mdag = mdag_init (G, ll, wl, d)
    if mdag.next_increment > 0:
      MDags.append (mdag)

  # Main GBA loop
  while MDags is not []:
    # Add the new greedy vector to the sequence
    Seq.append ([
      max ([
        mdag.next_increment - mdag.offset[succ]
        for mdag in MDags
        ])
      for succ in G.successors (r)
      ])

    for mdag in list (MDags):
      # Compute the distance increment
      m = min ([
        S[-1][s] + dag.offset[s]
        for s in G.successors (r)
        ])

      # Check whether this increment would have an impact on the nodes in ’mdag’
      if m < mdag.max_delta:
        # Load new constraints
        mdag.upMax (m)

        # Remove completed destinations
        if mdag.next_increment == 0:
          del (mdag)

  return Seq
```

Function 5: Sequence calculation

The worst case time complexity of this implementation of GBA is determined as follows:

- The complexity of the initialization is held by the calculation of the merged RSPDAG and the extraction of the first increment. Considering all destinations, this stage would have a cost of $O(|N| \times (|N|log_2(|N|) + |E|)$ with a priority queue system based on a *Fibonacci heap*. However, we rely in practice on a *binary heap* having a larger worst case complexity, in $O(|E|log_2(|N|))$, but allowing for shorter computing times on realistic ISP topologies;

```
def popMax (mdag, swallow_list):
  # Reset ’next_increment’ and ’max_delta’ variables
  mdag.next_increment = 0
  mdag.max_delta = 0

  # Swallow the DAG considering the current vector
  mdag.swallow (swallow_list)

  # Create a copy of the DAG
  GF = mdag.copy ()

  # Completely swallow this copy in order to extract the next increment
  while GF:
    # Extract the min Delta value among the nodes in GF
    GF.next_increment = min ([ GF.Delta[i] for i in GF.nodes ]) + 1

    # Retrieve the nodes associated to the min value
    min_nodes = [ i for i in GF.nodes if GF.Delta[i] == GF.next_increment ]

    # Swallow the graph from these nodes
    GF.swallow (min_nodes)

  # Retrieve ’max_delta’ and ’next_increment’ from the copy
  mdag.max_delta = GF.max_delta
  mdag.next_increment = GF.next_increment
```

Function 6: Satisfied constraints removal

```
def upMax (mdag, m):
  new_roots = []
  swallow_list = []
  while mdag.roots:
    current_node = mdag.roots.pop (0)
    if mdag.Delta[current_node] > m:
      swallow_list.append (current_node)
      mdag.roots.append (mdag.pred_old (current_node))
    else:
      new_roots.append (current_node)
  mdag.roots = new_roots
  mdag.popMax (swallow_list)
```

Function 7: Next increment extraction

- The number of iteration of the main loop corresponds to the length of the resulting sequence, and may thus be equal to $|N|^2$ at worst. Yet it can be limited in practice to a reasonable length, denoted p. This main loop hence comes at the cost of:
  - $O(p \times k \times |N|)$, where k is the degree of node r, for the calculation of the greedy vector;
  - and $O(min(p \times |N| \times |E|, |N|^2 \times |E|))$ for the extraction of the next increment.

Eventually, GBA has a worst case complexity in $O(|N|^4)$ if node r has a degree of $k = |N|$ (or if $|E| \approx |N|^2$ in general). However, in practice p can be picked as an arbitrary low value, such as $p \leq 5$, to limit the complexity of GBA to $O(|N|^3)$.

5 Conclusion

In this chapter, we first presented the basics of our approach, explaining how it can be used to prevent transient forwarding loops in the case of a router-wide modification. We also show that aiming for short update sequences through heterogeneous weight modifications may lead to a different kind of transient disruption. We then detailed our Greedy Backward Algorithm (GBA) for computing weight increment sequences that prevent normal transient loops. In particular, we proved that GBA yields sequences of minimal length for this problem. As for intermediate disruptions caused by heterogeneous weight increments, we proposed several solutions associated with different level of routing stability. The Adjusted Greedy Backward Algorithm (AGBA) is a variation of GBA producing sequences that prevent changes in the set of next-hops used by the modified router, thus ensures the absence of any intermediate disruption. We then described a relaxed solution, called Dynamic Greedy Backward Heuristic DGBH, that focuses on preventing transient loops caused by intermediate forwarding changes. This solution enables to compute shorter sequences at the expense of lower guarantees in terms of routing stability. We also show that it is possible to achieve the same effect without modifying the sequences produces by GBA, by temporarily disabling the synchronization between the control and data-plane of the modified router. Finally, we presented an efficient implementation of GBA, discussing several improvements we made in order to lower the time complexity and keep the computing time within reasonable limits, as we will show in the next chapter.

Chapter 3

Evaluations

Contents

In this chapter, we provide thorough evaluation of our algorithms on real and inferred network topologies, in order to assess the practicality of our approach in a realistic environment. We first give some insight about the graphs we are working on, presenting their intrinsic properties and showing how much each of them could be affected by transient forwarding loops. In a second part, we analyze the length of the sequences produced by Greedy Backward Algorithm (GBA)-based algorithms, and compare them with single-link based solutions. We show that, while GBA and its variations produce short enough sequences in most cases, uniform and link-by-link solutions are not suited for router-wide operations. Finally, we evaluate the computing time performances of our implementation. In particular, we show that the time required to compute a sequence remains reasonable even on very large topologies, and detail the effects of each of our algorithmic optimization. We also discuss some schemes for a practical deployment of our solutions.

1 Evaluation setup

1.1 Graph characteristics

Table 3.1 presents the main properties of our real Internet Service Provider (ISP) topologies. Internet2 [Int] and GEANT [gea] networks are two well-known networks whose weighted graphs are freely available. RENATER is the Internet provider for education and research institutions in France. Although we are not authorized to disclose their weighted topology, which we were provided as part of a research collaboration, a plain topology (without Interior Gateway Protocol (IGP) weights) is available online [REN]. The last six networks are real ISPs that we anonymized for confidentiality reasons. Numbers of links and nodes are also rounded for the same reason.
Note that our evaluation topologies have symmetric weights, i.e. the weight configured for the adjacency $A \rightarrow B$ is the same as the one for $B \rightarrow A$. This is not a requirement of our solutions, but seems a common practice in ISP networks.

We evaluated our algorithms on a wide set of real and inferred IP network topologies of various shape and size. Networks in table 3.2 are Rocketfuel inferred topologies obtained with *traceroute* campaigns [SMW02, MSWA02]. Using the popular route discovery tool, the authors gathered a large set of network traces from multiple vantage points all around the world. Then, they isolated each Autonomous System (AS) based on the routers IP addresses and domain names and used alias resolution mechanisms to retrieve the routing-layer topology. Finally, they assigned weights to the links according to a system of constraints, which was obtained by assuming that network traces represent

Network	$\|N\|$	$\|E\|$	Diameter	Max. degree	Weight space
Internet2	9	26	4	4	$[277, 1705]$ (13)
GEANT	22	72	4	6	$[1, 20050]$ (18)
RENATER	70	230	11	13	$[1, 1000]$ (14)
ISP 1	25	55	6	6	$[1, 11]$ (4)
ISP 2	55	200	5	20	$[10, 50000]$ (8)
ISP 3	110	350	11	8	$[1, 9999]$ (32)
ISP 4	150	400	13	9	$[1, 9999]$ (32)
ISP 5	200	800	13	14	$[1, 66666]$ (55)
ISP 6	1200	4000	12	56	$[1, 100010]$ (105)

Table 3.1: Real ISP graph properties

shortest path with respect to IGP weights. Such a technique allows for a reasonable approximation of the routing topology of a given AS. However, it is not possible to ensure that the whole network has been discovered, nor that the inferred weights are correct. For instance, equal-cost and backup paths that are no used by *traceroute* probes may not be detected during the measurement campaign. Also, the linear program obtained from the shortest path constraints does not have a unique solution. In particular, multiplying all inferred weights by the same factor leads to an equally valid solution.

