compared. If they are equal, the duplicate is discarded. If they are different, the older one is thrown out. To guard against errors on the links, all link state packets are acknowledged.

The data structure used by router *B* for the network shown in Fig. 5-12(a) is depicted in Fig. 5-13. Each row here corresponds to a recently arrived, but as yet not fully processed, link state packet. The table records where the packet originated, its sequence number and age, and the data. In addition, there are send and acknowledgement flags for each of *B*'s three links (to *A*, *C*, and *F*, respectively). The send flags mean that the packet must be sent on the indicated link. The acknowledgement flags mean that it must be acknowledged there.

			Send flags			ACK flags			
Source	Seq.	Age	A	C	F	A	C	F	Data
A	21	60	0	1	1	1	0	0	
F	21	60	1	1	0	0	0	1	
E	21	59	0	1	0	1	0	1	
C	20	60	1	0	1	0	1	0	
D	21	59	1	0	0	0	1	1	

Figure 5-13. The packet buffer for router *B* in Fig. 5-12(a).

In Fig. 5-13, the link state packet from *A* arrives directly, so it must be sent to *C* and *F* and acknowledged to *A*, as indicated by the flag bits. Similarly, the packet from *F* has to be forwarded to *A* and *C* and acknowledged to *F*.

However, the situation with the third packet, from *E*, is different. It arrives twice, once via *EAB* and once via *EFB*. Consequently, it has to be sent only to *C* but must be acknowledged to both *A* and *F*, as indicated by the bits.

If a duplicate arrives while the original is still in the buffer, bits have to be changed. For example, if a copy of *C*'s state arrives from *F* before the fourth entry in the table has been forwarded, the six bits will be changed to 100011 to indicate that the packet must be acknowledged to *F* but not sent there.

Computing the New Routes

Once a router has accumulated a full set of link state packets, it can construct the entire network graph because every link is represented. Every link is, in fact, represented twice, once for each direction. The different directions may even have different costs. The shortest-path computations may then find different paths from router *A* to *B* than from router *B* to *A*.

Now Dijkstra's algorithm can be run locally to construct the shortest paths to all possible destinations. The results of this algorithm tell the router which link to

use to reach each destination. This information is installed in the routing tables, and normal operation is resumed.

Compared to distance vector routing, link state routing requires more memory and computation. For a network with n routers, each of which has k neighbors, the memory required to store the input data is proportional to kn, which is at least as large as a routing table listing all the destinations. Also, the computation time grows faster than kn, even with the most efficient data structures, an issue in large networks. Nevertheless, in many practical situations, link state routing works well because it does not suffer from slow convergence problems.

Link state routing is widely used in actual networks, so a few words about some example protocols are in order. Many ISPs use the **IS-IS** (**Intermediate System-Intermediate System**) link state protocol (Oran, 1990). It was designed for an early network called DECnet, later adopted by ISO for use with the OSI protocols and then modified to handle other protocols as well, most notably, IP. **OSPF** (**Open Shortest Path First**) is the other main link state protocol. It was designed by IETF several years after IS-IS and adopted many of the innovations designed for IS-IS. These innovations include a self-stabilizing method of flooding link state updates, the concept of a designated router on a LAN, and the method of computing and supporting path splitting and multiple metrics. As a consequence, there is very little difference between IS-IS and OSPF. The most important difference is that IS-IS can carry information about multiple network layer protocols at the same time (e.g., IP, IPX, and AppleTalk). OSPF does not have this feature, and it is an advantage in large multiprotocol environments. We will go over OSPF in Sec. 5.6.6.

A general comment on routing algorithms is also in order. Link state, distance vector, and other algorithms rely on processing at all the routers to compute routes. Problems with the hardware or software at even a small number of routers can wreak havoc across the network. For example, if a router claims to have a link it does not have or forgets a link it does have, the network graph will be incorrect. If a router fails to forward packets or corrupts them while forwarding them, the route will not work as expected. Finally, if it runs out of memory or does the routing calculation wrong, bad things will happen. As the network grows into the range of tens or hundreds of thousands of nodes, the probability of some router failing occasionally becomes nonnegligible. The trick is to try to arrange to limit the damage when the inevitable happens. Perlman (1988) discusses these problems and their possible solutions in detail.

5.2.6 Hierarchical Routing

As networks grow in size, the router routing tables grow proportionally. Not only is router memory consumed by ever-increasing tables, but more CPU time is needed to scan them and more bandwidth is needed to send status reports about them. At a certain point, the network may grow to the point where it is no longer

feasible for every router to have an entry for every other router, so the routing will have to be done hierarchically, as it is in the telephone network.

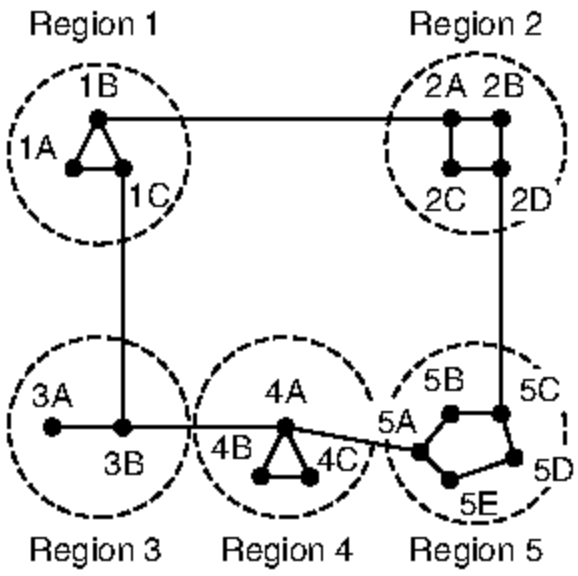
When hierarchical routing is used, the routers are divided into what we will call **regions**. Each router knows all the details about how to route packets to destinations within its own region but knows nothing about the internal structure of other regions. When different networks are interconnected, it is natural to regard each one as a separate region to free the routers in one network from having to know the topological structure of the other ones.

For huge networks, a two-level hierarchy may be insufficient; it may be necessary to group the regions into clusters, the clusters into zones, the zones into groups, and so on, until we run out of names for aggregations. As an example of a multilevel hierarchy, consider how a packet might be routed from Berkeley, California, to Malindi, Kenya. The Berkeley router would know the detailed topology within California but would send all out-of-state traffic to the Los Angeles router. The Los Angeles router would be able to route traffic directly to other domestic routers but would send all foreign traffic to New York. The New York router would be programmed to direct all traffic to the router in the destination country responsible for handling foreign traffic, say, in Nairobi. Finally, the packet would work its way down the tree in Kenya until it got to Malindi.

Figure 5-14 gives a quantitative example of routing in a two-level hierarchy with five regions. The full routing table for router *1A* has 17 entries, as shown in Fig. 5-14(b). When routing is done hierarchically, as in Fig. 5-14(c), there are entries for all the local routers, as before, but all other regions are condensed into a single router, so all traffic for region 2 goes via the *1B-2A* line, but the rest of the remote traffic goes via the *1C-3B* line. Hierarchical routing has reduced the table from 17 to 7 entries. As the ratio of the number of regions to the number of routers per region grows, the savings in table space increase.

Unfortunately, these gains in space are not free. There is a penalty to be paid: increased path length. For example, the best route from *1A* to *5C* is via region 2, but with hierarchical routing all traffic to region 5 goes via region 3, because that is better for most destinations in region 5.

When a single network becomes very large, an interesting question is "how many levels should the hierarchy have?" For example, consider a network with 720 routers. If there is no hierarchy, each router needs 720 routing table entries. If the network is partitioned into 24 regions of 30 routers each, each router needs 30 local entries plus 23 remote entries for a total of 53 entries. If a three-level hierarchy is chosen, with 8 clusters each containing 9 regions of 10 routers, each router needs 10 entries for local routers, 8 entries for routing to other regions within its own cluster, and 7 entries for distant clusters, for a total of 25 entries. Kamoun and Kleinrock (1979) discovered that the optimal number of levels for an N router network is $\ln N$, requiring a total of $e \ln N$ entries per router. They have also shown that the increase in effective mean path length caused by hierarchical routing is sufficiently small that it is usually acceptable.



(a)

Full table for 1A

Dest.	Line	Hops
1A	–	–
1B	1B	1
1C	1C	1
2A	1B	2
2B	1B	3
2C	1B	3
2D	1B	4
3A	1C	3
3B	1C	2
4A	1C	3
4B	1C	4
4C	1C	4
5A	1C	4
5B	1C	5
5C	1B	5
5D	1C	6
5E	1C	5

(b)

Hierarchical table for 1A

Dest.	Line	Hops
1A	–	–
1B	1B	1
1C	1C	1
2	1B	2
3	1C	2
4	1C	3
5	1C	4

(c)

Figure 5-14. Hierarchical routing.

5.2.7 Broadcast Routing

In some applications, hosts need to send messages to many or all other hosts. For example, a service distributing weather reports, stock market updates, or live radio programs might work best by sending to all machines and letting those that are interested read the data. Sending a packet to all destinations simultaneously is called **broadcasting**. Various methods have been proposed for doing it.

One broadcasting method that requires no special features from the network is for the source to simply send a distinct packet to each destination. Not only is the method wasteful of bandwidth and slow, but it also requires the source to have a complete list of all destinations. This method is not desirable in practice, even though it is widely applicable.

An improvement is **multidestination routing**, in which each packet contains either a list of destinations or a bit map indicating the desired destinations. When a packet arrives at a router, the router checks all the destinations to determine the set of output lines that will be needed. (An output line is needed if it is the best route to at least one of the destinations.) The router generates a new copy of the packet for each output line to be used and includes in each packet only those destinations that are to use the line. In effect, the destination set is partitioned among

the output lines. After a sufficient number of hops, each packet will carry only one destination like a normal packet. Multidestination routing is like using separately addressed packets, except that when several packets must follow the same route, one of them pays full fare and the rest ride free. The network bandwidth is therefore used more efficiently. However, this scheme still requires the source to know all the destinations, plus it is as much work for a router to determine where to send one multidestination packet as it is for multiple distinct packets.

We have already seen a better broadcast routing technique: flooding. When implemented with a sequence number per source, flooding uses links efficiently with a decision rule at routers that is relatively simple. Although flooding is ill-suited for ordinary point-to-point communication, it rates serious consideration for broadcasting. However, it turns out that we can do better still once the shortest path routes for regular packets have been computed.

The idea for **reverse path forwarding** is elegant and remarkably simple once it has been pointed out (Dalal and Metcalfe, 1978). When a broadcast packet arrives at a router, the router checks to see if the packet arrived on the link that is normally used for sending packets *toward* the source of the broadcast. If so, there is an excellent chance that the broadcast packet itself followed the best route from the router and is therefore the first copy to arrive at the router. This being the case, the router forwards copies of it onto all links except the one it arrived on. If, however, the broadcast packet arrived on a link other than the preferred one for reaching the source, the packet is discarded as a likely duplicate.

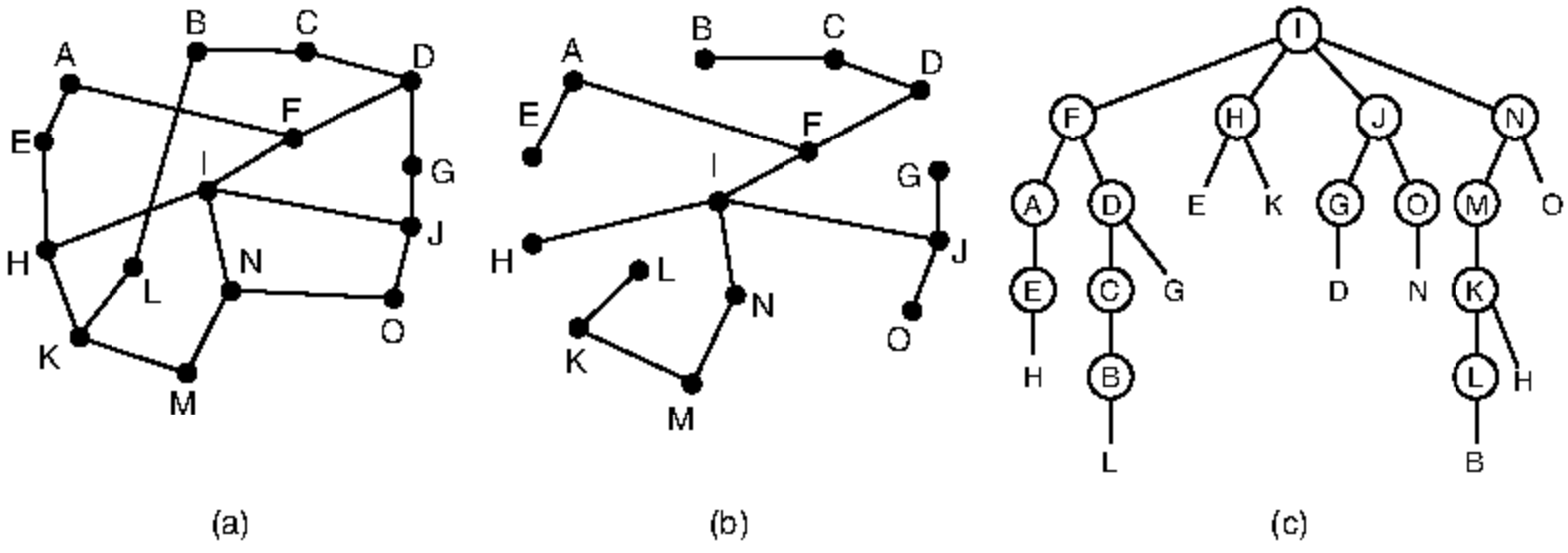


Figure 5-15. Reverse path forwarding. (a) A network. (b) A sink tree. (c) The tree built by reverse path forwarding.

An example of reverse path forwarding is shown in Fig. 5-15. Part (a) shows a network, part (b) shows a sink tree for router *I* of that network, and part (c) shows how the reverse path algorithm works. On the first hop, *I* sends packets to *F*, *H*, *J*, and *N*, as indicated by the second row of the tree. Each of these packets arrives on the preferred path to *I* (assuming that the preferred path falls along the sink tree) and is so indicated by a circle around the letter. On the second hop,

eight packets are generated, two by each of the routers that received a packet on the first hop. As it turns out, all eight of these arrive at previously unvisited routers, and five of these arrive along the preferred line. Of the six packets generated on the third hop, only three arrive on the preferred path (at *C*, *E*, and *K*); the others are duplicates. After five hops and 24 packets, the broadcasting terminates, compared with four hops and 14 packets had the sink tree been followed exactly.

The principal advantage of reverse path forwarding is that it is efficient while being easy to implement. It sends the broadcast packet over each link only once in each direction, just as in flooding, yet it requires only that routers know how to reach all destinations, without needing to remember sequence numbers (or use other mechanisms to stop the flood) or list all destinations in the packet.

Our last broadcast algorithm improves on the behavior of reverse path forwarding. It makes explicit use of the sink tree—or any other convenient spanning tree—for the router initiating the broadcast. A **spanning tree** is a subset of the network that includes all the routers but contains no loops. Sink trees are spanning trees. If each router knows which of its lines belong to the spanning tree, it can copy an incoming broadcast packet onto all the spanning tree lines except the one it arrived on. This method makes excellent use of bandwidth, generating the absolute minimum number of packets necessary to do the job. In Fig. 5-15, for example, when the sink tree of part (b) is used as the spanning tree, the broadcast packet is sent with the minimum 14 packets. The only problem is that each router must have knowledge of some spanning tree for the method to be applicable. Sometimes this information is available (e.g., with link state routing, all routers know the complete topology, so they can compute a spanning tree) but sometimes it is not (e.g., with distance vector routing).

5.2.8 Multicast Routing

Some applications, such as a multiplayer game or live video of a sports event streamed to many viewing locations, send packets to multiple receivers. Unless the group is very small, sending a distinct packet to each receiver is expensive. On the other hand, broadcasting a packet is wasteful if the group consists of, say, 1000 machines on a million-node network, so that most receivers are not interested in the message (or worse yet, they are definitely interested but are not supposed to see it). Thus, we need a way to send messages to well-defined groups that are numerically large in size but small compared to the network as a whole.

Sending a message to such a group is called **multicasting**, and the routing algorithm used is called **multicast routing**. All multicasting schemes require some way to create and destroy groups and to identify which routers are members of a group. How these tasks are accomplished is not of concern to the routing algorithm. For now, we will assume that each group is identified by a multicast address and that routers know the groups to which they belong. We will revisit group membership when we describe the network layer of the Internet in Sec. 5.6.

Multicast routing schemes build on the broadcast routing schemes we have already studied, sending packets along spanning trees to deliver the packets to the members of the group while making efficient use of bandwidth. However, the best spanning tree to use depends on whether the group is dense, with receivers scattered over most of the network, or sparse, with much of the network not belonging to the group. In this section we will consider both cases.

If the group is dense, broadcast is a good start because it efficiently gets the packet to all parts of the network. But broadcast will reach some routers that are not members of the group, which is wasteful. The solution explored by Deering and Cheriton (1990) is to prune the broadcast spanning tree by removing links that do not lead to members. The result is an efficient multicast spanning tree.

As an example, consider the two groups, 1 and 2, in the network shown in Fig. 5-16(a). Some routers are attached to hosts that belong to one or both of these groups, as indicated in the figure. A spanning tree for the leftmost router is shown in Fig. 5-16(b). This tree can be used for broadcast but is overkill for multicast, as can be seen from the two pruned versions that are shown next. In Fig. 5-16(c), all the links that do not lead to hosts that are members of group 1 have been removed. The result is the multicast spanning tree for the leftmost router to send to group 1. Packets are forwarded only along this spanning tree, which is more efficient than the broadcast tree because there are 7 links instead of 10. Fig. 5-16(d) shows the multicast spanning tree after pruning for group 2. It is efficient too, with only five links this time. It also shows that different multicast groups have different spanning trees.