Network	$\|N\|$	$\|E\|$	Diameter	Max. degree	Weight space
Exodus	79	294	10	12	$[1, 22]$ (17)
Ebone	87	322	11	11	$[1, 16]$ (14)
Telstra	108	306	8	18	$[1, 7]$ (6)
AboveNet	141	748	8	20	$[1, 20]$ (14)
Tiscali	161	656	10	29	$[1, 22]$ (20)
Sprint	315	1944	10	45	$[1, 16]$ (15)

Table 3.2: Inferred graph properties

Similar works have been recently conducted in [MDP+11, MMD+11] to provide the networking community with more accurate and up-to-date ISP topologies. *MERLIN* combines *mrinfo* [PMDB10] and *paris-traceroute* [VAC+08] probing in order to improve the network coverage. The former relies on Internet Group Management Protocol (IGMP) to collect the list of adjacencies of multicast-capable routers, including backup and unused links that would not have been discovered otherwise, while the latter is an improved version of *traceroute* designed to detect load-balancing and retrieve alternate routes.

Due to the large variations in terms of size and shape within our set of evaluation topologies, representing on the same figure the results obtain for all topologies may lead to loss of readability. We thus often split our evaluation sample as follows.

- Small ISPs (less than 100 nodes): Internet2, GEANT, RENATER, ISP1 and ISP2;

- Large and very large ISPs (100 to 1200 nodes): ISP 3 to 6;
- Rocketfuel inferred graphs: Exodus, EBone, Telstra, AboveNet, Tiscali and Sprint.

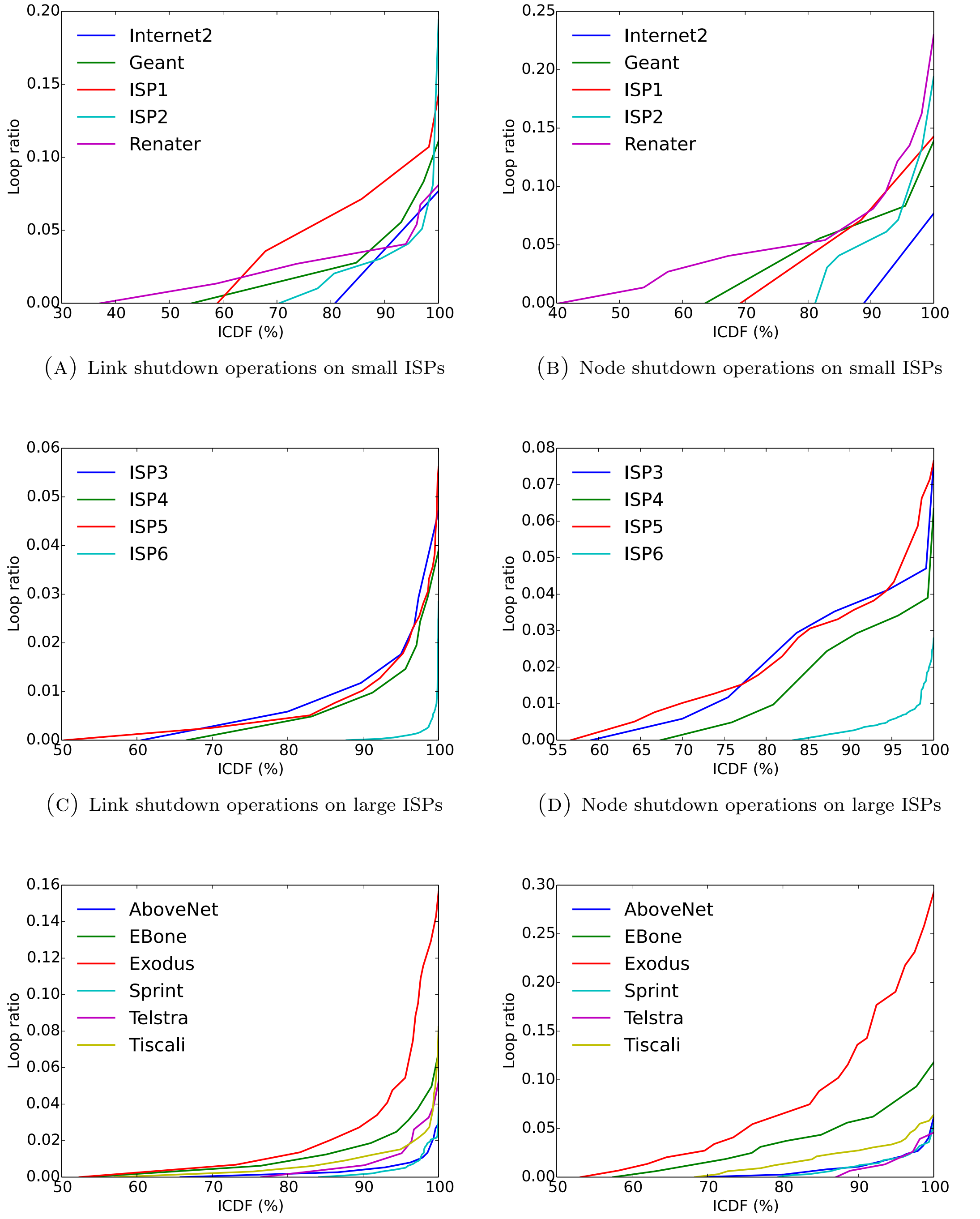
We perform our evaluations considering the removal operations of a single link or router from our topologies. As we mentioned before, our techniques also apply to link/router addition as well as any addition, removal, positive weight increment or decrement of any subset of the outgoing links of a single router. However, studying such extended use cases would require arbitrary decisions that may not represent realistic scenarii. In particular, we would have to choose which routers are to be connected by a new link and assign a weight to this link, bringing the network in a state it is no likely to ever take in practice. We thus limit our evaluations to operations that we know may happen.

1.2 Transient loop evaluations

On Fig. 3.1, we show the distribution of transient forwarding loops that could potentially occur for each link (left-hand side figures) and node (right-hand side figures) shutdown operations. For a given operation, we enumerate the elementary cycles on the graph resulting from the merging of the initial and final Reverse Shortest Path DAGs (RSPDAGs) for every destination, and compute the ratio of the number of cycles over the total number of links in the initial topology. Since our evaluation topologies are symmetrically weighted, all elementary cycles in the merged graphs are of size 2, so that the *loop ratio* presented in the figures represents the proportion of links on which the traffic may loop during the network convergence.

These figures are interesting for two reasons. On the one hand, they indicate the percentage of operations that are *safe* with respect to transient loops. *Safe* operations may not lead to transient forwarding loops during the convergence of the network, whatever the routers update order. On all but one topology, more than half of single link and node shutdown operations could be safely performed without requiring any intermediate update. Such safe components are, for one part, backup links or leaf nodes that are not used for transit, and can be removed from the network with few impact, if any, on the routing decisions. For the other part, they are surrounded by routers having a local backup solution. That is, their alternate shortest path to the destination does not involve sending the traffic backward.

The proportion of safe operations tends to be even larger for router removals, with five topologies (Internet2, ISP2, ISP6, Sprint and Telstra) above 80%. The main reason behind this phenomenon is that the outgoing links of the removed router are considered down in the final state. Hence no transient loop could occur in the direct vicinity of



(A) Link shutdown operations on small ISPs

(B) Node shutdown operations on small ISPs

(C) Link shutdown operations on large ISPs

(D) Node shutdown operations on large ISPs

(E) Link shutdown operations on Rocketfuel graphs

(F) Node shutdown operations on Rocketfuel graphs

FIGURE 3.1: Impact of shutdown operations on real and inferred ISP networks

the router. In practice, these links should not be taken down directly, for this could cause packet losses due to temporarily unreachable destinations, but the behavior can be mimicked using special routing configurations such as *overload bit*. The router positioning an overload bit continues to forward its traffic according to the initial topology, while the rest of the network should recompute new routing paths avoiding this router. Some ISPs topologies may also be more or less prone to transient loops. Overall, ISP6 (Fig .3.1c and 3.1d) is the safest topology, with 88% of link and 83% of node removals, while RENATER (Fig .3.1a and 3.1b) has the least proportion of safe operations, with 37% and 40%, respectively for links and nodes removals. Such a disparity can be explained by specific graph patterns. For example, topologies displaying high local redundancy are more likely to have local backup paths available, and thus less potential transient loops. On the contrary, ring patterns can require to go a long way backward before leaving the impacted subgraph, leading to many potential loops.

On the other hand, these figures also indicate how much the routing graph could be impacted by a shutdown operation. Although we only show potentialities of loops that may never arise in practice, the highest proportion of links can undergo transient loops, the more traffic is likely to be delayed or lost. The figures show that, on most networks, transient loops usually affect a very small part of the topology, representing between 1 and 2% of the links, and remains below 10% even for worst cases. They may however have a much greater impact on other networks, such as ISP1, ISP2, RENATER and Exodus, where transient loops can occur on more than 15% and up to 30% of the links. Interestingly enough, these are all small networks, with less than a hundred routers. While it is unlikely that the size alone has such an impact on how much the network is affected by transient inconsistencies, it is possible that small ISPs focus on coverage at the expense of less redundancy in their networks. This could result in more ring patterns where transient loops could occur. It is also worth mentioning that router removals, which represent more significant modifications to the network compared to single link shutdowns, result in more potential perturbations.

2 Sequence lengths

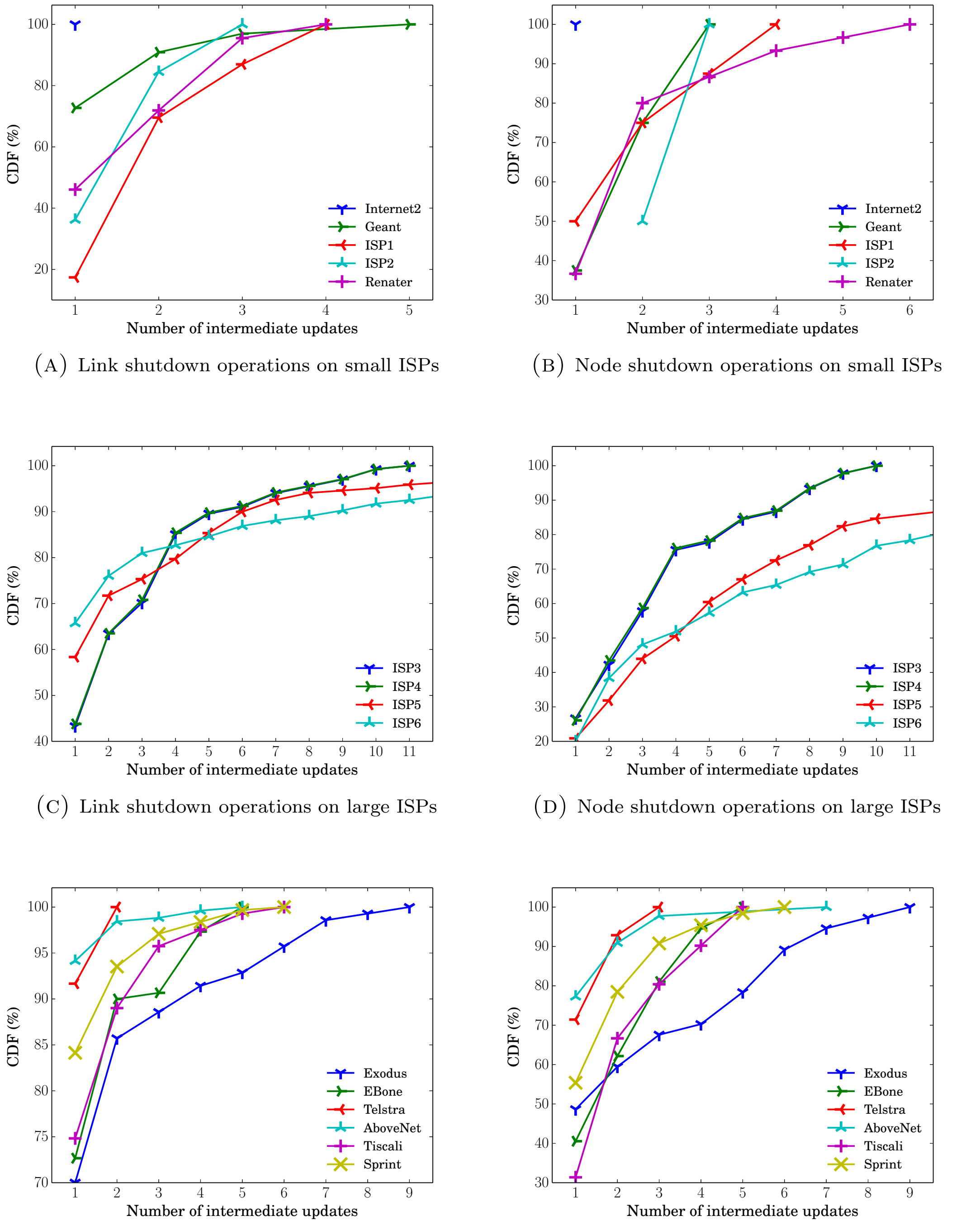
In chapter 2 we proved that our main algorithms, GBA and Adjusted Greedy Backward Algorithm (AGBA), compute sequences of minimal length in their respective categories. Such theoretical results take however no account of the practical limitations related to our approach. In particular, the time required for a sequence to be applied in the network, which directly depends on its length, may quickly become prohibitive in practice. While it seems realistic to delay a maintenance operation by a few seconds in order to apply a short sequence, processing a hundred updates long sequence may seriously hinder network management and increases the risk of concurrent modifications.

In this section, we evaluate the length of intermediate update sequences produced by our algorithms on real and inferred network configurations, showing how short they are in most cases. We first present the results yield by our standard GBA algorithm, assuming that intermediate inconsistencies are handled by an external mechanism. Then we compare these results with alternative sequence-based mechanisms, link-by-link and uniform, as well as GBA variations, AGBA and Dynamic Greedy Backward Heuristic (DGBH). Note that the results we present in this section, in particular the percentage of sequences of a given length, are always computed relatively to the number of *non-empty* sequences.

2.1 GBA sequences length

On Fig. 3.2, we show the cumulative length distribution of sequences computed for single link and node shutdown operations on each topology. As in the previous section, we split our set of evaluation topologies for the sake of clarity. For the same reason, we also limit the maximum length of sequences displayed on our figures to 11. This information can be found in Table 3.3, together with other relevant statistical data.