Various ways of pruning the spanning tree are possible. The simplest one can be used if link state routing is used and each router is aware of the complete topology, including which hosts belong to which groups. Each router can then construct its own pruned spanning tree for each sender to the group in question by constructing a sink tree for the sender as usual and then removing all links that do not connect group members to the sink node. **MOSPF** (**Multicast OSPF**) is an example of a link state protocol that works in this way (Moy, 1994).

With distance vector routing, a different pruning strategy can be followed. The basic algorithm is reverse path forwarding. However, whenever a router with no hosts interested in a particular group and no connections to other routers receives a multicast message for that group, it responds with a PRUNE message, telling the neighbor that sent the message not to send it any more multicasts from the sender for that group. When a router with no group members among its own hosts has received such messages on all the lines to which it sends the multicast, it, too, can respond with a PRUNE message. In this way, the spanning tree is recursively pruned. **DVMRP** (**Distance Vector Multicast Routing Protocol**) is an example of a multicast routing protocol that works this way (Waitzman et al., 1988).

Pruning results in efficient spanning trees that use only the links that are actually needed to reach members of the group. One potential disadvantage is that it is lots of work for routers, especially for large networks. Suppose that a network

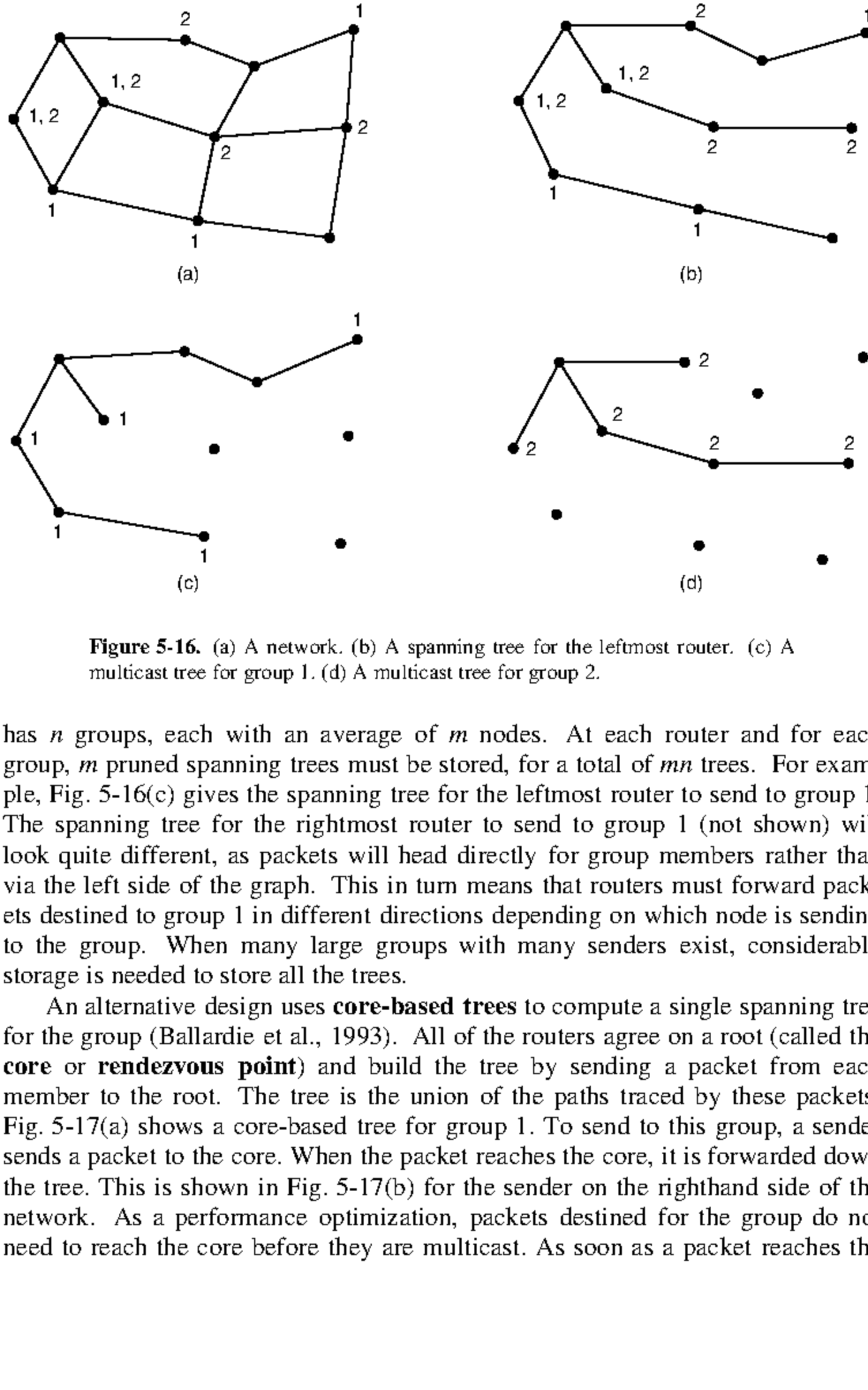


Figure 5-16. (a) A network. (b) A spanning tree for the leftmost router. (c) A multicast tree for group 1. (d) A multicast tree for group 2.

has n groups, each with an average of m nodes. At each router and for each group, m pruned spanning trees must be stored, for a total of mn trees. For example, Fig. 5-16(c) gives the spanning tree for the leftmost router to send to group 1. The spanning tree for the rightmost router to send to group 1 (not shown) will look quite different, as packets will head directly for group members rather than via the left side of the graph. This in turn means that routers must forward packets destined to group 1 in different directions depending on which node is sending to the group. When many large groups with many senders exist, considerable storage is needed to store all the trees.

An alternative design uses **core-based trees** to compute a single spanning tree for the group (Ballardie et al., 1993). All of the routers agree on a root (called the **core** or **rendezvous point**) and build the tree by sending a packet from each member to the root. The tree is the union of the paths traced by these packets. Fig. 5-17(a) shows a core-based tree for group 1. To send to this group, a sender sends a packet to the core. When the packet reaches the core, it is forwarded down the tree. This is shown in Fig. 5-17(b) for the sender on the righthand side of the network. As a performance optimization, packets destined for the group do not need to reach the core before they are multicast. As soon as a packet reaches the

tree, it can be forwarded up toward the root, as well as down all the other branches. This is the case for the sender at the top of Fig. 5-17(b).

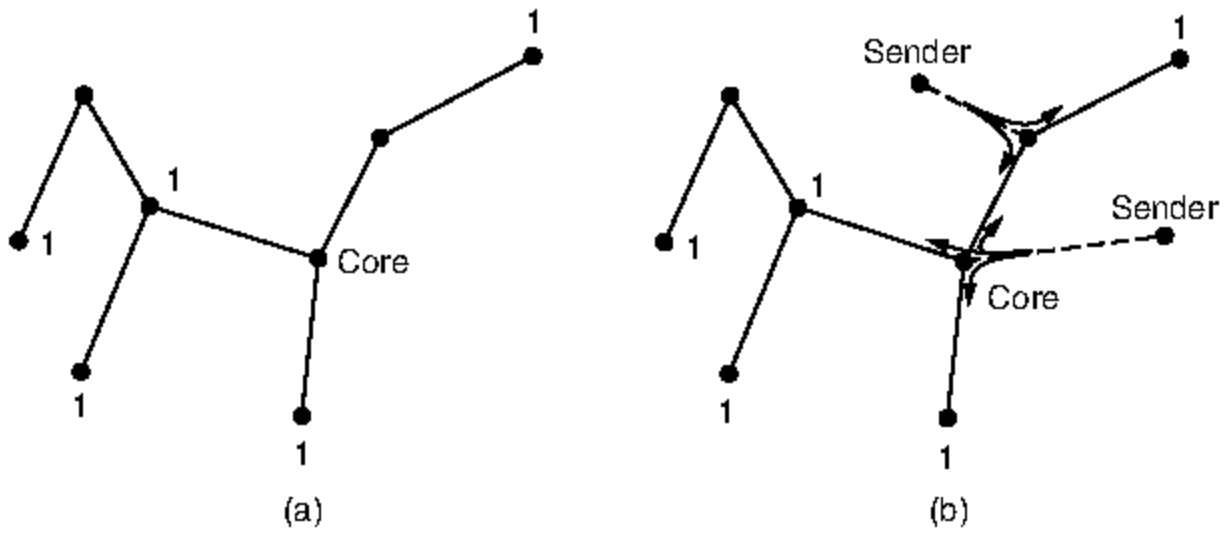


Figure 5-17. (a) Core-based tree for group 1. (b) Sending to group 1.

Having a shared tree is not optimal for all sources. For example, in Fig. 5-17(b), the packet from the sender on the righthand side reaches the top-right group member via the core in three hops, instead of directly. The inefficiency depends on where the core and senders are located, but often it is reasonable when the core is in the middle of the senders. When there is only a single sender, as in a video that is streamed to a group, using the sender as the core is optimal.

Also of note is that shared trees can be a major savings in storage costs, messages sent, and computation. Each router has to keep only one tree per group, instead of m trees. Further, routers that are not part of the tree do no work at all to support the group. For this reason, shared tree approaches like core-based trees are used for multicasting to sparse groups in the Internet as part of popular protocols such as **PIM** (**Protocol Independent Multicast**) (Fenner et al., 2006).

5.2.9 Anycast Routing

So far, we have covered delivery models in which a source sends to a single destination (called **unicast**), to all destinations (called broadcast), and to a group of destinations (called multicast). Another delivery model, called **anycast** is sometimes also useful. In anycast, a packet is delivered to the nearest member of a group (Partridge et al., 1993). Schemes that find these paths are called **anycast routing**.

Why would we want anycast? Sometimes nodes provide a service, such as time of day or content distribution for which it is getting the right information all that matters, not the node that is contacted; any node will do. For example, anycast is used in the Internet as part of DNS, as we will see in Chap. 7.

Luckily, we will not have to devise new routing schemes for anycast because regular distance vector and link state routing can produce anycast routes. Suppose

we want to anycast to the members of group 1. They will all be given the address "1," instead of different addresses. Distance vector routing will distribute vectors as usual, and nodes will choose the shortest path to destination 1. This will result in nodes sending to the nearest instance of destination 1. The routes are shown in Fig. 5-18(a). This procedure works because the routing protocol does not realize that there are multiple instances of destination 1. That is, it believes that all the instances of node 1 are the same node, as in the topology shown in Fig. 5-18(b).

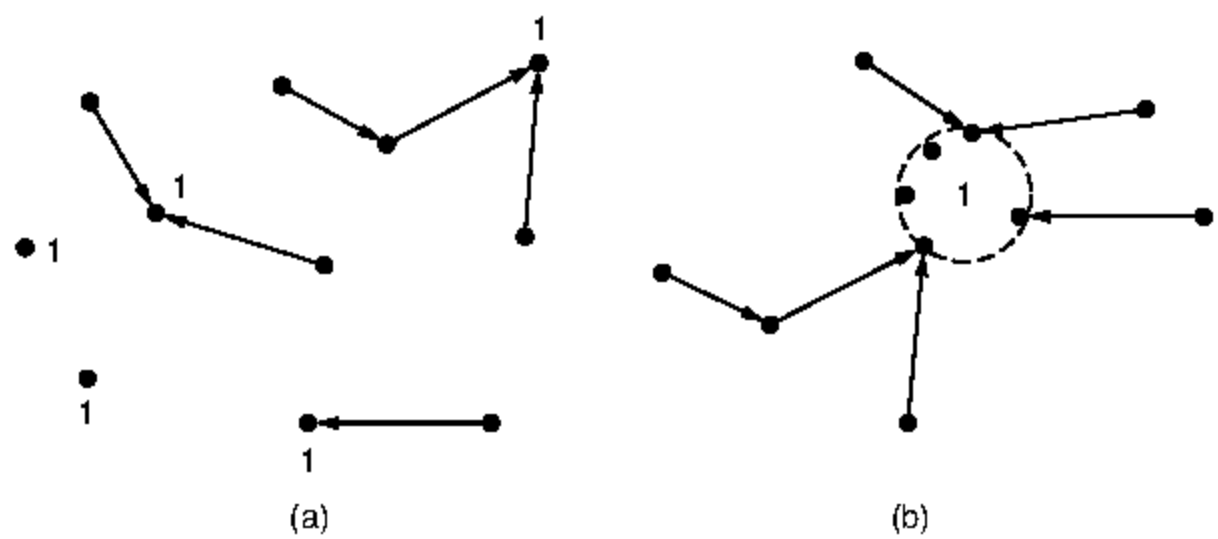


Figure 5-18. (a) Anycast routes to group 1. (b) Topology seen by the routing protocol.

This procedure works for link state routing as well, although there is the added consideration that the routing protocol must not find seemingly short paths that pass through node 1. This would result in jumps through hyperspace, since the instances of node 1 are really nodes located in different parts of the network. However, link state protocols already make this distinction between routers and hosts. We glossed over this fact earlier because it was not needed for our discussion.

5.2.10 Routing for Mobile Hosts

Millions of people use computers while on the go, from truly mobile situations with wireless devices in moving cars, to nomadic situations in which laptop computers are used in a series of different locations. We will use the term **mobile hosts** to mean either category, as distinct from stationary hosts that never move. Increasingly, people want to stay connected wherever in the world they may be, as easily as if they were at home. These mobile hosts introduce a new complication: to route a packet to a mobile host, the network first has to find it.

The model of the world that we will consider is one in which all hosts are assumed to have a permanent **home location** that never changes. Each hosts also has a permanent home address that can be used to determine its home location, analogous to the way the telephone number 1-212-5551212 indicates the United States (country code 1) and Manhattan (212). The routing goal in systems with

mobile hosts is to make it possible to send packets to mobile hosts using their fixed home addresses and have the packets efficiently reach them wherever they may be. The trick, of course, is to find them.

Some discussion of this model is in order. A different model would be to recompute routes as the mobile host moves and the topology changes. We could then simply use the routing schemes described earlier in this section. However, with a growing number of mobile hosts, this model would soon lead to the entire network endlessly computing new routes. Using the home addresses greatly reduces this burden.

Another alternative would be to provide mobility above the network layer, which is what typically happens with laptops today. When they are moved to new Internet locations, laptops acquire new network addresses. There is no association between the old and new addresses; the network does not know that they belonged to the same laptop. In this model, a laptop can be used to browse the Web, but other hosts cannot send packets to it (for example, for an incoming call), without building a higher layer location service, for example, signing into Skype *again* after moving. Moreover, connections cannot be maintained while the host is moving; new connections must be started up instead. Network-layer mobility is useful to fix these problems.

The basic idea used for mobile routing in the Internet and cellular networks is for the mobile host to tell a host at the home location where it is now. This host, which acts on behalf of the mobile host, is called the **home agent**. Once it knows where the mobile host is currently located, it can forward packets so that they are delivered.

Fig. 5-19 shows mobile routing in action. A sender in the northwest city of Seattle wants to send a packet to a host normally located across the United States in New York. The case of interest to us is when the mobile host is not at home. Instead, it is temporarily in San Diego.

The mobile host in San Diego must acquire a local network address before it can use the network. This happens in the normal way that hosts obtain network addresses; we will cover how this works for the Internet later in this chapter. The local address is called a **care of address**. Once the mobile host has this address, it can tell its home agent where it is now. It does this by sending a registration message to the home agent (step 1) with the care of address. The message is shown with a dashed line in Fig. 5-19 to indicate that it is a control message, not a data message.

Next, the sender sends a data packet to the mobile host using its permanent address (step 2). This packet is routed by the network to the host's home location because that is where the home address belongs. In New York, the home agent intercepts this packet because the mobile host is away from home. It then wraps or **encapsulates** the packet with a new header and sends this bundle to the care of address (step 3). This mechanism is called **tunneling**. It is very important in the Internet so we will look at it in more detail later.

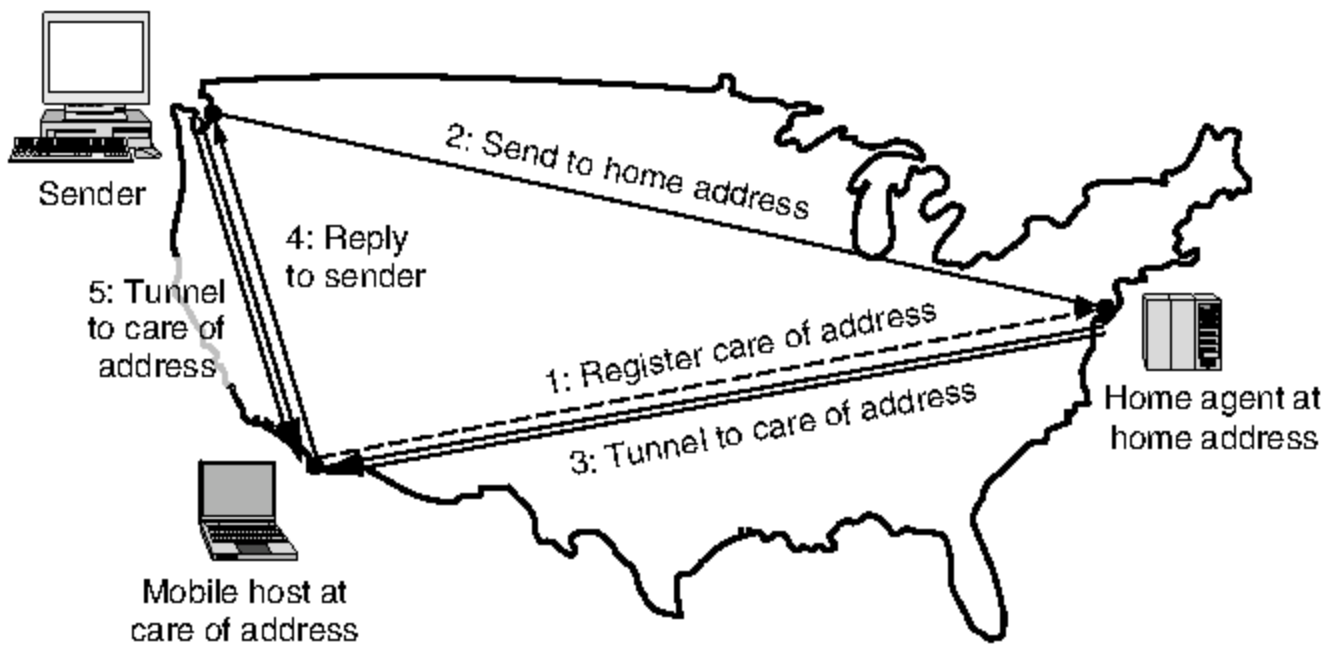


Figure 5-19. Packet routing for mobile hosts.