At first glance, these figures show that most sequences produced by GBA are short. On small ISP networks, more than half *unsafe* operations requires only 1 or 2 intermediate updates, while even the longest sequences contain no more than 5 elements for link-shutdown operations (Fig. 3.2a) and 6 for node-shutdowns (Fig. 3.2b).
GBA sequences are only slightly longer for Rocketfuel topologies (Fig. 3.2e and 3.2f). Worst case shutdown operations may require a few more updates, yet 50% of the routers, and even 85% of the links, can be safely shut down with a couple of intermediates updates. With up to 9 elements, longest sequences are produced for Exodus network, which is one of our smallest topologies. This tends to indicate that sequences length does not directly depends on the size of the network.



(A) Link shutdown operations on small ISPs

(B) Node shutdown operations on small ISPs

(C) Link shutdown operations on large ISPs

(D) Node shutdown operations on large ISPs

(E) Link shutdown operations on Rocketfuel graphs

(F) Node shutdown operations on Rocketfuel graphs

FIGURE 3.2: GBA sequences lengths for shutdown operations on real and inferred ISP networks

	$\lvert S\rvert \leq 5$	$\lvert S\rvert \leq 10$	max	$\lvert S\rvert \leq 5$	$\lvert S\rvert \leq 10$	max	Int. FC	Int. TL
Internet2	100 %	100 %	1	100 %	100 %	1	0 %	0 %
Geant	100 %	100 %	5	100 %	100 %	3	37.50 %	12.50 %
ISP1	100 %	100 %	4	100 %	100 %	4	50 %	0 %
ISP2	100 %	100 %	3	100 %	100 %	3	40 %	30 %
Renater	100 %	100 %	4	96.67 %	100 %	6	50 %	6.67 %
ISP3	89.55 %	99.25 %	11	77.78 %	100 %	10	82.22 %	24.44 %
ISP4	89.78 %	99.27 %	11	78.26 %	100 %	10	80.43 %	26.09 %
ISP5	85.35 %	95.12 %	39	60.44 %	84.62 %	33	80.22 %	18.68 %
ISP6	84.62 %	91.74 %	61	57.30 %	76.76 %	57	81.08 %	35.14 %
Exodus	92.86 %	100 %	9	78.38 %	100 %	9	91.89 %	8.11 %
EBone	100 %	100 %	5	100 %	100 %	5	62.16 %	2.70 %
Telstra	100 %	100 %	2	100 %	100 %	3	78.57 %	14.29 %
AboveNet	100 %	100 %	5	97.73 %	100 %	7	56.82 %	2.27 %
Tiscali	99.29 %	100 %	6	100 %	100 %	5	74.51 %	9.80 %
Sprint	99.68 %	100 %	6	98.46 %	100 %	6	81.54 %	15.38 %

(A) Link shutdown (B) Node shutdown

TABLE 3.3: GBA sequence lengths and possible disruptions

However, Fig. 3.2c and 3.2d show that shutdown operations on large networks may require very long sequences with up to 61 intermediate updates. Fortunately, such extreme cases only happen for a few operations and our investigations give us reason to believe that they are due to inconsistencies in the network at the moment the snapshot was taken. In most cases, shutdown operations can be safely performed after a reasonable number of intermediate updates. Five updates are sufficient for more than 55% of node and 80% of link shutdowns, while ten cover respectively 75 and 90% of the operations. One may also notice on these figures that the results produced for ISP3 and ISP4 are very close to each other. Indeed, both topologies actually represent two versions of the same network, which vary in terms of node and edges, but still exhibit the same network patterns.

Longer sequences may not be usable in practice, for they would require the actual operation to be overly delayed. Since our algorithms produce sequences of minimal length, it is not possible to shorten them without causing potential transient loops to no longer be covered. However, we could consider some transient loops more important than others and assign priority levels based on how impactful they might be for the traffic passing through the network. For example, transient loops located at the edge of the network, which are less likely to heavily disturb the traffic, could be assigned a lower priority than those involving more central routers. Such system could be included within the sequence computation process, making the constraints associated with lowest priority loops to be automatically ignored if the number of intermediate vectors grows too large. Similarly, it is also possible to exclude the least used prefixes from the list of destinations, making the associated transient loops not to be considered by GBA. Because all these decisions depend on the needs and concerns of each operator we did not however perform thorough evaluation of this approach.

The last two columns on Table 3.3 respectively present the proportion of node shutdown sequences that incur intermediate forwarding changes (FC) and transient loops (TL). These results show that at least one intermediate forwarding change occur for about 80% of node shutdown sequences on most networks. Besides, a significant part of non-empty sequences could lead to intermediate transient loops. This proportion tends to increase with the size of the topology, with more than 35% of affected sequences on ISP6, and emphasizes the need for intermediate disruption avoidance mechanisms.

2.2 Comparison with GBA alternatives

Since we proved that GBA yields sequences of minimal length, different algorithms, whether or not they are based on GBA, may only produce longer, or equally long, sequences. It could however be interesting to determine how longer these sequences are compared to GBA, in order to evaluate if it could be worth relaxing the property of minimality to benefit from other properties held by these alternatives, such as routing stability or computing efficiency. In this section, we present the length of sequences computed for node shutdown operations by alternative algorithms. Table 3.4 shows statistical information for all topologies, while thorough comparison on our largest real evaluation network is provided in Fig 3.3 and 3.4.

The *link-by-link* heuristic (Table 3.4a) consists in safely shutting down one after the other the outgoing links of the router to be removed, using either GBA or any other algorithm producing minimal update sequences for single link operations. Contrary to other router modification algorithm, this method does not benefit from the opportunity of simultaneously modifying the weight configured on multiple links. It is thus not affected by the intermediate transient loop problem. The traffic may be forwarded through intermediate next-hops when the weight on one link is increased, yet the loops that could arise from this situation will be considered as normal transient loops. On the other hand, intermediate update sequences are much longer than those produced by multi-link update based algorithms. Our evaluations report that the proportion of shutdown operations requiring 5 intermediate updates or less is 20 percentage points lower with this method compared to standard GBA and worst case sequence are more than 3 times as long.