When the encapsulated packet arrives at the care of address, the mobile host unwraps it and retrieves the packet from the sender. The mobile host then sends its reply packet directly to the sender (step 4). The overall route is called **triangle routing** because it may be circuitous if the remote location is far from the home location. As part of step 4, the sender may learn the current care of address. Subsequent packets can be routed directly to the mobile host by tunneling them to the care of address (step 5), bypassing the home location entirely. If connectivity is lost for any reason as the mobile moves, the home address can always be used to reach the mobile.

An important aspect that we have omitted from this description is security. In general, when a host or router gets a message of the form "Starting right now, please send all of Stephany's mail to me," it might have a couple of questions about whom it is talking to and whether this is a good idea. Security information is included in the messages so that their validity can be checked with cryptographic protocols that we will study in Chap. 8.

There are many variations on mobile routing. The scheme above is modeled on IPv6 mobility, the form of mobility used in the Internet (Johnson et al., 2004) and as part of IP-based cellular networks such as UMTS. We showed the sender to be a stationary node for simplicity, but the designs let both nodes be mobile hosts. Alternatively, the host may be part of a mobile network, for example a computer in a plane. Extensions of the basic scheme support mobile networks with no work on the part of the hosts (Devarapalli et al., 2005).

Some schemes make use of a foreign (i.e., remote) agent, similar to the home agent but at the foreign location, or analogous to the VLR (Visitor Location Register) in cellular networks. However, in more recent schemes, the foreign agent is not needed; mobile hosts act as their own foreign agents. In either case, knowledge of the temporary location of the mobile host is limited to a small number of

hosts (e.g., the mobile, home agent, and senders) so that the many routers in a large network do not need to recompute routes.

For more information about mobile routing, see also Perkins (1998, 2002) and Snoeren and Balakrishnan (2000).

5.2.11 Routing in Ad Hoc Networks

We have now seen how to do routing when the hosts are mobile but the routers are fixed. An even more extreme case is one in which the routers themselves are mobile. Among the possibilities are emergency workers at an earthquake site, military vehicles on a battlefield, a fleet of ships at sea, or a gathering of people with laptop computers in an area lacking 802.11.

In all these cases, and others, each node communicates wirelessly and acts as both a host and a router. Networks of nodes that just happen to be near each other are called **ad hoc networks** or **MANETs** (**Mobile Ad hoc NETworks**). Let us now examine them briefly. More information can be found in Perkins (2001).

What makes ad hoc networks different from wired networks is that the topology is suddenly tossed out the window. Nodes can come and go or appear in new places at the drop of a bit. With a wired network, if a router has a valid path to some destination, that path continues to be valid barring failures, which are hopefully rare. With an ad hoc network, the topology may be changing all the time, so the desirability and even the validity of paths can change spontaneously without warning. Needless to say, these circumstances make routing in ad hoc networks more challenging than routing in their fixed counterparts.

Many, many routing algorithms for ad hoc networks have been proposed. However, since ad hoc networks have been little used in practice compared to mobile networks, it is unclear which of these protocols are most useful. As an example, we will look at one of the most popular routing algorithms, **AODV** (**Ad hoc On-demand Distance Vector**) (Perkins and Royer, 1999). It is a relative of the distance vector algorithm that has been adapted to work in a mobile environment, in which nodes often have limited bandwidth and battery lifetimes. Let us now see how it discovers and maintains routes.

Route Discovery

In AODV, routes to a destination are discovered on demand, that is, only when a somebody wants to send a packet to that destination. This saves much work that would otherwise be wasted when the topology changes before the route is used. At any instant, the topology of an ad hoc network can be described by a graph of connected nodes. Two nodes are connected (i.e., have an arc between them in the graph) if they can communicate directly using their radios. A basic but adequate model that is sufficient for our purposes is that each node can communicate with all other nodes that lie within its coverage circle. Real networks are

more complicated, with buildings, hills, and other obstacles that block communication, and nodes for which *A* is connected to *B* but *B* is not connected to *A* because *A* has a more powerful transmitter than *B*. However, for simplicity, we will assume all connections are symmetric.

To describe the algorithm, consider the newly formed ad hoc network of Fig. 5-20. Suppose that a process at node *A* wants to send a packet to node *I*. The AODV algorithm maintains a distance vector table at each node, keyed by destination, giving information about that destination, including the neighbor to which to send packets to reach the destination. First, *A* looks in its table and does not find an entry for *I*. It now has to discover a route to *I*. This property of discovering routes only when they are needed is what makes this algorithm "on demand."

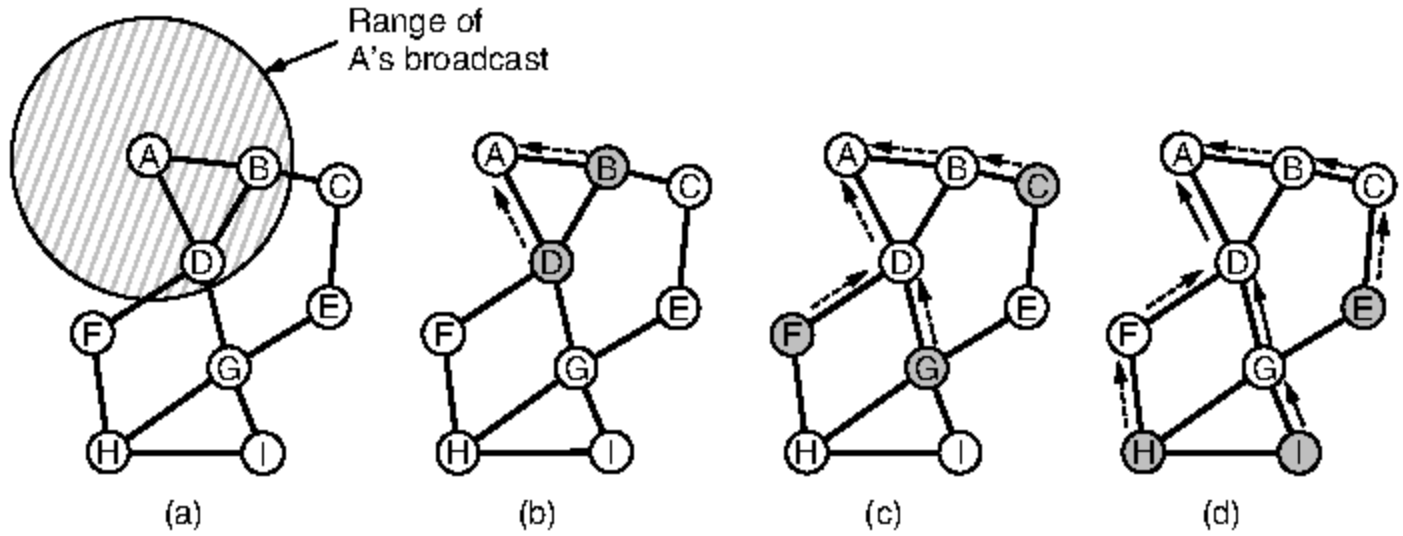


Figure 5-20. (a) Range of *A*'s broadcast. (b) After *B* and *D* receive it. (c) After *C*, *F*, and *G* receive it. (d) After *E*, *H*, and *I* receive it. The shaded nodes are new recipients. The dashed lines show possible reverse routes. The solid lines show the discovered route.

To locate *I*, *A* constructs a ROUTE REQUEST packet and broadcasts it using flooding, as described in Sec. 5.2.3. The transmission from *A* reaches *B* and *D*, as illustrated in Fig. 5-20(a). Each node rebroadcasts the request, which continues to reach nodes *F*, *G*, and *C* in Fig. 5-20(c) and nodes *H*, *E*, and *I* in Fig. 5-20(d). A sequence number set at the source is used to weed out duplicates during the flood. For example, *D* discards the transmission from *B* in Fig. 5-20(c) because it has already forwarded the request.

Eventually, the request reaches node *I*, which constructs a ROUTE REPLY packet. This packet is unicast to the sender along the reverse of the path followed by the request. For this to work, each intermediate node must remember the node that sent it the request. The arrows in Fig. 5-20(b)–(d) show the reverse route information that is stored. Each intermediate node also increments a hop count as it forwards the reply. This tells the nodes how far they are from the destination. The replies tell each intermediate node which neighbor to use to reach the destination: it is the node that sent them the reply. Intermediate nodes *G* and *D* put the

best route they hear into their routing tables as they process the reply. When the reply reaches *A*, a new route, *ADGI*, has been created.

In a large network, the algorithm generates many broadcasts, even for destinations that are close by. To reduce overhead, the scope of the broadcasts is limited using the IP packet's *Time to live* field. This field is initialized by the sender and decremented on each hop. If it hits 0, the packet is discarded instead of being broadcast. The route discovery process is then modified as follows. To locate a destination, the sender broadcasts a ROUTE REQUEST packet with *Time to live* set to 1. If no response comes back within a reasonable time, another one is sent, this time with *Time to live* set to 2. Subsequent attempts use 3, 4, 5, etc. In this way, the search is first attempted locally, then in increasingly wider rings.

Route Maintenance

Because nodes can move or be switched off, the topology can change spontaneously. For example, in Fig. 5-20, if *G* is switched off, *A* will not realize that the route it was using to *I* (*ADGI*) is no longer valid. The algorithm needs to be able to deal with this. Periodically, each node broadcasts a *Hello* message. Each of its neighbors is expected to respond to it. If no response is forthcoming, the broadcaster knows that that neighbor has moved out of range or failed and is no longer connected to it. Similarly, if it tries to send a packet to a neighbor that does not respond, it learns that the neighbor is no longer available.

This information is used to purge routes that no longer work. For each possible destination, each node, *N*, keeps track of its active neighbors that have fed it a packet for that destination during the last ΔT seconds. When any of *N*'s neighbors becomes unreachable, it checks its routing table to see which destinations have routes using the now-gone neighbor. For each of these routes, the active neighbors are informed that their route via *N* is now invalid and must be purged from their routing tables. In our example, *D* purges its entries for *G* and *I* from its routing table and notifies *A*, which purges its entry for *I*. In the general case, the active neighbors tell their active neighbors, and so on, recursively, until all routes depending on the now-gone node are purged from all routing tables.

At this stage, the invalid routes have been purged from the network, and senders can find new, valid routes by using the discovery mechanism that we described. However, there is a complication. Recall that distance vector protocols can suffer from slow convergence or count-to-infinity problems after a topology change in which they confuse old, invalid routes with new, valid routes.

To ensure rapid convergence, routes include a sequence number that is controlled by the destination. The destination sequence number is like a logical clock. The destination increments it every time that it sends a fresh ROUTE REPLY. Senders ask for a fresh route by including in the ROUTE REQUEST the destination sequence number of the last route they used, which will either be the sequence number of the route that was just purged, or 0 as an initial value. The

request will be broadcast until a route with a higher sequence number is found. Intermediate nodes store the routes that have a higher sequence number, or the fewest hops for the current sequence number.

In the spirit of an on demand protocol, intermediate nodes only store the routes that are in use. Other route information learned during broadcasts is timed out after a short delay. Discovering and storing only the routes that are used helps to save bandwidth and battery life compared to a standard distance vector protocol that periodically broadcasts updates.

So far, we have considered only a single route, from *A* to *I*. To further save resources, route discovery and maintenance are shared when routes overlap. For instance, if *B* also wants to send packets to *I*, it will perform route discovery. However, in this case the request will first reach *D*, which already has a route to *I*. Node *D* can then generate a reply to tell *B* the route without any additional work being required.

There are many other ad hoc routing schemes. Another well-known on demand scheme is DSR (Dynamic Source Routing) (Johnson et al., 2001). A different strategy based on geography is explored by GPSR (Greedy Perimeter Stateless Routing) (Karp and Kung, 2000). If all nodes know their geographic positions, forwarding to a destination can proceed without route computation by simply heading in the right direction and circling back to escape any dead ends. Which protocols win out will depend on the kinds of ad hoc networks that prove useful in practice.

5.3 CONGESTION CONTROL ALGORITHMS

Too many packets present in (a part of) the network causes packet delay and loss that degrades performance. This situation is called **congestion**. The network and transport layers share the responsibility for handling congestion. Since congestion occurs within the network, it is the network layer that directly experiences it and must ultimately determine what to do with the excess packets. However, the most effective way to control congestion is to reduce the load that the transport layer is placing on the network. This requires the network and transport layers to work together. In this chapter we will look at the network aspects of congestion. In Chap. 6, we will complete the topic by covering the transport aspects of congestion.

Figure 5-21 depicts the onset of congestion. When the number of packets hosts send into the network is well within its carrying capacity, the number delivered is proportional to the number sent. If twice as many are sent, twice as many are delivered. However, as the offered load approaches the carrying capacity, bursts of traffic occasionally fill up the buffers inside routers and some packets are lost. These lost packets consume some of the capacity, so the number of delivered packets falls below the ideal curve. The network is now congested.

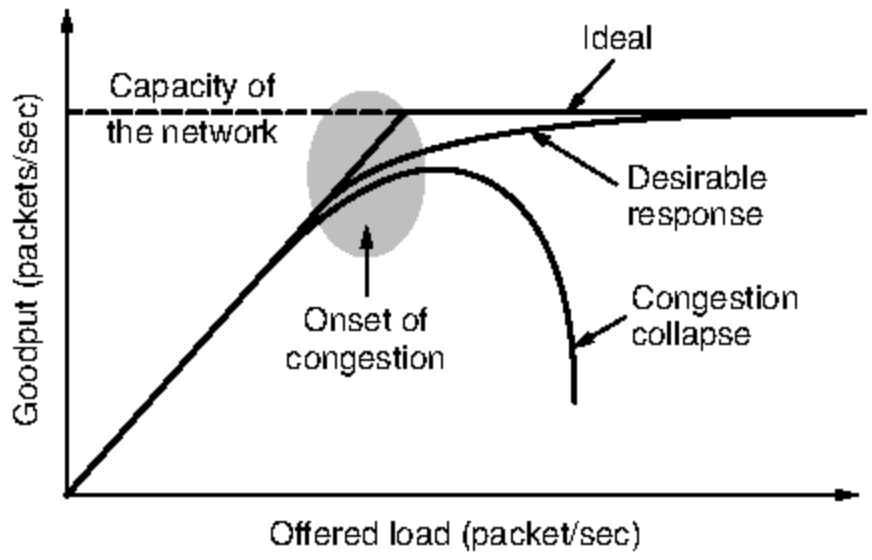


Figure 5-21. With too much traffic, performance drops sharply.

Unless the network is well designed, it may experience a **congestion collapse**, in which performance plummets as the offered load increases beyond the capacity. This can happen because packets can be sufficiently delayed inside the network that they are no longer useful when they leave the network. For example, in the early Internet, the time a packet spent waiting for a backlog of packets ahead of it to be sent over a slow 56-kbps link could reach the maximum time it was allowed to remain in the network. It then had to be thrown away. A different failure mode occurs when senders retransmit packets that are greatly delayed, thinking that they have been lost. In this case, copies of the same packet will be delivered by the network, again wasting its capacity. To capture these factors, the y-axis of Fig. 5-21 is given as **goodput**, which is the rate at which *useful* packets are delivered by the network.

We would like to design networks that avoid congestion where possible and do not suffer from congestion collapse if they do become congested. Unfortunately, congestion cannot wholly be avoided. If all of a sudden, streams of packets begin arriving on three or four input lines and all need the same output line, a queue will build up. If there is insufficient memory to hold all of them, packets will be lost. Adding more memory may help up to a point, but Nagle (1987) realized that if routers have an infinite amount of memory, congestion gets worse, not better. This is because by the time packets get to the front of the queue, they have already timed out (repeatedly) and duplicates have been sent. This makes matters worse, not better—it leads to congestion collapse.

Low-bandwidth links or routers that process packets more slowly than the line rate can also become congested. In this case, the situation can be improved by directing some of the traffic away from the bottleneck to other parts of the network. Eventually, however, all regions of the network will be congested. In this situation, there is no alternative but to shed load or build a faster network.

It is worth pointing out the difference between congestion control and flow control, as the relationship is a very subtle one. Congestion control has to do with

making sure the network is able to carry the offered traffic. It is a global issue, involving the behavior of all the hosts and routers. Flow control, in contrast, relates to the traffic between a particular sender and a particular receiver. Its job is to make sure that a fast sender cannot continually transmit data faster than the receiver is able to absorb it.

To see the difference between these two concepts, consider a network made up of 100-Gbps fiber optic links on which a supercomputer is trying to force feed a large file to a personal computer that is capable of handling only 1 Gbps. Although there is no congestion (the network itself is not in trouble), flow control is needed to force the supercomputer to stop frequently to give the personal computer a chance to breathe.

At the other extreme, consider a network with 1-Mbps lines and 1000 large computers, half of which are trying to transfer files at 100 kbps to the other half. Here, the problem is not that of fast senders overpowering slow receivers, but that the total offered traffic exceeds what the network can handle.

The reason congestion control and flow control are often confused is that the best way to handle both problems is to get the host to slow down. Thus, a host can get a "slow down" message either because the receiver cannot handle the load or because the network cannot handle it. We will come back to this point in Chap. 6.

We will start our study of congestion control by looking at the approaches that can be used at different time scales. Then we will look at approaches to preventing congestion from occurring in the first place, followed by approaches for coping with it once it has set in.

5.3.1 Approaches to Congestion Control

The presence of congestion means that the load is (temporarily) greater than the resources (in a part of the network) can handle. Two solutions come to mind: increase the resources or decrease the load. As shown in Fig. 5-22, these solutions are usually applied on different time scales to either prevent congestion or react to it once it has occurred.

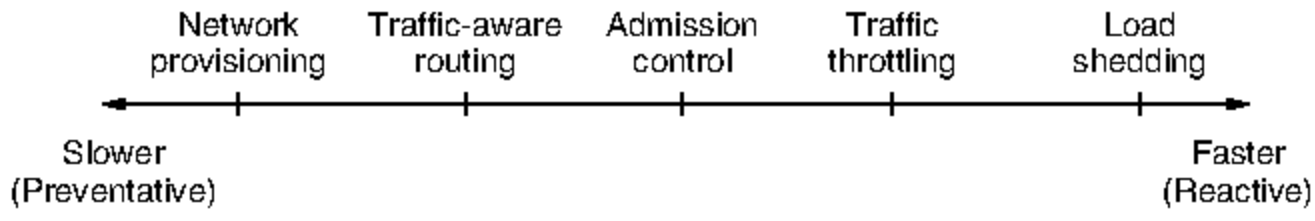


Figure 5-22. Timescales of approaches to congestion control.

The most basic way to avoid congestion is to build a network that is well matched to the traffic that it carries. If there is a low-bandwidth link on the path along which most traffic is directed, congestion is likely. Sometimes resources

can be added dynamically when there is serious congestion, for example, turning on spare routers or enabling lines that are normally used only as backups (to make the system fault tolerant) or purchasing bandwidth on the open market. More often, links and routers that are regularly heavily utilized are upgraded at the earliest opportunity. This is called **provisioning** and happens on a time scale of months, driven by long-term traffic trends.

To make the most of the existing network capacity, routes can be tailored to traffic patterns that change during the day as network users wake and sleep in different time zones. For example, routes may be changed to shift traffic away from heavily used paths by changing the shortest path weights. Some local radio stations have helicopters flying around their cities to report on road congestion to make it possible for their mobile listeners to route their packets (cars) around hotspots. This is called **traffic-aware routing**. Splitting traffic across multiple paths is also helpful.

However, sometimes it is not possible to increase capacity. The only way then to beat back the congestion is to decrease the load. In a virtual-circuit network, new connections can be refused if they would cause the network to become congested. This is called **admission control**.

At a finer granularity, when congestion is imminent the network can deliver feedback to the sources whose traffic flows are responsible for the problem. The network can request these sources to throttle their traffic, or it can slow down the traffic itself.

Two difficulties with this approach are how to identify the onset of congestion, and how to inform the source that needs to slow down. To tackle the first issue, routers can monitor the average load, queueing delay, or packet loss. In all cases, rising numbers indicate growing congestion.

To tackle the second issue, routers must participate in a feedback loop with the sources. For a scheme to work correctly, the time scale must be adjusted carefully. If every time two packets arrive in a row, a router yells STOP and every time a router is idle for 20 μsec, it yells GO, the system will oscillate wildly and never converge. On the other hand, if it waits 30 minutes to make sure before saying anything, the congestion-control mechanism will react too sluggishly to be of any use. Delivering timely feedback is a nontrivial matter. An added concern is having routers send more messages when the network is already congested.

Finally, when all else fails, the network is forced to discard packets that it cannot deliver. The general name for this is **load shedding**. A good policy for choosing which packets to discard can help to prevent congestion collapse.

5.3.2 Traffic-Aware Routing

The first approach we will examine is traffic-aware routing. The routing schemes we looked at in Sec 5.2 used fixed link weights. These schemes adapted to changes in topology, but not to changes in load. The goal in taking load into

account when computing routes is to shift traffic away from hotspots that will be the first places in the network to experience congestion.

The most direct way to do this is to set the link weight to be a function of the (fixed) link bandwidth and propagation delay plus the (variable) measured load or average queuing delay. Least-weight paths will then favor paths that are more lightly loaded, all else being equal.

Traffic-aware routing was used in the early Internet according to this model (Khanna and Zinky, 1989). However, there is a peril. Consider the network of Fig. 5-23, which is divided into two parts, East and West, connected by two links, *CF* and *EI*. Suppose that most of the traffic between East and West is using link *CF*, and, as a result, this link is heavily loaded with long delays. Including queueing delay in the weight used for the shortest path calculation will make *EI* more attractive. After the new routing tables have been installed, most of the East-West traffic will now go over *EI*, loading this link. Consequently, in the next update, *CF* will appear to be the shortest path. As a result, the routing tables may oscillate wildly, leading to erratic routing and many potential problems.

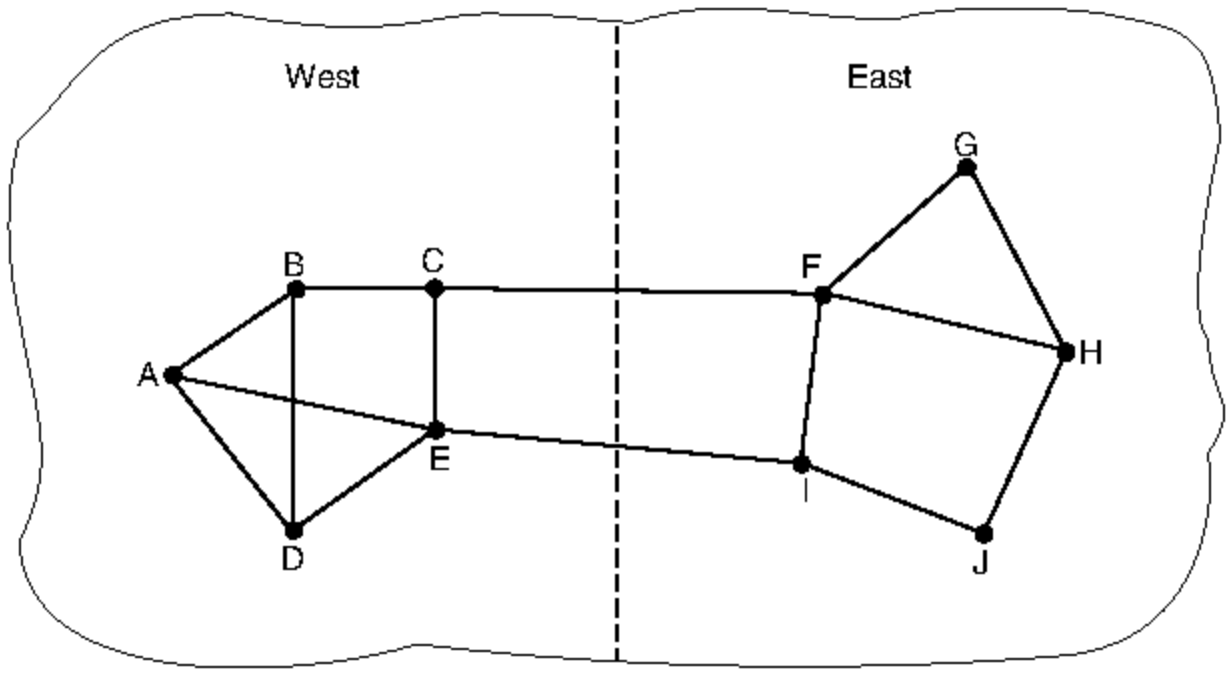


Figure 5-23. A network in which the East and West parts are connected by two links.

If load is ignored and only bandwidth and propagation delay are considered, this problem does not occur. Attempts to include load but change weights within a narrow range only slow down routing oscillations. Two techniques can contribute to a successful solution. The first is multipath routing, in which there can be multiple paths from a source to a destination. In our example this means that the traffic can be spread across both of the East to West links. The second one is for the routing scheme to shift traffic across routes slowly enough that it is able to converge, as in the scheme of Gallagher (1977).

Given these difficulties, in the Internet routing protocols do not generally adjust their routes depending on the load. Instead, adjustments are made outside the routing protocol by slowly changing its inputs. This is called **traffic engineering**.

5.3.3 Admission Control

One technique that is widely used in virtual-circuit networks to keep congestion at bay is **admission control**. The idea is simple: do not set up a new virtual circuit unless the network can carry the added traffic without becoming congested. Thus, attempts to set up a virtual circuit may fail. This is better than the alternative, as letting more people in when the network is busy just makes matters worse. By analogy, in the telephone system, when a switch gets overloaded it practices admission control by not giving dial tones.

The trick with this approach is working out when a new virtual circuit will lead to congestion. The task is straightforward in the telephone network because of the fixed bandwidth of calls (64 kbps for uncompressed audio). However, virtual circuits in computer networks come in all shapes and sizes. Thus, the circuit must come with some characterization of its traffic if we are to apply admission control.

Traffic is often described in terms of its rate and shape. The problem of how to describe it in a simple yet meaningful way is difficult because traffic is typically bursty—the average rate is only half the story. For example, traffic that varies while browsing the Web is more difficult to handle than a streaming movie with the same long-term throughput because the bursts of Web traffic are more likely to congest routers in the network. A commonly used descriptor that captures this effect is the **leaky bucket** or **token bucket**. A leaky bucket has two parameters that bound the average rate and the instantaneous burst size of traffic. Since leaky buckets are widely used for quality of service, we will go over them in detail in Sec. 5.4.

Armed with traffic descriptions, the network can decide whether to admit the new virtual circuit. One possibility is for the network to reserve enough capacity along the paths of each of its virtual circuits that congestion will not occur. In this case, the traffic description is a service agreement for what the network will guarantee its users. We have prevented congestion but veered into the related topic of quality of service a little too early; we will return to it in the next section.

Even without making guarantees, the network can use traffic descriptions for admission control. The task is then to estimate how many circuits will fit within the carrying capacity of the network without congestion. Suppose that virtual circuits that may blast traffic at rates up to 10 Mbps all pass through the same 100-Mbps physical link. How many circuits should be admitted? Clearly, 10 circuits can be admitted without risking congestion, but this is wasteful in the normal case since it may rarely happen that all 10 are transmitting full blast at the same time. In real networks, measurements of past behavior that capture the statistics of transmissions can be used to estimate the number of circuits to admit, to trade better performance for acceptable risk.

Admission control can also be combined with traffic-aware routing by considering routes around traffic hotspots as part of the setup procedure. For example,

consider the network illustrated in Fig. 5-24(a), in which two routers are congested, as indicated.

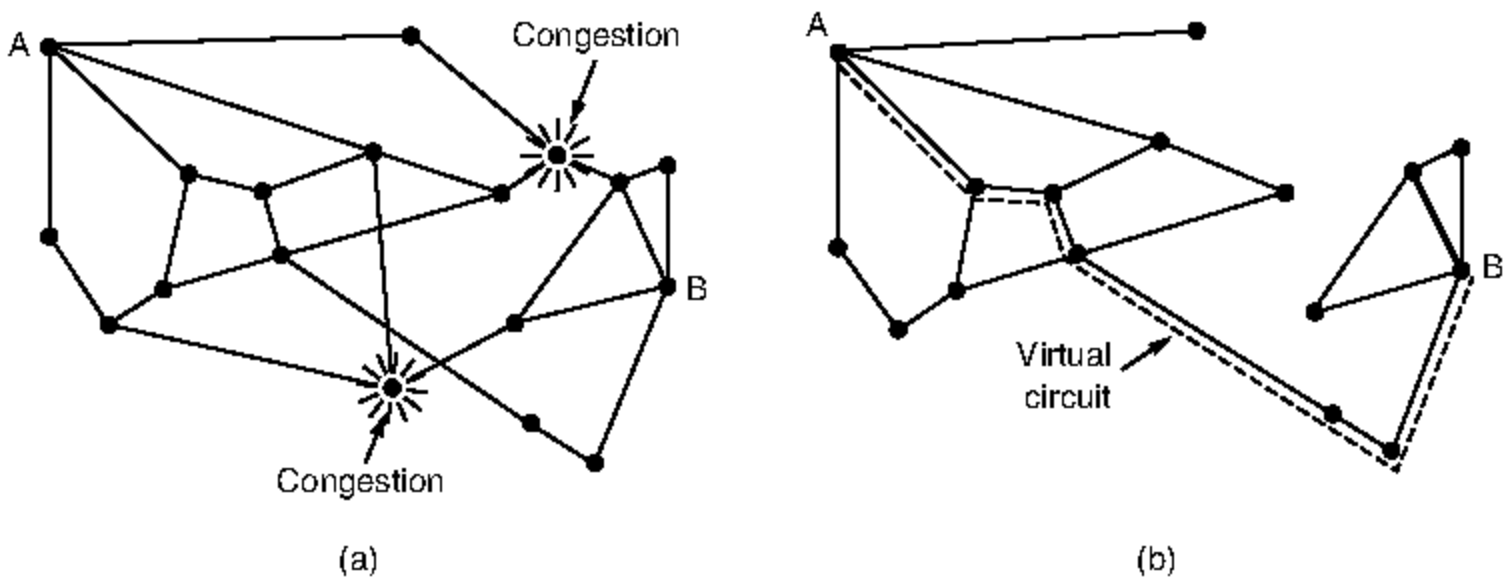


Figure 5-24. (a) A congested network. (b) The portion of the network that is not congested. A virtual circuit from *A* to *B* is also shown.

Suppose that a host attached to router *A* wants to set up a connection to a host attached to router *B*. Normally, this connection would pass through one of the congested routers. To avoid this situation, we can redraw the network as shown in Fig. 5-24(b), omitting the congested routers and all of their lines. The dashed line shows a possible route for the virtual circuit that avoids the congested routers. Shaikh et al. (1999) give a design for this kind of load-sensitive routing.

5.3.4 Traffic Throttling

In the Internet and many other computer networks, senders adjust their transmissions to send as much traffic as the network can readily deliver. In this setting, the network aims to operate just before the onset of congestion. When congestion is imminent, it must tell the senders to throttle back their transmissions and slow down. This feedback is business as usual rather than an exceptional situation. The term **congestion avoidance** is sometimes used to contrast this operating point with the one in which the network has become (overly) congested.

Let us now look at some approaches to throttling traffic that can be used in both datagram networks and virtual-circuit networks. Each approach must solve two problems. First, routers must determine when congestion is approaching, ideally before it has arrived. To do so, each router can continuously monitor the resources it is using. Three possibilities are the utilization of the output links, the buffering of queued packets inside the router, and the number of packets that are lost due to insufficient buffering. Of these possibilities, the second one is the most useful. Averages of utilization do not directly account for the burstiness of

most traffic—a utilization of 50% may be low for smooth traffic and too high for highly variable traffic. Counts of packet losses come too late. Congestion has already set in by the time that packets are lost.

The queueing delay inside routers directly captures any congestion experienced by packets. It should be low most of time, but will jump when there is a burst of traffic that generates a backlog. To maintain a good estimate of the queueing delay, d, a sample of the instantaneous queue length, s, can be made periodically and d updated according to

$$d_{\text{new}} = \alpha d_{\text{old}} + (1 - \alpha)s$$

where the constant α determines how fast the router forgets recent history. This is called an **EWMA** (**Exponentially Weighted Moving Average**). It smoothes out fluctuations and is equivalent to a low-pass filter. Whenever d moves above the threshold, the router notes the onset of congestion.

The second problem is that routers must deliver timely feedback to the senders that are causing the congestion. Congestion is experienced in the network, but relieving congestion requires action on behalf of the senders that are using the network. To deliver feedback, the router must identify the appropriate senders. It must then warn them carefully, without sending many more packets into the already congested network. Different schemes use different feedback mechanisms, as we will now describe.

Choke Packets

The most direct way to notify a sender of congestion is to tell it directly. In this approach, the router selects a congested packet and sends a **choke packet** back to the source host, giving it the destination found in the packet. The original packet may be tagged (a header bit is turned on) so that it will not generate any more choke packets farther along the path and then forwarded in the usual way. To avoid increasing load on the network during a time of congestion, the router may only send choke packets at a low rate.

When the source host gets the choke packet, it is required to reduce the traffic sent to the specified destination, for example, by 50%. In a datagram network, simply picking packets at random when there is congestion is likely to cause choke packets to be sent to fast senders, because they will have the most packets in the queue. The feedback implicit in this protocol can help prevent congestion yet not throttle any sender unless it causes trouble. For the same reason, it is likely that multiple choke packets will be sent to a given host and destination. The host should ignore these additional chokes for the fixed time interval until its reduction in traffic takes effect. After that period, further choke packets indicate that the network is still congested.

An example of a choke packet used in the early Internet is the SOURCE-QUENCH message (Postel, 1981). It never caught on, though, partly because the

circumstances in which it was generated and the effect it had were not clearly specified. The modern Internet uses an alternative notification design that we will describe next.