The sequences used to produce these results were obtained considering the removal of the links in an arbitrary order. In most cases, this order has an impact on the final

	$\lvert S \rvert \leq 5$	max	$\lvert S \rvert \leq 5$	max	$\lvert S \rvert \leq 5$	max	$\lvert S \rvert \leq 5$	max
Internet2	100 %	3	100 %	1	100 %	1	100 %	1
Geant	72.73 %	9	100 %	5	100 %	3	100 %	3
ISP1	70.00 %	13	100 %	4	100 %	4	100 %	4
ISP2	45.45 %	22	70.00 %	7	100 %	3	100 %	3
Renater	72.97 %	19	90.00 %	8	96.67 %	6	96.67 %	6
ISP3	53.57 %	26	55.56 %	21	57.78 %	14	73.33 %	11
ISP4	55.93 %	26	56.52 %	21	58.70 %	14	73.91 %	11
ISP5	56.00 %	66	41.76 %	63	50.55 %	41	60.44 %	33
ISP6	49.62 %	174	50.27 %	147	54.05 %	67	57.30 %	57
Exodus	61.82 %	21	70.27 %	11	70.27 %	10	78.38 %	9
EBone	72.58 %	17	91.89 %	7	94.59 %	7	100 %	5
Telstra	83.87 %	8	100 %	4	100 %	4	100 %	3
AboveNet	80.23 %	20	97.73 %	7	97.73 %	7	97.73 %	7
Tiscali	71.88 %	20	94.12 %	6	98.04 %	6	100 %	5
Sprint	74.05 %	23	95.38 %	9	98.46 %	6	98.46 %	6

(a) Link-by-link (b) Uniform (c) AGBA (d) DGBH

Table 3.4: Lengths of node shutdown sequences produced by GBA alternatives

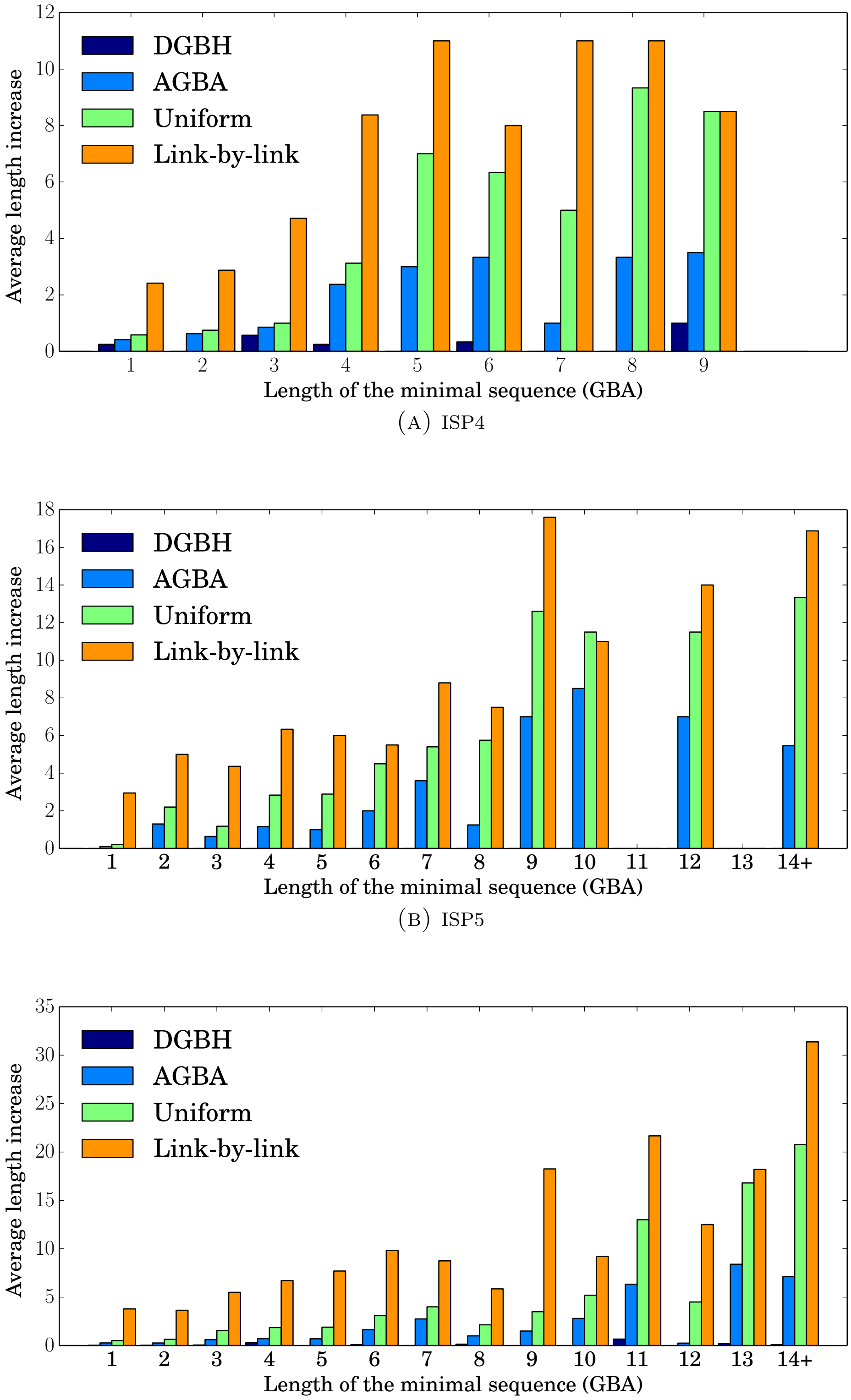
sequence length, and it may be possible to devise an algorithm finding the best possible order in all situation. Such algorithm could be based, for example, on the initial weights of the links to be modified, choosing to remove backup links first or, on the contrary, to keep them for last. Such sequences could not however be shorter than those produced by GBA, and would involve intermediate next-hop changes. We thus chose not to further investigate this approach.

The uniform algorithm works by increasing (or decreasing) the weight of all outgoing links of the modified by the same value at each step. It is somehow similar to a single link modification sequence applied at the granularity of a router and such sequences can be computed using the same algorithms. Uniform sequences naturally preserves the initial routing decisions of the modified node, thus preventing intermediate next-hop changes. Fig 3.3 shows that the overhead in terms of sequence length compared to GBA is usually smaller than the link-by-link heuristic. The gap between GBA and uniform is very narrow for short sequences, but it grows larger for longer sequences (Fig. 3.4). Eventually, worst case sequences reported on Table 3.4b are almost as long as those produced by the link-by-link heuristic.

AGBA adds additional constraints to GBA vectors in order to ensure the same property as uniform updates on intermediate next-hop changes. In a sense, AGBA tends to *uniformize* GBA sequences, while keeping the possibility to perform non-uniform updates as long as it does not impact the routing stability. In fact, Fig 3.4 shows that the distribution of AGBA sequence lengths often is mid-distance between GBA and uniform sequences. Longest sequences are also much shorter compared to those computed by the two previous algorithms. According to Table 3.4c, AGBA produces almost the same proportion of *short enough* sequences as GBA on most networks, making it a realistic alternative to GBA for network operators concerned about intermediate next-hop changes.

Contrary to GBA and AGBA, DGBH does not necessarily yield sequences of minimal length, so that shorter sequences may exist that ensure the same property, namely the absence of intermediate transient loops. However, our results show that this heuristic performs well in practice. More than 99% of all sequences, for every single router shutdown operation and every topology, are of exactly the same length as those produced by GBA (Table 3.4d). Besides, among the sequences whose length is increased, 65% only contain one extra intermediate update, 23% have 2 more elements, and the most stretched sequences, which represent the 12 remaining percent, are increased by 4 additional elements. These could be considered a negligible overhead compared to the total length of the sequences.

Seeing these results, we would recommend that practical deployment of our solutions be based on AGBA, for it offers the best trade-off between sequence length and routing stability. However, should sequence lengths be a problem, DGBH would be an acceptable reduced solution that still prevents any kind transient loops.