Explicit Congestion Notification

Instead of generating additional packets to warn of congestion, a router can tag any packet it forwards (by setting a bit in the packet's header) to signal that it is experiencing congestion. When the network delivers the packet, the destination can note that there is congestion and inform the sender when it sends a reply packet. The sender can then throttle its transmissions as before.

This design is called **ECN** (**Explicit Congestion Notification**) and is used in the Internet (Ramakrishnan et al., 2001). It is a refinement of early congestion signaling protocols, notably the binary feedback scheme of Ramakrishnan and Jain (1988) that was used in the DECNET architecture. Two bits in the IP packet header are used to record whether the packet has experienced congestion. Packets are unmarked when they are sent, as illustrated in Fig. 5-25. If any of the routers they pass through is congested, that router will then mark the packet as having experienced congestion as it is forwarded. The destination will then echo any marks back to the sender as an explicit congestion signal in its next reply packet. This is shown with a dashed line in the figure to indicate that it happens above the IP level (e.g., in TCP). The sender must then throttle its transmissions, as in the case of choke packets.

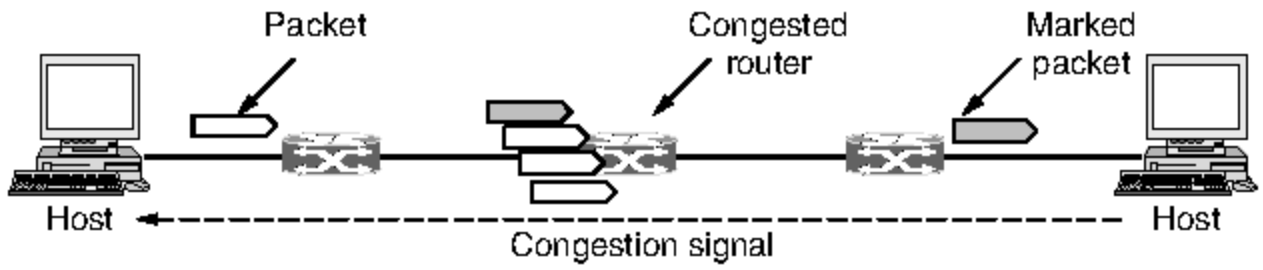


Figure 5-25. Explicit congestion notification

Hop-by-Hop Backpressure

At high speeds or over long distances, many new packets may be transmitted after congestion has been signaled because of the delay before the signal takes effect. Consider, for example, a host in San Francisco (router *A* in Fig. 5-26) that is sending traffic to a host in New York (router *D* in Fig. 5-26) at the OC-3 speed of 155 Mbps. If the New York host begins to run out of buffers, it will take about 40 msec for a choke packet to get back to San Francisco to tell it to slow down. An ECN indication will take even longer because it is delivered via the destination. Choke packet propagation is illustrated as the second, third, and fourth steps in

Fig. 5-26(a). In those 40 msec, another 6.2 megabits will have been sent. Even if the host in San Francisco completely shuts down immediately, the 6.2 megabits in the pipe will continue to pour in and have to be dealt with. Only in the seventh diagram in Fig. 5-26(a) will the New York router notice a slower flow.

An alternative approach is to have the choke packet take effect at every hop it passes through, as shown in the sequence of Fig. 5-26(b). Here, as soon as the choke packet reaches *F*, *F* is required to reduce the flow to *D*. Doing so will require *F* to devote more buffers to the connection, since the source is still sending away at full blast, but it gives *D* immediate relief, like a headache remedy in a television commercial. In the next step, the choke packet reaches *E*, which tells *E* to reduce the flow to *F*. This action puts a greater demand on *E*'s buffers but gives *F* immediate relief. Finally, the choke packet reaches *A* and the flow genuinely slows down.

The net effect of this hop-by-hop scheme is to provide quick relief at the point of congestion, at the price of using up more buffers upstream. In this way, congestion can be nipped in the bud without losing any packets. The idea is discussed in detail by Mishra et al. (1996).

5.3.5 Load Shedding

When none of the above methods make the congestion disappear, routers can bring out the heavy artillery: load shedding. **Load shedding** is a fancy way of saying that when routers are being inundated by packets that they cannot handle, they just throw them away. The term comes from the world of electrical power generation, where it refers to the practice of utilities intentionally blacking out certain areas to save the entire grid from collapsing on hot summer days when the demand for electricity greatly exceeds the supply.

The key question for a router drowning in packets is which packets to drop. The preferred choice may depend on the type of applications that use the network. For a file transfer, an old packet is worth more than a new one. This is because dropping packet 6 and keeping packets 7 through 10, for example, will only force the receiver to do more work to buffer data that it cannot yet use. In contrast, for real-time media, a new packet is worth more than an old one. This is because packets become useless if they are delayed and miss the time at which they must be played out to the user.

The former policy (old is better than new) is often called **wine** and the latter (new is better than old) is often called **milk** because most people would rather drink new milk and old wine than the alternative.

More intelligent load shedding requires cooperation from the senders. An example is packets that carry routing information. These packets are more important than regular data packets because they establish routes; if they are lost, the network may lose connectivity. Another example is that algorithms for compressing video, like MPEG, periodically transmit an entire frame and then send subsequent

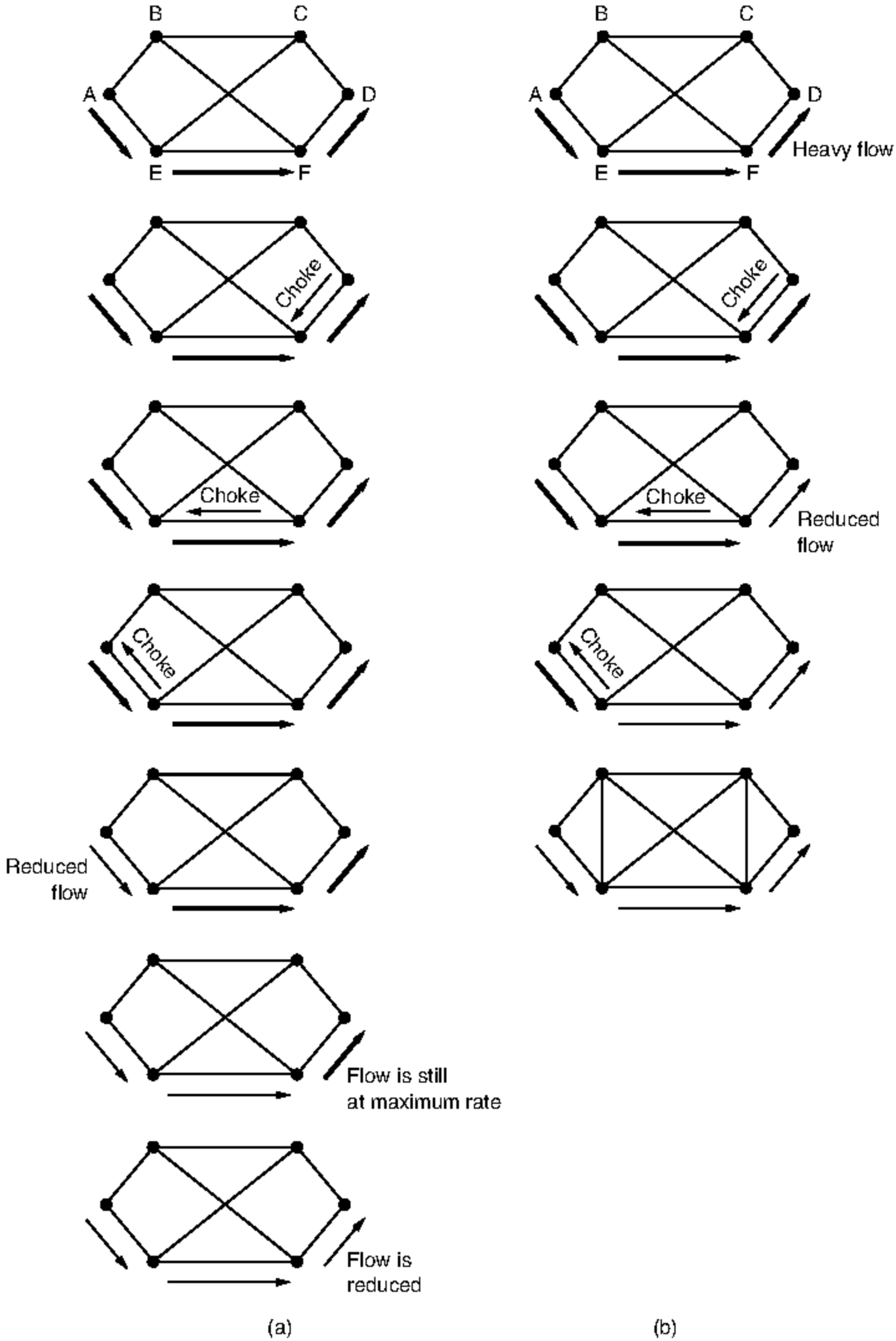


Figure 5-26. (a) A choke packet that affects only the source. (b) A choke packet that affects each hop it passes through.

frames as differences from the last full frame. In this case, dropping a packet that is part of a difference is preferable to dropping one that is part of a full frame because future packets depend on the full frame.

To implement an intelligent discard policy, applications must mark their packets to indicate to the network how important they are. Then, when packets have to be discarded, routers can first drop packets from the least important class, then the next most important class, and so on.

Of course, unless there is some significant incentive to avoid marking every packet as VERY IMPORTANT—NEVER, EVER DISCARD, nobody will do it. Often accounting and money are used to discourage frivolous marking. For example, the network might let senders send faster than the service they purchased allows if they mark excess packets as low priority. Such a strategy is actually not a bad idea because it makes more efficient use of idle resources, allowing hosts to use them as long as nobody else is interested, but without establishing a right to them when times get tough.

Random Early Detection

Dealing with congestion when it first starts is more effective than letting it gum up the works and then trying to deal with it. This observation leads to an interesting twist on load shedding, which is to discard packets before all the buffer space is really exhausted.

The motivation for this idea is that most Internet hosts do not yet get congestion signals from routers in the form of ECN. Instead, the only reliable indication of congestion that hosts get from the network is packet loss. After all, it is difficult to build a router that does not drop packets when it is overloaded. Transport protocols such as TCP are thus hardwired to react to loss as congestion, slowing down the source in response. The reasoning behind this logic is that TCP was designed for wired networks and wired networks are very reliable, so lost packets are mostly due to buffer overruns rather than transmission errors. Wireless links must recover transmission errors at the link layer (so they are not seen at the network layer) to work well with TCP.

This situation can be exploited to help reduce congestion. By having routers drop packets early, before the situation has become hopeless, there is time for the source to take action before it is too late. A popular algorithm for doing this is called **RED** (**Random Early Detection**) (Floyd and Jacobson, 1993). To determine when to start discarding, routers maintain a running average of their queue lengths. When the average queue length on some link exceeds a threshold, the link is said to be congested and a small fraction of the packets are dropped at random. Picking packets at random makes it more likely that the fastest senders will see a packet drop; this is the best option since the router cannot tell which source is causing the most trouble in a datagram network. The affected sender will notice the loss when there is no acknowledgement, and then the transport protocol

will slow down. The lost packet is thus delivering the same message as a choke packet, but implicitly, without the router sending any explicit signal.

RED routers improve performance compared to routers that drop packets only when their buffers are full, though they may require tuning to work well. For example, the ideal number of packets to drop depends on how many senders need to be notified of congestion. However, ECN is the preferred option if it is available. It works in exactly the same manner, but delivers a congestion signal explicitly rather than as a loss; RED is used when hosts cannot receive explicit signals.

5.4 QUALITY OF SERVICE

The techniques we looked at in the previous sections are designed to reduce congestion and improve network performance. However, there are applications (and customers) that demand stronger performance guarantees from the network than "the best that could be done under the circumstances.'' Multimedia applications in particular, often need a minimum throughput and maximum latency to work. In this section, we will continue our study of network performance, but now with a sharper focus on ways to provide quality of service that is matched to application needs. This is an area in which the Internet is undergoing a long-term upgrade.

An easy solution to provide good quality of service is to build a network with enough capacity for whatever traffic will be thrown at it. The name for this solution is **overprovisioning**. The resulting network will carry application traffic without significant loss and, assuming a decent routing scheme, will deliver packets with low latency. Performance doesn't get any better than this. To some extent, the telephone system is overprovisioned because it is rare to pick up a telephone and not get a dial tone instantly. There is simply so much capacity available that demand can almost always be met.

The trouble with this solution is that it is expensive. It is basically solving a problem by throwing money at it. Quality of service mechanisms let a network with less capacity meet application requirements just as well at a lower cost. Moreover, overprovisioning is based on expected traffic. All bets are off if the traffic pattern changes too much. With quality of service mechanisms, the network can honor the performance guarantees that it makes even when traffic spikes, at the cost of turning down some requests.

Four issues must be addressed to ensure quality of service:

1. What applications need from the network.
2. How to regulate the traffic that enters the network.
3. How to reserve resources at routers to guarantee performance.
4. Whether the network can safely accept more traffic.

No single technique deals efficiently with all these issues. Instead, a variety of techniques have been developed for use at the network (and transport) layer. Practical quality-of-service solutions combine multiple techniques. To this end, we will describe two versions of quality of service for the Internet called Integrated Services and Differentiated Services.

5.4.1 Application Requirements

A stream of packets from a source to a destination is called a **flow** (Clark, 1988). A flow might be all the packets of a connection in a connection-oriented network, or all the packets sent from one process to another process in a connectionless network. The needs of each flow can be characterized by four primary parameters: bandwidth, delay, jitter, and loss. Together, these determine the **QoS** (**Quality of Service**) the flow requires.

Several common applications and the stringency of their network requirements are listed in Fig. 5-27. Note that network requirements are less demanding than application requirements in those cases that the application can improve on the service provided by the network. In particular, networks do not need to be lossless for reliable file transfer, and they do not need to deliver packets with identical delays for audio and video playout. Some amount of loss can be repaired with retransmissions, and some amount of jitter can be smoothed by buffering packets at the receiver. However, there is nothing applications can do to remedy the situation if the network provides too little bandwidth or too much delay.

Application	Bandwidth	Delay	Jitter	Loss
Email	Low	Low	Low	Medium
File sharing	High	Low	Low	Medium
Web access	Medium	Medium	Low	Medium
Remote login	Low	Medium	Medium	Medium
Audio on demand	Low	Low	High	Low
Video on demand	High	Low	High	Low
Telephony	Low	High	High	Low
Videoconferencing	High	High	High	Low

Figure 5-27. Stringency of applications' quality-of-service requirements.

The applications differ in their bandwidth needs, with email, audio in all forms, and remote login not needing much, but file sharing and video in all forms needing a great deal.

More interesting are the delay requirements. File transfer applications, including email and video, are not delay sensitive. If all packets are delayed uniformly by a few seconds, no harm is done. Interactive applications, such as Web

surfing and remote login, are more delay sensitive. Real-time applications, such as telephony and videoconferencing, have strict delay requirements. If all the words in a telephone call are each delayed by too long, the users will find the connection unacceptable. On the other hand, playing audio or video files from a server does not require low delay.

The variation (i.e., standard deviation) in the delay or packet arrival times is called **jitter**. The first three applications in Fig. 5-27 are not sensitive to the packets arriving with irregular time intervals between them. Remote login is somewhat sensitive to that, since updates on the screen will appear in little bursts if the connection suffers much jitter. Video and especially audio are extremely sensitive to jitter. If a user is watching a video over the network and the frames are all delayed by exactly 2.000 seconds, no harm is done. But if the transmission time varies randomly between 1 and 2 seconds, the result will be terrible unless the application hides the jitter. For audio, a jitter of even a few milliseconds is clearly audible.

The first four applications have more stringent requirements on loss than audio and video because all bits must be delivered correctly. This goal is usually achieved with retransmissions of packets that are lost in the network by the transport layer. This is wasted work; it would be better if the network refused packets it was likely to lose in the first place. Audio and video applications can tolerate some lost packets without retransmission because people do not notice short pauses or occasional skipped frames.

To accommodate a variety of applications, networks may support different categories of QoS. An influential example comes from ATM networks, which were once part of a grand vision for networking but have since become a niche technology. They support:

1. Constant bit rate (e.g., telephony).
2. Real-time variable bit rate (e.g., compressed videoconferencing).
3. Non-real-time variable bit rate (e.g., watching a movie on demand).
4. Available bit rate (e.g., file transfer).

These categories are also useful for other purposes and other networks. Constant bit rate is an attempt to simulate a wire by providing a uniform bandwidth and a uniform delay. Variable bit rate occurs when video is compressed, with some frames compressing more than others. Sending a frame with a lot of detail in it may require sending many bits, whereas a shot of a white wall may compress extremely well. Movies on demand are not actually real time because a few seconds of video can easily be buffered at the receiver before playback starts, so jitter on the network merely causes the amount of stored-but-not-played video to vary. Available bit rate is for applications such as email that are not sensitive to delay or jitter and will take what bandwidth they can get.

5.4.2 Traffic Shaping

Before the network can make QoS guarantees, it must know what traffic is being guaranteed. In the telephone network, this characterization is simple. For example, a voice call (in uncompressed format) needs 64 kbps and consists of one 8-bit sample every 125 μsec. However, traffic in data networks is **bursty**. It typically arrives at nonuniform rates as the traffic rate varies (e.g., videoconferencing with compression), users interact with applications (e.g., browsing a new Web page), and computers switch between tasks. Bursts of traffic are more difficult to handle than constant-rate traffic because they can fill buffers and cause packets to be lost.