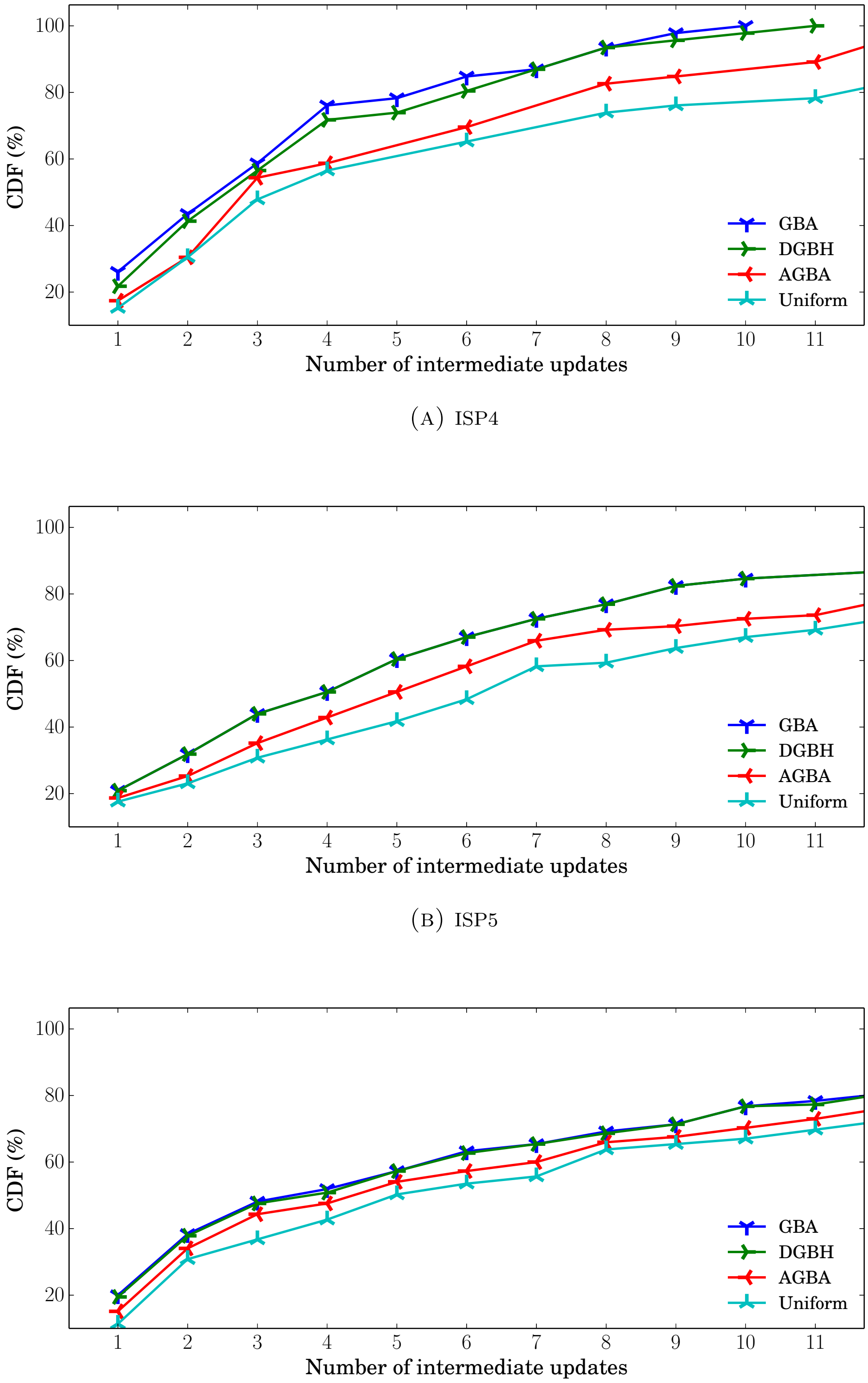


(A) ISP4

(B) ISP5

(C) ISP6

FIGURE 3.3: Sequence length overhead compared to standard GBA



(A) ISP4

(B) ISP5

(C) ISP6

FIGURE 3.4: Length distribution of sequences produced by GBA alternatives

3 Computing times

We showed in the previous section that most intermediate sequences computed using variations of GBA are short enough for practical application. This is however not sufficient to make our approach a realistic solution to be used in production networks. Our algorithms indeed involve heavy graph calculations that could lead to long computing times. In this section, we show that in spite of a polynomial time worst case complexity, actual computing times are very small, and they remain reasonable even on our largest evaluation networks. We also assess the efficiency of the implementation improvements presented at the end of Chapter 2, detailing how each individual improvement affects global performances. These results were obtained using a standard desktop computer architecture based on an *Intel Core 2 Duo (E7200)* CPU with a clock rate of 2.53 GHz.

Since GBA variations, AGBA and DGBH, have a negligible impact on computing times, we focus here on the performances of the basic GBA algorithm. We also merge together the computing times obtained for link and node shutdown operations, for they present no significant differences.

3.1 GBA performances

Table 3.5 displays various statistical information on the time required to compute shutdown sequences. On all but the three largest topologies, worst-case sequences are computed in only a few milliseconds. This duration increases to some tens of milliseconds for ISP5 and Sprint networks, which respectively contain 200 and 315 routers. For ISP6, any sequence can be computed within a couple of seconds, which remains a reasonable duration considering that the topology belongs to one of the biggest Tier-1 Internet providers in the world. Besides, operations for which no transient loop may occur are ignored in these results. Such cases are detected after a few microseconds, even on ISP6, and cause an empty sequence to be generated. We thus envision a practical deployment of our solutions according to one of the following schemes.

- A network management tool pre-computes and stores a sequence for every predictable event. These include links and routers shutdowns, as well as any other operation pre-configured by the operator. When one of these operation, or any other whose associated sequence could be deduced from a stored one, is to be performed in the network, the corresponding sequence could either be automatically applied on the router using a reconfiguration script, or be displayed for the operator to execute it manually.

	Min	Median	Max	Mean	3^{rd} quartile	9^{th} decile
Internet2	0.06 ms	0.06 ms	0.06 ms	0.06 ms	0.06 ms	0.06 ms
Geant	0.21 ms	0.29 ms	0.35 ms	0.28 ms	0.30 ms	0.33 ms
ISP1	0.34 ms	0.39 ms	0.51 ms	0.41 ms	0.47 ms	0.51 ms
ISP2	1.43 ms	1.98 ms	2.68 ms	1.96 ms	2.08 ms	2.67 ms
Renater	0.35 ms	1.33 ms	2.68 ms	1.28 ms	1.48 ms	1.78 ms
ISP3	0.49 ms	6.08 ms	10.91 ms	6.08 ms	7.25 ms	7.75 ms
ISP4	0.99 ms	10.17 ms	18.04 ms	10.18 ms	12.07 ms	12.95 ms
ISP5	0.64 ms	26.58 ms	49.63 ms	23.80 ms	30.01 ms	34.64 ms
ISP6	3.63 ms	**1.59 s**	**2.15 s**	**1.40 s**	**1.70 s**	**1.77 s**
Exodus	0.88 ms	3.16 ms	5.32 ms	3.15 ms	3.90 ms	4.69 ms
EBone	0.33 ms	4.17 ms	7.41 ms	3.87 ms	4.91 ms	6.04 ms
Telstra	5.01 ms	6.42 ms	9.20 ms	6.61 ms	7.53 ms	8.81 ms
AboveNet	0.49 ms	11.13 ms	16.68 ms	10.45 ms	12.37 ms	15.82 ms
Tiscali	0.91 ms	15.31 ms	22.68 ms	14.56 ms	17.26 ms	19.87 ms
Sprint	1.71 ms	68.42 ms	109.47 ms	62.77 ms	76.30 ms	80.17 ms

TABLE 3.5: Computing time statistics for all node and link shutdown operations

- Sequences are pre-computed and stored on a distributed mode, each router being in charge of its own modifications. Then, when a command is passed on the router that would cause a topological modification, the associated sequence is applied.

- Sequences are computed on-the-fly by the router itself at the time a command is passed, and directly processed. This option seems realistic at least on small networks considering how low the computing times are.

3.2 Algorithmic improvements evaluation

Fig. 3.5 shows how much computing time was saved by each of our algorithmic improvements on the three largest topologies. It appear on these figures that, for a given topology, a basic GBA implementation computes most of the intermediate update sequences in a nearly constant time. This duration is of about 9 ms, 25 ms and 1.5 s, respectively for ISP 4, 5 and 6, and corresponds to the time required to check whether transient loop could occur for each destination in the network.