Traffic shaping is a technique for regulating the average rate and burstiness of a flow of data that enters the network. The goal is to allow applications to transmit a wide variety of traffic that suits their needs, including some bursts, yet have a simple and useful way to describe the possible traffic patterns to the network. When a flow is set up, the user and the network (i.e., the customer and the provider) agree on a certain traffic pattern (i.e., shape) for that flow. In effect, the customer says to the provider "My transmission pattern will look like this; can you handle it?"

Sometimes this agreement is called an **SLA** (**Service Level Agreement**), especially when it is made over aggregate flows and long periods of time, such as all of the traffic for a given customer. As long as the customer fulfills her part of the bargain and only sends packets according to the agreed-on contract, the provider promises to deliver them all in a timely fashion.

Traffic shaping reduces congestion and thus helps the network live up to its promise. However, to make it work, there is also the issue of how the provider can tell if the customer is following the agreement and what to do if the customer is not. Packets in excess of the agreed pattern might be dropped by the network, or they might be marked as having lower priority. Monitoring a traffic flow is called **traffic policing**.

Shaping and policing are not so important for peer-to-peer and other transfers that will consume any and all available bandwidth, but they are of great importance for real-time data, such as audio and video connections, which have stringent quality-of-service requirements.

Leaky and Token Buckets

We have already seen one way to limit the amount of data an application sends: the sliding window, which uses one parameter to limit how much data is in transit at any given time, which indirectly limits the rate. Now we will look at a more general way to characterize traffic, with the leaky bucket and token bucket algorithms. The formulations are slightly different but give an equivalent result.

Try to imagine a bucket with a small hole in the bottom, as illustrated in Fig. 5-28(b). No matter the rate at which water enters the bucket, the outflow is at a constant rate, R, when there is any water in the bucket and zero when the bucket is empty. Also, once the bucket is full to capacity B, any additional water entering it spills over the sides and is lost.

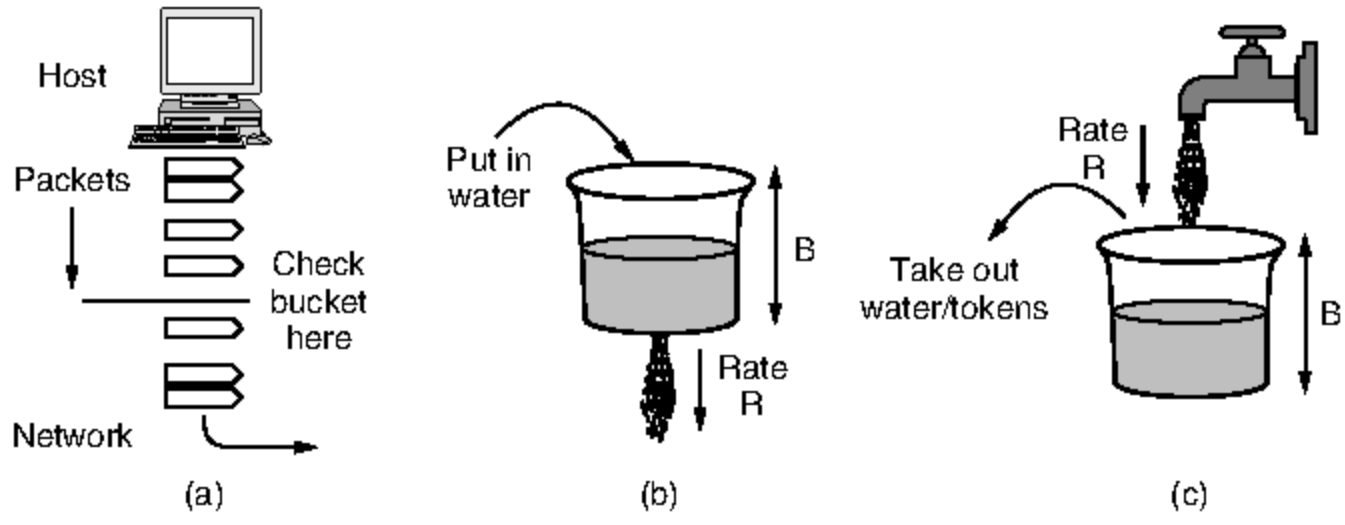


Figure 5-28. (a) Shaping packets. (b) A leaky bucket. (c) A token bucket.

This bucket can be used to shape or police packets entering the network, as shown in Fig. 5-28(a). Conceptually, each host is connected to the network by an interface containing a leaky bucket. To send a packet into the network, it must be possible to put more water into the bucket. If a packet arrives when the bucket is full, the packet must either be queued until enough water leaks out to hold it or be discarded. The former might happen at a host shaping its traffic for the network as part of the operating system. The latter might happen in hardware at a provider network interface that is policing traffic entering the network. This technique was proposed by Turner (1986) and is called the **leaky bucket algorithm**.

A different but equivalent formulation is to imagine the network interface as a bucket that is being filled, as shown in Fig. 5-28(c). The tap is running at rate R and the bucket has a capacity of B, as before. Now, to send a packet we must be able to take water, or tokens, as the contents are commonly called, out of the bucket (rather than putting water into the bucket). No more than a fixed number of tokens, B, can accumulate in the bucket, and if the bucket is empty, we must wait until more tokens arrive before we can send another packet. This algorithm is called the **token bucket algorithm**.

Leaky and token buckets limit the long-term rate of a flow but allow short-term bursts up to a maximum regulated length to pass through unaltered and without suffering any artificial delays. Large bursts will be smoothed by a leaky bucket traffic shaper to reduce congestion in the network. As an example, imagine that a computer can produce data at up to 1000 Mbps (125 million bytes/sec) and that the first link of the network also runs at this speed. The pattern of traffic the host generates is shown in Fig. 5-29(a). This pattern is bursty. The average

rate over one second is 200 Mbps, even though the host sends a burst of 16,000 KB at the top speed of 1000 Mbps (for 1/8 of the second).

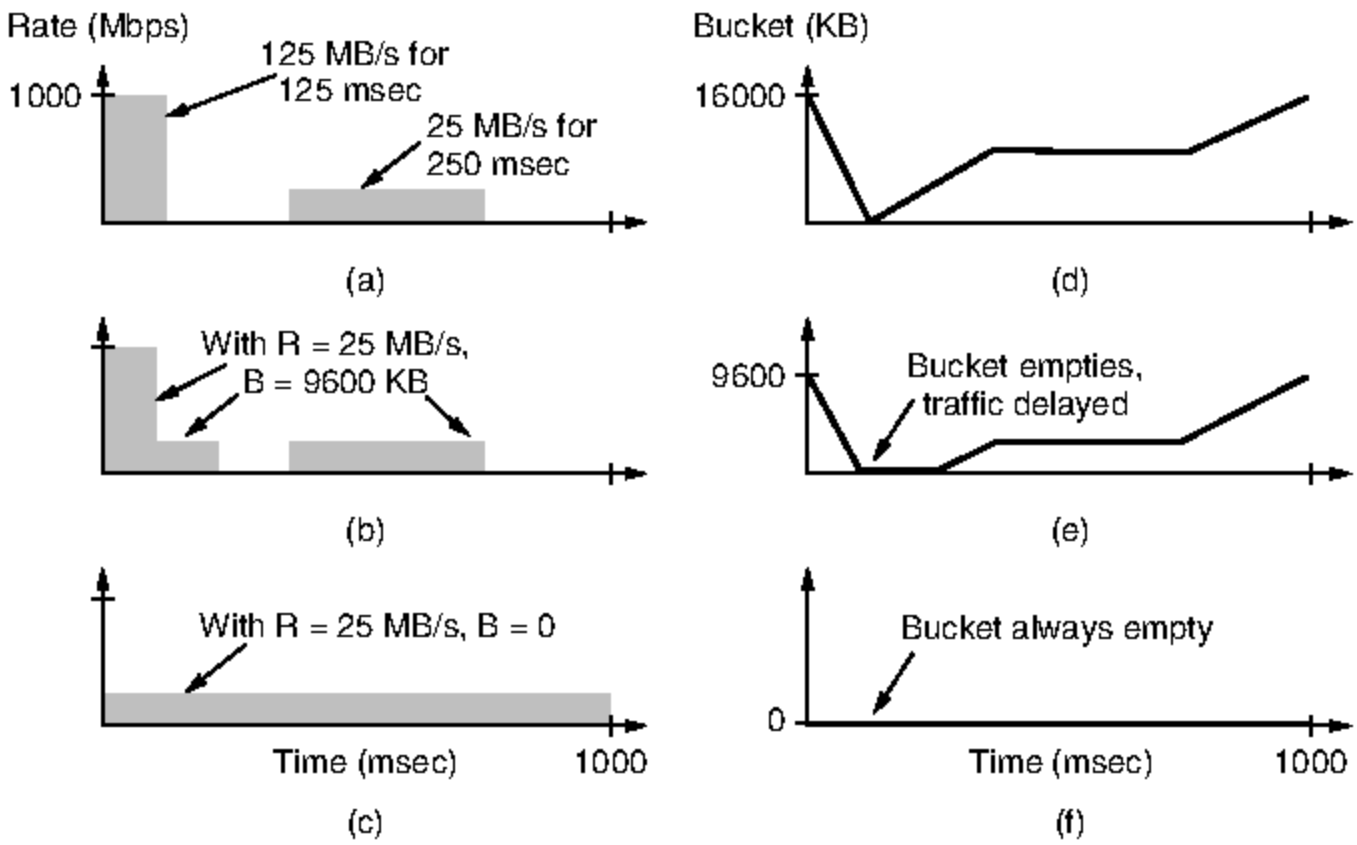


Figure 5-29. (a) Traffic from a host. Output shaped by a token bucket of rate 200 Mbps and capacity (b) 9600 KB and (c) 0 KB. Token bucket level for shaping with rate 200 Mbps and capacity (d) 16,000 KB, (e) 9600 KB, and (f) 0 KB.

Now suppose that the routers can accept data at the top speed only for short intervals, until their buffers fill up. The buffer size is 9600 KB, smaller than the traffic burst. For long intervals, the routers work best at rates not exceeding 200 Mbps (say, because this is all the bandwidth given to the customer). The implication is that if traffic is sent in this pattern, some of it will be dropped in the network because it does not fit into the buffers at routers.

To avoid this packet loss, we can shape the traffic at the host with a token bucket. If we use a rate, R, of 200 Mbps and a capacity, B, of 9600 KB, the traffic will fall within what the network can handle. The output of this token bucket is shown in Fig. 5-29(b). The host can send full throttle at 1000 Mbps for a short while until it has drained the bucket. Then it has to cut back to 200 Mbps until the burst has been sent. The effect is to spread out the burst over time because it was too large to handle all at once. The level of the token bucket is shown in Fig. 5-29(e). It starts off full and is depleted by the initial burst. When it reaches zero, new packets can be sent only at the rate at which the buffer is filling; there can be no more bursts until the bucket has recovered. The bucket fills when no traffic is being sent and stays flat when traffic is being sent at the fill rate.

We can also shape the traffic to be less bursty. Fig. 5-29(c) shows the output of a token bucket with $R = 200$ Mbps and a capacity of 0. This is the extreme case

in which the traffic has been completely smoothed. No bursts are allowed, and the traffic enters the network at a steady rate. The corresponding bucket level, shown in Fig. 5-29(f), is always empty. Traffic is being queued on the host for release into the network and there is always a packet waiting to be sent when it is allowed.

Finally, Fig. 5-29(d) shows the bucket level for a token bucket with $R = 200$ Mbps and a capacity of $B = 16{,}000$ KB. This is the smallest token bucket through which the traffic passes unaltered. It might be used at a router in the network to police the traffic that the host sends. If the host is sending traffic that conforms to the token bucket on which it has agreed with the network, the traffic will fit through that same token bucket run at the router at the edge of the network. If the host sends at a faster or burstier rate, the token bucket will run out of water. If this happens, a traffic policer will know that the traffic is not as described. It will then either drop the excess packets or lower their priority, depending on the design of the network. In our example, the bucket empties only momentarily, at the end of the initial burst, then recovers enough for the next burst.

Leaky and token buckets are easy to implement. We will now describe the operation of a token bucket. Even though we have described water flowing continuously into and out of the bucket, real implementations must work with discrete quantities. A token bucket is implemented with a counter for the level of the bucket. The counter is advanced by $R/\Delta T$ units at every clock tick of ΔT seconds. This would be 200 Kbit every 1 msec in our example above. Every time a unit of traffic is sent into the network, the counter is decremented, and traffic may be sent until the counter reaches zero.

When the packets are all the same size, the bucket level can just be counted in packets (e.g., 200 Mbit is 20 packets of 1250 bytes). However, often variable-sized packets are being used. In this case, the bucket level is counted in bytes. If the residual byte count is too low to send a large packet, the packet must wait until the next tick (or even longer, if the fill rate is small).

Calculating the length of the maximum burst (until the bucket empties) is slightly tricky. It is longer than just 9600 KB divided by 125 MB/sec because while the burst is being output, more tokens arrive. If we call the burst length S sec., the maximum output rate M bytes/sec, the token bucket capacity B bytes, and the token arrival rate R bytes/sec, we can see that an output burst contains a maximum of $B + RS$ bytes. We also know that the number of bytes in a maximum-speed burst of length S seconds is MS. Hence, we have

$$B + RS = MS$$

We can solve this equation to get $S = B/(M - R)$. For our parameters of $B = 9600$ KB, $M = 125$ MB/sec, and $R = 25$ MB/sec, we get a burst time of about 94 msec.

A potential problem with the token bucket algorithm is that it reduces large bursts down to the long-term rate R. It is frequently desirable to reduce the peak rate, but without going down to the long-term rate (and also without raising the

long-term rate to allow more traffic into the network). One way to get smoother traffic is to insert a second token bucket after the first one. The rate of the second bucket should be much higher than the first one. Basically, the first bucket characterizes the traffic, fixing its average rate but allowing some bursts. The second bucket reduces the peak rate at which the bursts are sent into the network. For example, if the rate of the second token bucket is set to be 500 Mbps and the capacity is set to 0, the initial burst will enter the network at a peak rate of 500 Mbps, which is lower than the 1000 Mbps rate we had previously.

Using all of these buckets can be a bit tricky. When token buckets are used for traffic shaping at hosts, packets are queued and delayed until the buckets permit them to be sent. When token buckets are used for traffic policing at routers in the network, the algorithm is simulated to make sure that no more packets are sent than permitted. Nevertheless, these tools provide ways to shape the network traffic into more manageable forms to assist in meeting quality-of-service requirements.

5.4.3 Packet Scheduling

Being able to regulate the shape of the offered traffic is a good start. However, to provide a performance guarantee, we must reserve sufficient resources along the route that the packets take through the network. To do this, we are assuming that the packets of a flow follow the same route. Spraying them over routers at random makes it hard to guarantee anything. As a consequence, something similar to a virtual circuit has to be set up from the source to the destination, and all the packets that belong to the flow must follow this route.

Algorithms that allocate router resources among the packets of a flow and between competing flows are called **packet scheduling algorithms**. Three different kinds of resources can potentially be reserved for different flows:

1. Bandwidth.
2. Buffer space.
3. CPU cycles.

The first one, bandwidth, is the most obvious. If a flow requires 1 Mbps and the outgoing line has a capacity of 2 Mbps, trying to direct three flows through that line is not going to work. Thus, reserving bandwidth means not oversubscribing any output line.

A second resource that is often in short supply is buffer space. When a packet arrives, it is buffered inside the router until it can be transmitted on the chosen outgoing line. The purpose of the buffer is to absorb small bursts of traffic as the flows contend with each other. If no buffer is available, the packet has to be discarded since there is no place to put it. For good quality of service, some buffers might be reserved for a specific flow so that flow does not have to compete for

buffers with other flows. Up to some maximum value, there will always be a buffer available when the flow needs one.

Finally, CPU cycles may also be a scarce resource. It takes router CPU time to process a packet, so a router can process only a certain number of packets per second. While modern routers are able to process most packets quickly, some kinds of packets require greater CPU processing, such as the ICMP packets we will describe in Sec. 5.6. Making sure that the CPU is not overloaded is needed to ensure timely processing of these packets.

Packet scheduling algorithms allocate bandwidth and other router resources by determining which of the buffered packets to send on the output line next. We already described the most straightforward scheduler when explaining how routers work. Each router buffers packets in a queue for each output line until they can be sent, and they are sent in the same order that they arrived. This algorithm is known as **FIFO** (**First-In First-Out**), or equivalently **FCFS** (**First-Come First-Serve**).

FIFO routers usually drop newly arriving packets when the queue is full. Since the newly arrived packet would have been placed at the end of the queue, this behavior is called **tail drop**. It is intuitive, and you may be wondering what alternatives exist. In fact, the RED algorithm we described in Sec. 5.3.5 chose a newly arriving packet to drop at random when the average queue length grew large. The other scheduling algorithms that we will describe also create other opportunities for deciding which packet to drop when the buffers are full.

FIFO scheduling is simple to implement, but it is not suited to providing good quality of service because when there are multiple flows, one flow can easily affect the performance of the other flows. If the first flow is aggressive and sends large bursts of packets, they will lodge in the queue. Processing packets in the order of their arrival means that the aggressive sender can hog most of the capacity of the routers its packets traverse, starving the other flows and reducing their quality of service. To add insult to injury, the packets of the other flows that do get through are likely to be delayed because they had to sit in the queue behind many packets from the aggressive sender.

Many packet scheduling algorithms have been devised that provide stronger isolation between flows and thwart attempts at interference (Bhatti and Crowcroft, 2000). One of the first ones was the **fair queueing** algorithm devised by Nagle (1987). The essence of this algorithm is that routers have separate queues, one for each flow for a given output line. When the line becomes idle, the router scans the queues round-robin, as shown in Fig. 5-30. It then takes the first packet on the next queue. In this way, with n hosts competing for the output line, each host gets to send one out of every n packets. It is fair in the sense that all flows get to send packets at the same rate. Sending more packets will not improve this rate.