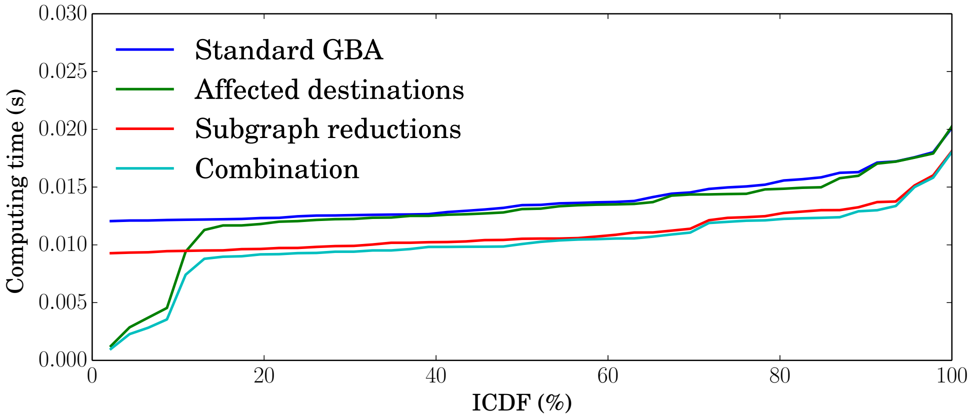
The *affected destinations* improvement lightens this detection phase by pruning from the set of destinations to be checked all nodes that were not initially sending traffic via the modified component. This results in a strong decrease of the computing time for more than 10%, 40% and 35% of the operations for ISP 4, 5 and 6, respectively. However, for operations affecting a large proportion of the destinations, the overhead involved by the preliminary phase may exceed the benefits it provides, having thus slight negative effects in some worst cases.

On another hand, the *subgraph reductions* improves almost all graph related calculations, accelerating both the initial transient loop detection phase and the merged graph update that is performed at each iteration of GBA. As a result, this algorithmic improvement reduces the constant detection time for all operations by 10 to 20%, but also the actual sequence computation time based on how long the sequence is and how complex graph calculation are. It is thus generally more effective on worst cases, as one may notice on Fig 3.5c were the maximum computing times are almost halved.

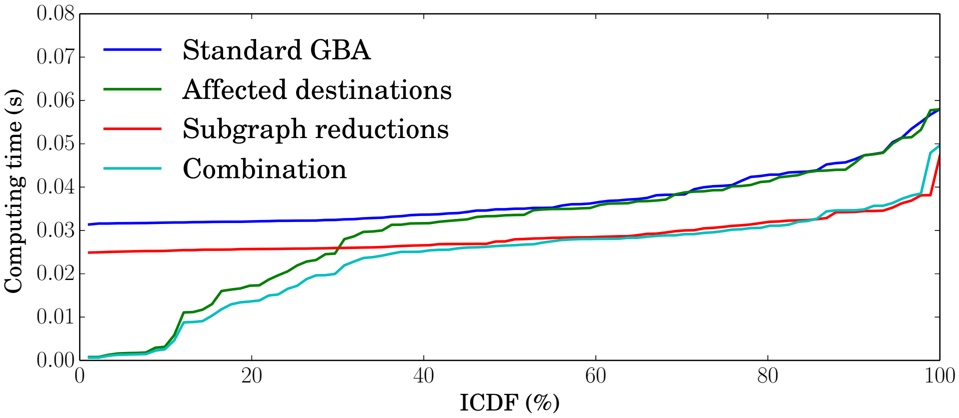
Eventually, both improvements can be used together to combine their positive effects (*combination*), while the overhead of computing impacted destinations fades away as it is also affected by the subgraph reductions.

4 Conclusion

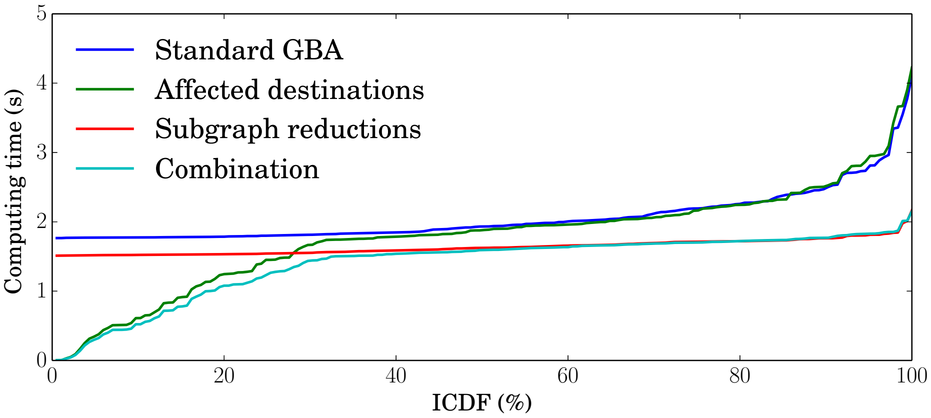
In this chapter, we evaluated various aspects of our solutions on real and inferred ISP topologies or different shapes and sizes. First, we showed that transient loops could occur for a large proportion of link and node shutdown operations. Although these loops are usually limited to a few links, our results reveal that the removal of key components may have a much wider impact on some networks. We then assessed the practicality of GBA, compared to single-link based algorithms. It appears that sequences produced by GBA enable to perform most link and router removal operations with only a few intermediate updates, while *link-by-link* and *uniform* sequences are impractical for router-wide reconfigurations. In addition, we compared the sequence lengths achieved by GBA with its more constrained variations. For almost all operations, DGBH computes sequences of exactly the same length as GBA, ensuring at no cost the absence of intermediate transient loops. On the other hand, the prevention of intermediate forwarding changes, with AGBA, requires a few extra weight increments. Yet these sequences remain significantly shorter than uniform ones. Finally, we evaluated the computing time efficiency of our implementation of GBA. Our results show that, on common hardware, most sequences are computed within tens of milliseconds, while worst cases barely exceed 2 seconds. Such promising results open the way for a deployment of our solutions on production networks.



(A) ISP4



(B) ISP5



(C) ISP6

FIGURE 3.5: Evaluation of implementation improvements on computing time distribution

Conclusion

Over the past decades, the development of the Internet has led to the emergence of real-time applications such as voice over IP, videoconferences and online gaming. These usages raised new concerns regarding the quality of service achieved by Internet Service Providers (ISPs), leading to more and more stringent Service Level Agreements (SLAs). However the protocols currently used to direct and carry traffic through the Internet have been designed in a best effort perspective, and do no allow to guarantee high service availability in the presence of topological changes due to routing events. In particular, the distributed nature of link-state routing protocols implies that transient routing loops may occur after each topological change. The duration of such loops is about one second in practice, increasing transmission delays, and causing congestions as well as packet losses.

In order to enable network operators to perform maintenance operations, and adapt their routing policies in real time according to traffic fluctuations, we proposed practical and incrementally deployable solutions to prevent transient disruptions in case of router-wide modifications. Our proposals do not require any protocol modification, as they only rely on basic functionalities of link-state routing protocols. The approach consists in progressively reconfiguring the weights on the outgoing links of the modified router. Through fine tuning of these weights we may indeed implicitly control the order in which routers impacted by the change update their forwarding decisions, and prevent transient loops. We presented a theoretical framework for this approach, proving that a transient loop free sequence always exists, and defining necessary and sufficient constraints to ensure the prevention of each loop. Based on this framework, we provided an algorithm to compute weight update sequences of minimal length, such that no transient loop can occur between two subsequent updates.

This aim for minimality may jeopardize routing stability, possibly causing transient loops around the modified router that could not have occurred otherwise. To deal with these side effects of our approach, we devised several solutions that achieve several tradeoffs between the level of disruption avoidance and the sequence lengths. While the use of *local delay*, in conjunction with sequences computed by our Greedy Backward Algorithm (GBA), allows to effectively prevent any potential transient loop, it is also possible to adjust the sequences themselves by considering additional constraints during the calculation process. These conditions may either only focus on transient loop avoidance, or enforce routing stability by preventing the use of intermediate forwarding paths. That is, to ensure that intermediate vectors may only lead to the use of PRE or POST edges. We proposed two variations of GBA, respectively called Dynamic Greedy Backward Heuristic (DGBH) and Adjusted Greedy Backward Algorithm (AGBA), to compute short sequences satisfying either set of conditions, along with static *loop-constraints*.

In order to assess the practicality of our approach in a production network, we evaluated various aspects of the solutions we propose on real and inferred ISP topologies. Our results show that transient loops could occur for a large proportion of link and node shutdown operations, and may affect a significant part of the network. This emphasizes the need for efficient means to solve this problem. As such, our proposal performs decently, allowing to safely proceed to most link and router removal operations with only a few intermediate updates, while single-link based algorithms appear impractical to handle router-wide reconfigurations. In addition, we compared the sequence lengths achieved by GBA with its more constrained extensions. For almost all operations, the algorithmic prevention of intermediate transient loops comes at no cost, the length of a sequence computed with DGBH being exactly the same as the one produced by GBA. These sequences could thus be applied regardless of the availability of a local-delay feature on the modified router. Although the sequence stretching caused by the use of AGBA is not negligible, final sequences often remain significantly shorter than minimal uniform ones, while achieving the same guarantees in terms of path stability. Besides, thanks to various algorithmic improvements, we have been able to develop a time-efficient implementation of our solutions. Our evaluations of this implementation show that most sequences can be computed within tens of milliseconds, even on large networks, and the few worst cases do not exceed a couple of seconds. In practice, our algorithms could be implemented in a network management tool, and sequences computed *offline* before being manually applied on the modified router. Yet we hope that such performances, combined with the need for transient loop avoidance mechanisms in ISP networks, will allow for our solutions to be eventually integrated in router software.

Perspectives

Even though we claim our approach to be easily deployable in practice, its actual impact on a production network is yet to be evaluated. For example, practical aspects of link-state routing that we approximated in our theoretical framework, such as network destinations and multipoint links, may complicate the deployment of our solutions. Short-term goals on this subject should thus include extensive analysis in real environments. Currently, one of the main concerns raised by operators regarding this approach is the possible negative interaction with inter-domain routing. That is, the effects of successive ISP weight reconfigurations on Border Gateway Protocol (BGP) convergence. This aspect will be paid a particular attention during the measurement and evaluation campaign we recently started in collaboration with RENATER, as it may also influence the delay between two subsequent updates and the maximum length of a sequence. As for longer sequences, thorough analysis based on real traffic matrices is required to determine which conditions could be relaxed, or which constraints could be ignored, with the least impact on the network. On the contrary, it could also be possible to adapt the network topology itself. In particular, enforcing logical and physical patterns that increase fast re-routing coverage and decrease sequence length. Another source of long sequences is the prevention of intermediate changes, whose practical impact on connection-based transport layer protocols remains unknown. It would thus be interesting to compare the throughput of a TCP-like flow subject to several intermediate route deflections with the one achieved in case of a direct rerouting.

On a more theoretical perspective, and especially if intermediate forwarding changes have a negligible impact on network performances, future works could take interest in formally studying the complexity of the Minimal Intermediate Loop-free Problem (MILP) and, if MILP is in P, finding a P-time algorithm to optimally solve it. Long-term objectives to provide ever shorter sequences would also include investigating the opportunity of performing weight updates of opposite sign to the intended modification. For example, considering link weight decrements before applying an always increasing sequence may, in certain cases, enable to reduce its overall length. In the same spirit, it could be possible to perform weight reconfigurations on links farther away from the modified router. Finally, the approach could be further extended to the more general use cases of Shared Risk Link Groups (SRLGs) and arbitrary k-links modifications anywhere in the network. Again, the formal analysis of the computational complexity of such minimization problems is left open for future works.

List of publications

International journals

[1] F. Clad, S. Vissicchio, P. Merindol, P. Francois and J.-J. Pansiot, “Computing Minimal Update Sequences for Graceful Router-wide Reconfigurations”, *to appear in IEEE/ACM Transactions on Networking*, 2014 (14 pages)

[2] F. Clad, P. Merindol, J.-J. Pansiot, P. Francois and O. Bonaventure, “Graceful Convergence in Link-State IP Networks”, *in IEEE/ACM Transactions on Networking*, Volume 22, Issue 1, February 2014 (13 pages)

International conference with selection committee

[3] F. Clad, P. Merindol, S. Vissicchio, J.-J. Pansiot and P. Francois, “Graceful Router Updates for Link-State Protocols”, *in proceedings of IEEE International Conference on Network Protocols (ICNP'13)*, Goettingen, Germany, October 2013 (10 pages)

National conference with selection committee

[4] F. Clad, P. Merindol and J.-J. Pansiot “L'art de reconfigurer un noeud de routage”, 16$^{\text{èmes}}$ *Rencontres Francophones sur les Aspects Algorithmiques des Télécommunications (AlgoTel'14)*, Le-Bois-Plage-en-Ré, France, June 2014 (4 pages)

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List of Abbreviations

AGBA	Adjusted Greedy Backward Algorithm
AS	Autonomous System
BFD	Bidirectionnal Forwarding Detection
BGP	Border Gateway Protocol
CPC	Change Prevention Condition
CSPF	Contrained Shortest Path First
DAG	Directed Acyclic Graph
DGBH	Dynamic Greedy Backward Heuristic
DLC	Dynamic Loop Constraint
ECMP	Equal-Cost Multi-Path
FIB	Forwarding Information Base
FIFR	Failure Insensitive Fast Rerouting
GBA	Greedy Backward Algorithm
ICMP	Internet Control Message Protocol
IETF	Internet Engineering Task Force
IGMP	Internet Group Management Protocol
IGP	Interior Gateway Protocol
IGRP	Interior Gateway Routing Protocol
IS-IS	Intermediate System to Intermediate System
ISO	International Organization for Standardization

ISP	Internet Service Provider
ISPF	Incremental Shortest Path First
LFA	Loop-Free Alternate
LOF	Loss of Frame
LOS	Loss of Signal
LSA	Link-State Advertisement
LSDB	Link-state Database
LSP	Link-State Packet
LSP	Label Switching Path
MCLP	Minimal intermediate Change-free and Loop-free Problem
MILP	Minimal Intermediate Loop-free Problem
MLP	Minimal Loop-free Problem
MP	Merge Point
MPLS	Multiprotocol Label Switching
MTU	Maximum Transmission Unit
oFIB	ordered FIB
OSPF	Open Shortest Path First
PLR	Point of Local Repair
PLSN	Path Locking via Safe Neighbors
RIB	Routing Information Base
RIP	Routing Internet Protocol
RLFA	Remote Loop-Free Alternate
RSPDAG	Reverse Shortest Path DAG
RSPT	Reverse Shortest Path Tree
SDH	Synchronous Digital Hierarchy
SITN	Ships-in-the-Night
SLA	Service Level Agreement
SONET	Synchronous Optical Networking

SPDAG	Shortest Path DAG
SPT	Shortest Path Tree
SRLG	Shared Risk Link Group
TOS	Type of Service
TTL	time-to-live

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