Although a start, the algorithm has a flaw: it gives more bandwidth to hosts that use large packets than to hosts that use small packets. Demers et al. (1990) suggested an improvement in which the round-robin is done in such a way as to

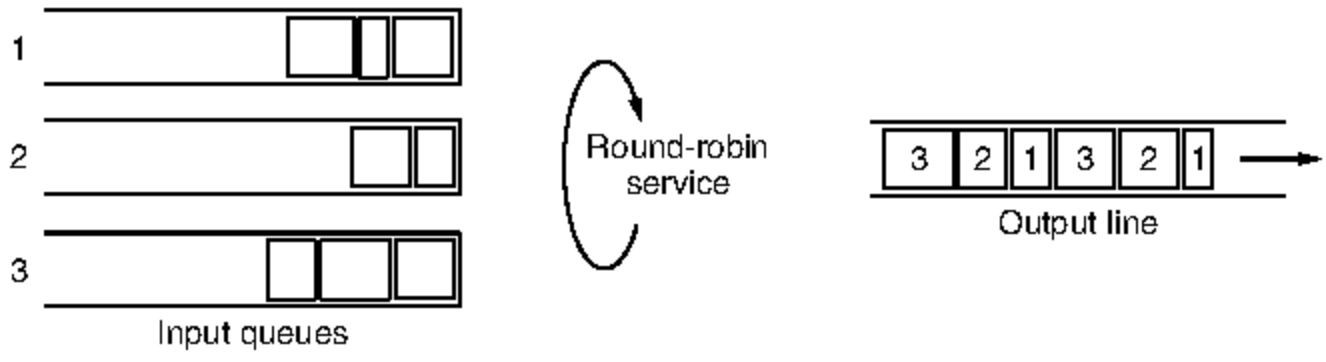
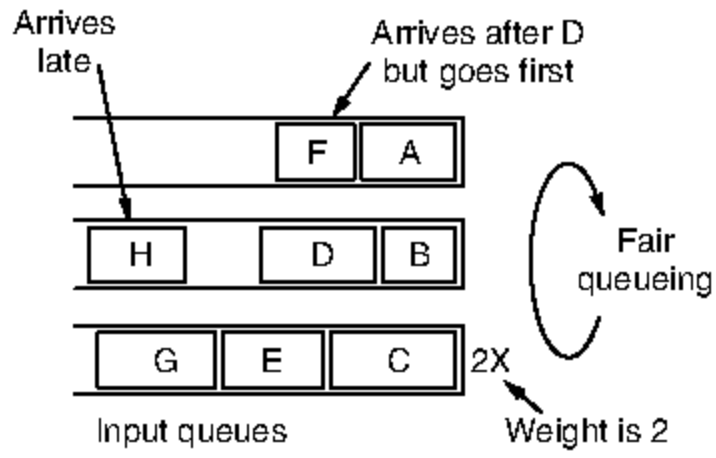


Figure 5-30. Round-robin fair queueing.

simulate a byte-by-byte round-robin, instead of a packet-by-packet round-robin. The trick is to compute a virtual time that is the number of the round at which each packet would finish being sent. Each round drains a byte from all of the queues that have data to send. The packets are then sorted in order of their finishing times and sent in that order.

This algorithm and an example of finish times for packets arriving in three flows are illustrated in Fig. 5-31. If a packet has length L, the round at which it will finish is simply L rounds after the start time. The start time is either the finish time of the previous packet, or the arrival time of the packet, if the queue is empty when it arrives.



Packet	Arrival time	Length	Finish time	Output order
A	0	8	8	1
B	5	6	11	3
C	5	10	10	2
D	8	9	20	7
E	8	8	14	4
F	10	6	16	5
G	11	10	19	6
H	20	8	28	8

(b)

Figure 5-31. (a) Weighted Fair Queueing. (b) Finishing times for the packets.

From the table in Fig. 5-32(b), and looking only at the first two packets in the top two queues, packets arrive in the order A, B, D, and F. Packet A arrives at round 0 and is 8 bytes long, so its finish time is round 8. Similarly the finish time for packet B is 11. Packet D arrives while B is being sent. Its finish time is 9 byte-rounds after it starts when B finishes, or 20. Similarly, the finish time for F is 16. In the absence of new arrivals, the relative sending order is A, B, F, D, even though F arrived after D. It is possible that another small packet will arrive on the top flow and obtain a finish time before D. It will only jump ahead of D if the

transmission of that packet has not started. Fair queueing does not preempt packets that are currently being transmitted. Because packets are sent in their entirety, fair queueing is only an approximation of the ideal byte-by-byte scheme. But it is a very good approximation, staying within one packet transmission of the ideal scheme at all times.

One shortcoming of this algorithm in practice is that it gives all hosts the same priority. In many situations, it is desirable to give, for example, video servers more bandwidth than, say, file servers. This is easily possible by giving the video server two or more bytes per round. This modified algorithm is called **WFQ (Weighted Fair Queueing)**. Letting the number of bytes per round be the weight of a flow, W, we can now give the formula for computing the finish time:

$$F_i = \max(A_i, F_{i-1}) + L_i/W$$

where A_i is the arrival time, F_i is the finish time, and L_i is the length of packet i. The bottom queue of Fig. 5-31(a) has a weight of 2, so its packets are sent more quickly as you can see in the finish times given in Fig. 5-31(b).

Another practical consideration is implementation complexity. WFQ requires that packets be inserted by their finish time into a sorted queue. With N flows, this is at best an $O(\log N)$ operation per packet, which is difficult to achieve for many flows in high-speed routers. Shreedhar and Varghese (1995) describe an approximation called **deficit round robin** that can be implemented very efficiently, with only $O(1)$ operations per packet. WFQ is widely used given this approximation.

Other kinds of scheduling algorithms exist, too. A simple example is priority scheduling, in which each packet is marked with a priority. High-priority packets are always sent before any low-priority packets that are buffered. Within a priority, packets are sent in FIFO order. However, priority scheduling has the disadvantage that a burst of high-priority packets can starve low-priority packets, which may have to wait indefinitely. WFQ often provides a better alternative. By giving the high-priority queue a large weight, say 3, high-priority packets will often go through a short line (as relatively few packets should be high priority) yet some fraction of low priority packets will continue to be sent even when there is high priority traffic. A high and low priority system is essentially a two-queue WFQ system in which the high priority has infinite weight.

As a final example of a scheduler, packets might carry timestamps and be sent in timestamp order. Clark et al. (1992) describe a design in which the timestamp records how far the packet is behind or ahead of schedule as it is sent through a sequence of routers on the path. Packets that have been queued behind other packets at a router will tend to be behind schedule, and the packets that have been serviced first will tend to be ahead of schedule. Sending packets in order of their timestamps has the beneficial effect of speeding up slow packets while at the same time slowing down fast packets. The result is that all packets are delivered by the network with a more consistent delay.

5.4.4 Admission Control

We have now seen all the necessary elements for QoS and it is time to put them together to actually provide it. QoS guarantees are established through the process of admission control. We first saw admission control used to control congestion, which is a performance guarantee, albeit a weak one. The guarantees we are considering now are stronger, but the model is the same. The user offers a flow with an accompanying QoS requirement to the network. The network then decides whether to accept or reject the flow based on its capacity and the commitments it has made to other flows. If it accepts, the network reserves capacity in advance at routers to guarantee QoS when traffic is sent on the new flow.

The reservations must be made at all of the routers along the route that the packets take through the network. Any routers on the path without reservations might become congested, and a single congested router can break the QoS guarantee. Many routing algorithms find the single best path between each source and each destination and send all traffic over the best path. This may cause some flows to be rejected if there is not enough spare capacity along the best path. QoS guarantees for new flows may still be accommodated by choosing a different route for the flow that has excess capacity. This is called **QoS routing**. Chen and Nahrstedt (1998) give an overview of these techniques. It is also possible to split the traffic for each destination over multiple paths to more easily find excess capacity. A simple method is for routers to choose equal-cost paths and to divide the traffic equally or in proportion to the capacity of the outgoing links. However, more sophisticated algorithms are also available (Nelakuditi and Zhang, 2002).

Given a path, the decision to accept or reject a flow is not a simple matter of comparing the resources (bandwidth, buffers, cycles) requested by the flow with the router's excess capacity in those three dimensions. It is a little more complicated than that. To start with, although some applications may know about their bandwidth requirements, few know about buffers or CPU cycles, so at the minimum, a different way is needed to describe flows and translate this description to router resources. We will get to this shortly.

Next, some applications are far more tolerant of an occasional missed deadline than others. The applications must choose from the type of guarantees that the network can make, whether hard guarantees or behavior that will hold most of the time. All else being equal, everyone would like hard guarantees, but the difficulty is that they are expensive because they constrain worst case behavior. Guarantees for most of the packets are often sufficient for applications, and more flows with this guarantee can be supported for a fixed capacity.

Finally, some applications may be willing to haggle about the flow parameters and others may not. For example, a movie viewer that normally runs at 30 frames/sec may be willing to drop back to 25 frames/sec if there is not enough free bandwidth to support 30 frames/sec. Similarly, the number of pixels per frame, audio bandwidth, and other properties may be adjustable.

Because many parties may be involved in the flow negotiation (the sender, the receiver, and all the routers along the path between them), flows must be described accurately in terms of specific parameters that can be negotiated. A set of such parameters is called a **flow specification**. Typically, the sender (e.g., the video server) produces a flow specification proposing the parameters it would like to use. As the specification propagates along the route, each router examines it and modifies the parameters as need be. The modifications can only reduce the flow, not increase it (e.g., a lower data rate, not a higher one). When it gets to the other end, the parameters can be established.

As an example of what can be in a flow specification, consider the example of Fig. 5-32. This is based on RFCs 2210 and 2211 for Integrated Services, a QoS design we will cover in the next section. It has five parameters. The first two parameters, the *token bucket rate* and *token bucket size*, use a token bucket to give the maximum sustained rate the sender may transmit, averaged over a long time interval, and the largest burst it can send over a short time interval.

Parameter	Unit
Token bucket rate	Bytes/sec
Token bucket size	Bytes
Peak data rate	Bytes/sec
Minimum packet size	Bytes
Maximum packet size	Bytes

Figure 5-32. An example flow specification.

The third parameter, the *peak data rate*, is the maximum transmission rate tolerated, even for brief time intervals. The sender must never exceed this rate even for short bursts.

The last two parameters specify the minimum and maximum packet sizes, including the transport and network layer headers (e.g., TCP and IP). The minimum size is useful because processing each packet takes some fixed time, no matter how short. A router may be prepared to handle 10,000 packets/sec of 1 KB each, but not be prepared to handle 100,000 packets/sec of 50 bytes each, even though this represents a lower data rate. The maximum packet size is important due to internal network limitations that may not be exceeded. For example, if part of the path goes over an Ethernet, the maximum packet size will be restricted to no more than 1500 bytes no matter what the rest of the network can handle.

An interesting question is how a router turns a flow specification into a set of specific resource reservations. At first glance, it might appear that if a router has a link that runs at, say, 1 Gbps and the average packet is 1000 bits, it can process 1 million packets/sec. This observation is not the case, though, because there will always be idle periods on the link due to statistical fluctuations in the load. If the

link needs every bit of capacity to get its work done, idling for even a few bits creates a backlog it can never get rid of.

Even with a load slightly below the theoretical capacity, queues can build up and delays can occur. Consider a situation in which packets arrive at random with a mean arrival rate of λ packets/sec. The packets have random lengths and can be sent on the link with a mean service rate of μ packets/sec. Under the assumption that both the arrival and service distributions are Poisson distributions (what is called an M/M/1 queueing system, where "M" stands for Markov, i.e., Poisson), it can be proven using queueing theory that the mean delay experienced by a packet, T, is

$$T = \frac{1}{\mu} \times \frac{1}{1 - \lambda/\mu} = \frac{1}{\mu} \times \frac{1}{1 - \rho}$$

where $\rho = \lambda/\mu$ is the CPU utilization. The first factor, $1/\mu$, is what the service time would be in the absence of competition. The second factor is the slowdown due to competition with other flows. For example, if $\lambda = 950{,}000$ packets/sec and $\mu = 1{,}000{,}000$ packets/sec, then $\rho = 0.95$ and the mean delay experienced by each packet will be 20 μsec instead of 1 μsec. This time accounts for both the queueing time and the service time, as can be seen when the load is very low ($\lambda/\mu \approx 0$). If there are, say, 30 routers along the flow's route, queueing delay alone will account for 600 μsec of delay.

One method of relating flow specifications to router resources that correspond to bandwidth and delay performance guarantees is given by Parekh and Gallagher (1993, 1994). It is based on traffic sources shaped by (R, B) token buckets and WFQ at routers. Each flow is given a WFQ weight W large enough to drain its token bucket rate R as shown in Fig. 5-33. For example, if the flow has a rate of 1 Mbps and the router and output link have a capacity of 1 Gbps, the weight for the flow must be greater than 1/1000th of the total of the weights for all of the flows at that router for the output link. This guarantees the flow a minimum bandwidth. If it cannot be given a large enough rate, the flow cannot be admitted.

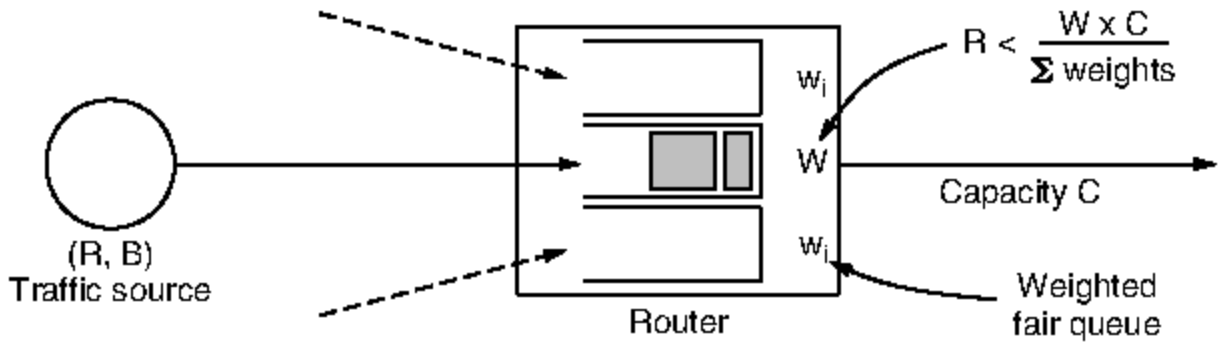


Figure 5-33. Bandwidth and delay guarantees with token buckets and WFQ.

The largest queueing delay the flow will see is a function of the burst size of the token bucket. Consider the two extreme cases. If the traffic is smooth, without

any bursts, packets will be drained from the router just as quickly as they arrive. There will be no queueing delay (ignoring packetization effects). On the other hand, if the traffic is saved up in bursts, then a maximum-size burst, B, may arrive at the router all at once. In this case the maximum queueing delay, D, will be the time taken to drain this burst at the guaranteed bandwidth, or B/R (again, ignoring packetization effects). If this delay is too large, the flow must request more bandwidth from the network.

These guarantees are hard. The token buckets bound the burstiness of the source, and fair queueing isolates the bandwidth given to different flows. This means that the flow will meet its bandwidth and delay guarantees regardless of how the other competing flows behave at the router. Those other flows cannot break the guarantee even by saving up traffic and all sending at once.

Moreover, the result holds for a path through multiple routers in any network topology. Each flow gets a minimum bandwidth because that bandwidth is guaranteed at each router. The reason each flow gets a maximum delay is more subtle. In the worst case that a burst of traffic hits the first router and competes with the traffic of other flows, it will be delayed up to the maximum delay of D. However, this delay will also smooth the burst. In turn, this means that the burst will incur no further queueing delays at later routers. The overall queueing delay will be at most D.

5.4.5 Integrated Services

Between 1995 and 1997, IETF put a lot of effort into devising an architecture for streaming multimedia. This work resulted in over two dozen RFCs, starting with RFCs 2205–2212. The generic name for this work is **integrated services**. It was aimed at both unicast and multicast applications. An example of the former is a single user streaming a video clip from a news site. An example of the latter is a collection of digital television stations broadcasting their programs as streams of IP packets to many receivers at various locations. Below we will concentrate on multicast, since unicast is a special case of multicast.

In many multicast applications, groups can change membership dynamically, for example, as people enter a video conference and then get bored and switch to a soap opera or the croquet channel. Under these conditions, the approach of having the senders reserve bandwidth in advance does not work well, since it would require each sender to track all entries and exits of its audience. For a system designed to transmit television with millions of subscribers, it would not work at all.

RSVP—The Resource reSerVation Protocol

The main part of the integrated services architecture that is visible to the users of the network is **RSVP**. It is described in RFCs 2205–2210. This protocol is used for making the reservations; other protocols are used for sending the data.

RSVP allows multiple senders to transmit to multiple groups of receivers, permits individual receivers to switch channels freely, and optimizes bandwidth use while at the same time eliminating congestion.

In its simplest form, the protocol uses multicast routing using spanning trees, as discussed earlier. Each group is assigned a group address. To send to a group, a sender puts the group's address in its packets. The standard multicast routing algorithm then builds a spanning tree covering all group members. The routing algorithm is not part of RSVP. The only difference from normal multicasting is a little extra information that is multicast to the group periodically to tell the routers along the tree to maintain certain data structures in their memories.

As an example, consider the network of Fig. 5-34(a). Hosts 1 and 2 are multicast senders, and hosts 3, 4, and 5 are multicast receivers. In this example, the senders and receivers are disjoint, but in general, the two sets may overlap. The multicast trees for hosts 1 and 2 are shown in Fig. 5-34(b) and Fig. 5-34(c), respectively.

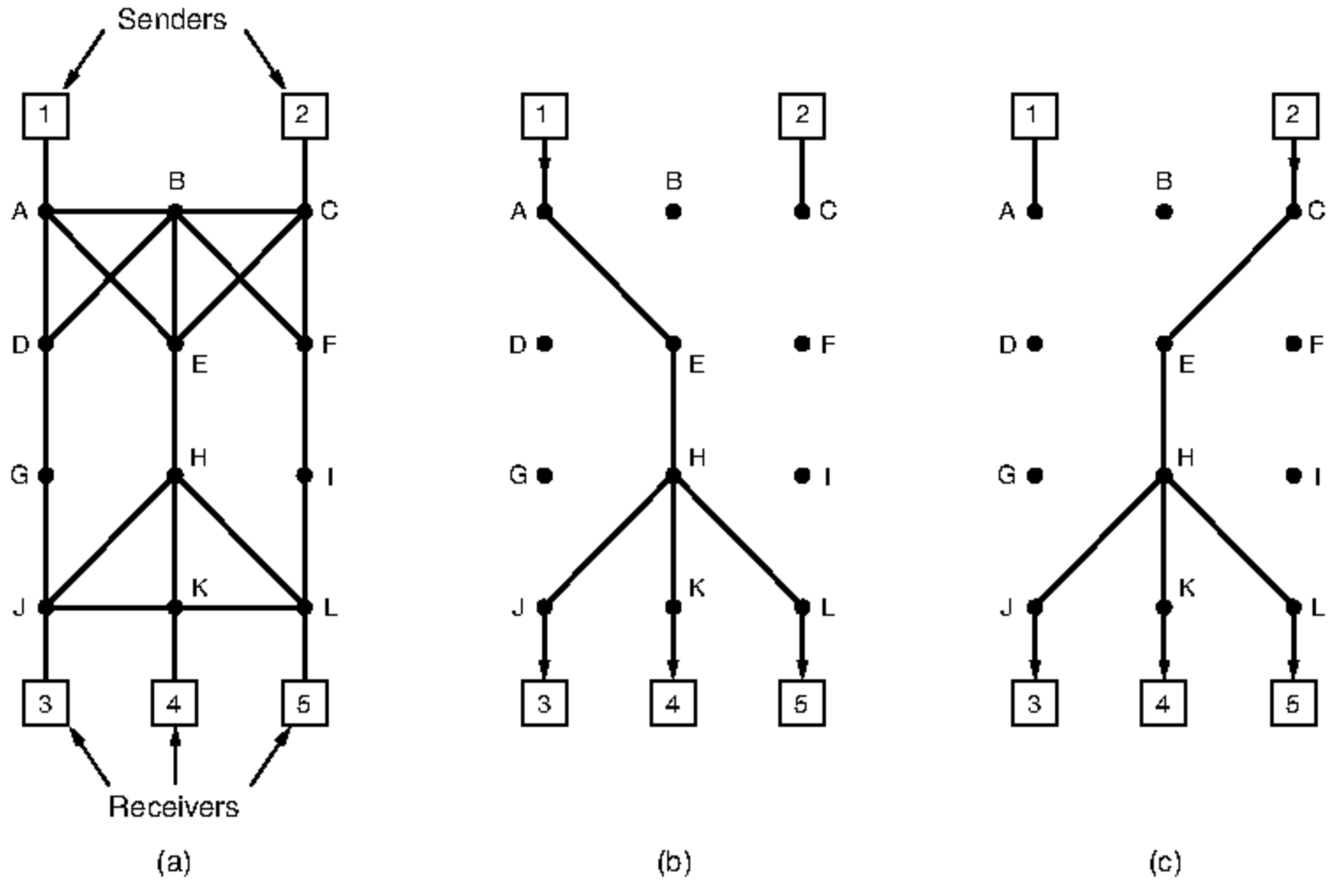


Figure 5-34. (a) A network. (b) The multicast spanning tree for host 1. (c) The multicast spanning tree for host 2.

To get better reception and eliminate congestion, any of the receivers in a group can send a reservation message up the tree to the sender. The message is propagated using the reverse path forwarding algorithm discussed earlier. At each

hop, the router notes the reservation and reserves the necessary bandwidth. We saw in the previous section how a weighted fair queueing scheduler can be used to make this reservation. If insufficient bandwidth is available, it reports back failure. By the time the message gets back to the source, bandwidth has been reserved all the way from the sender to the receiver making the reservation request along the spanning tree.

An example of such a reservation is shown in Fig. 5-35(a). Here host 3 has requested a channel to host 1. Once it has been established, packets can flow from 1 to 3 without congestion. Now consider what happens if host 3 next reserves a channel to the other sender, host 2, so the user can watch two television programs at once. A second path is reserved, as illustrated in Fig. 5-35(b). Note that two separate channels are needed from host 3 to router *E* because two independent streams are being transmitted.

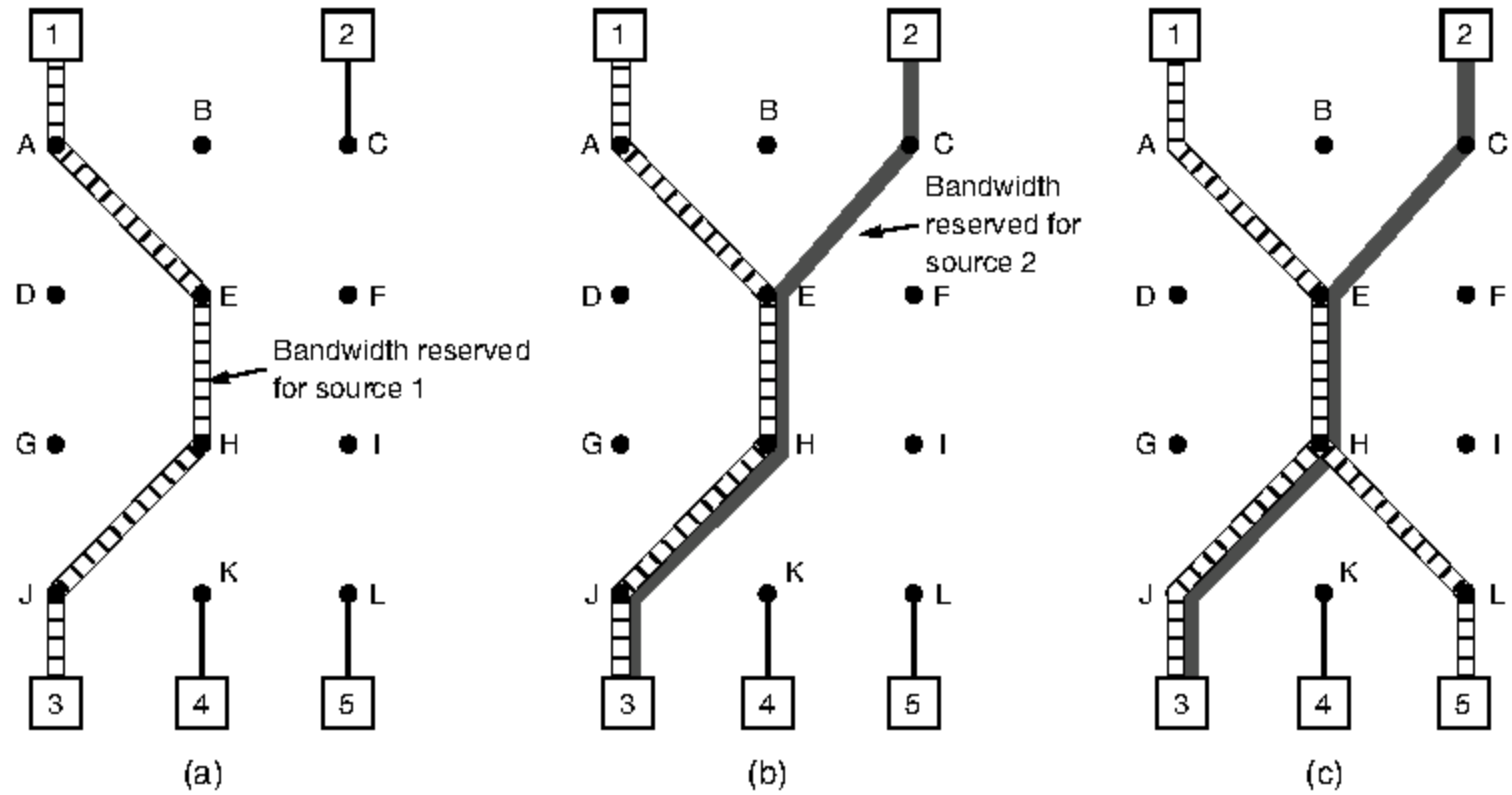


Figure 5-35. (a) Host 3 requests a channel to host 1. (b) Host 3 then requests a second channel, to host 2. (c) Host 5 requests a channel to host 1.

Finally, in Fig. 5-35(c), host 5 decides to watch the program being transmitted by host 1 and also makes a reservation. First, dedicated bandwidth is reserved as far as router *H*. However, this router sees that it already has a feed from host 1, so if the necessary bandwidth has already been reserved, it does not have to reserve any more. Note that hosts 3 and 5 might have asked for different amounts of bandwidth (e.g., if host 3 is playing on a small screen and only wants the low-resolution information), so the capacity reserved must be large enough to satisfy the greediest receiver.

When making a reservation, a receiver can (optionally) specify one or more sources that it wants to receive from. It can also specify whether these choices

are fixed for the duration of the reservation or whether the receiver wants to keep open the option of changing sources later. The routers use this information to optimize bandwidth planning. In particular, two receivers are only set up to share a path if they both agree not to change sources later on.

The reason for this strategy in the fully dynamic case is that reserved bandwidth is decoupled from the choice of source. Once a receiver has reserved bandwidth, it can switch to another source and keep that portion of the existing path that is valid for the new source. If host 2 is transmitting several video streams in real time, for example a TV broadcaster with multiple channels, host 3 may switch between them at will without changing its reservation: the routers do not care what program the receiver is watching.

5.4.6 Differentiated Services

Flow-based algorithms have the potential to offer good quality of service to one or more flows because they reserve whatever resources are needed along the route. However, they also have a downside. They require an advance setup to establish each flow, something that does not scale well when there are thousands or millions of flows. Also, they maintain internal per-flow state in the routers, making them vulnerable to router crashes. Finally, the changes required to the router code are substantial and involve complex router-to-router exchanges for setting up the flows. As a consequence, while work continues to advance integrated services, few deployments of it or anything like it exist yet.

For these reasons, IETF has also devised a simpler approach to quality of service, one that can be largely implemented locally in each router without advance setup and without having the whole path involved. This approach is known as **class-based** (as opposed to flow-based) quality of service. IETF has standardized an architecture for it, called **differentiated services**, which is described in RFCs 2474, 2475, and numerous others. We will now describe it.

Differentiated services can be offered by a set of routers forming an administrative domain (e.g., an ISP or a telco). The administration defines a set of service classes with corresponding forwarding rules. If a customer subscribes to differentiated services, customer packets entering the domain are marked with the class to which they belong. This information is carried in the *Differentiated services* field of IPv4 and IPv6 packets (described in Sec. 5.6). The classes are defined as **per hop behaviors** because they correspond to the treatment the packet will receive at each router, not a guarantee across the network. Better service is provided to packets with some per-hop behaviors (e.g., premium service) than to others (e.g., regular service). Traffic within a class may be required to conform to some specific shape, such as a leaky bucket with some specified drain rate. An operator with a nose for business might charge extra for each premium packet transported or might allow up to N premium packets per month for a fixed additional monthly fee. Note that this scheme requires no advance setup, no resource

reservation, and no time-consuming end-to-end negotiation for each flow, as with integrated services. This makes differentiated services relatively easy to implement.

Class-based service also occurs in other industries. For example, package delivery companies often offer overnight, two-day, and three-day service. Airlines offer first class, business class, and cattle-class service. Long-distance trains often have multiple service classes. Even the Paris subway has two different service classes. For packets, the classes may differ in terms of delay, jitter, and probability of being discarded in the event of congestion, among other possibilities (but probably not roomier Ethernet frames).

To make the difference between flow-based quality of service and class-based quality of service clearer, consider an example: Internet telephony. With a flow-based scheme, each telephone call gets its own resources and guarantees. With a class-based scheme, all the telephone calls together get the resources reserved for the class telephony. These resources cannot be taken away by packets from the Web browsing class or other classes, but no telephone call gets any private resources reserved for it alone.

Expedited Forwarding

The choice of service classes is up to each operator, but since packets are often forwarded between networks run by different operators, IETF has defined some network-independent service classes. The simplest class is **expedited forwarding**, so let us start with that one. It is described in RFC 3246.

The idea behind expedited forwarding is very simple. Two classes of service are available: regular and expedited. The vast majority of the traffic is expected to be regular, but a limited fraction of the packets are expedited. The expedited packets should be able to transit the network as though no other packets were present. In this way they will get low loss, low delay and low jitter service—just what is needed for VoIP. A symbolic representation of this "two-tube" system is given in Fig. 5-36. Note that there is still just one physical line. The two logical pipes shown in the figure represent a way to reserve bandwidth for different classes of service, not a second physical line.

One way to implement this strategy is as follows. Packets are classified as expedited or regular and marked accordingly. This step might be done on the sending host or in the ingress (first) router. The advantage of doing classification on the sending host is that more information is available about which packets belong to which flows. This task may be performed by networking software or even the operating system, to avoid having to change existing applications. For example, it is becoming common for VoIP packets to be marked for expedited service by hosts. If the packets pass through a corporate network or ISP that supports expedited service, they will receive preferential treatment. If the network does not support expedited service, no harm is done.

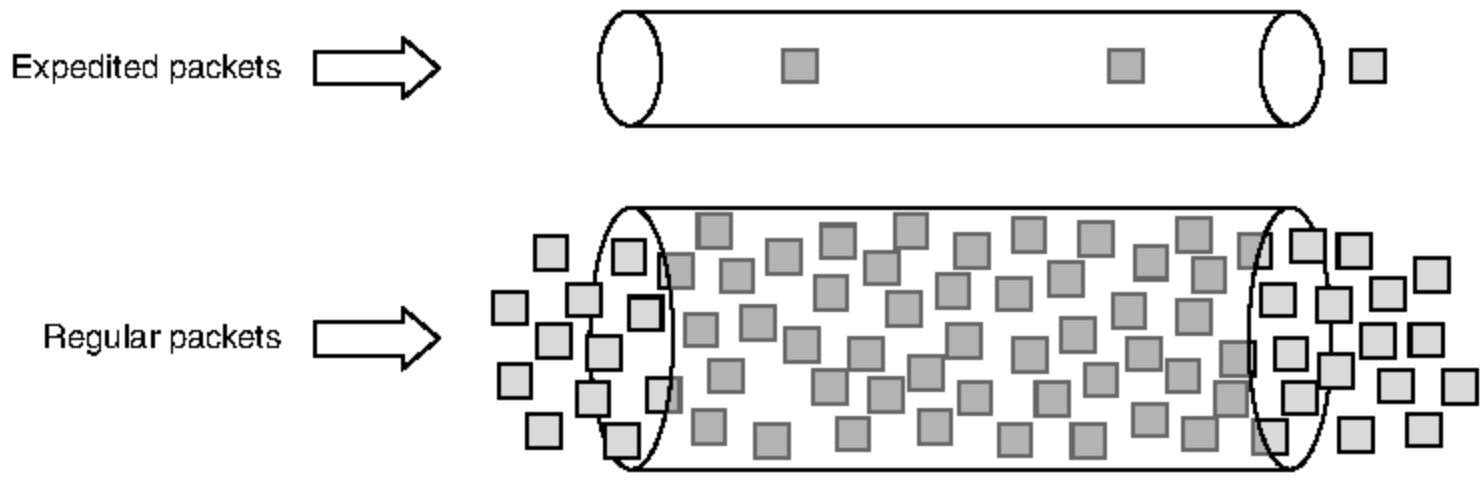


Figure 5-36. Expedited packets experience a traffic-free network.

Of course, if the marking is done by the host, the ingress router is likely to police the traffic to make sure that customers are not sending more expedited traffic than they have paid for. Within the network, the routers may have two output queues for each outgoing line, one for expedited packets and one for regular packets. When a packet arrives, it is queued accordingly. The expedited queue is given priority over the regular one, for example, by using a priority scheduler. In this way, expedited packets see an unloaded network, even when there is, in fact, a heavy load of regular traffic.

Assured Forwarding

A somewhat more elaborate scheme for managing the service classes is called **assured forwarding**. It is described in RFC 2597. Assured forwarding specifies that there shall be four priority classes, each class having its own resources. The top three classes might be called gold, silver, and bronze. In addition, it defines three discard classes for packets that are experiencing congestion: low, medium, and high. Taken together, these two factors define 12 service classes.

Figure 5-37 shows one way packets might be processed under assured forwarding. The first step is to classify the packets into one of the four priority classes. As before, this step might be done on the sending host (as shown in the figure) or in the ingress router, and the rate of higher-priority packets may be limited by the operator as part of the service offering.

The next step is to determine the discard class for each packet. This is done by passing the packets of each priority class through a traffic policer such as a token bucket. The policer lets all of the traffic through, but it identifies packets that fit within small bursts as low discard, packets that exceed small bursts as medium discard, and packets that exceed large bursts as high discard. The combination of priority and discard class is then encoded in each packet.

Finally, the packets are processed by routers in the network with a packet scheduler that distinguishes the different classes. A common choice is to use

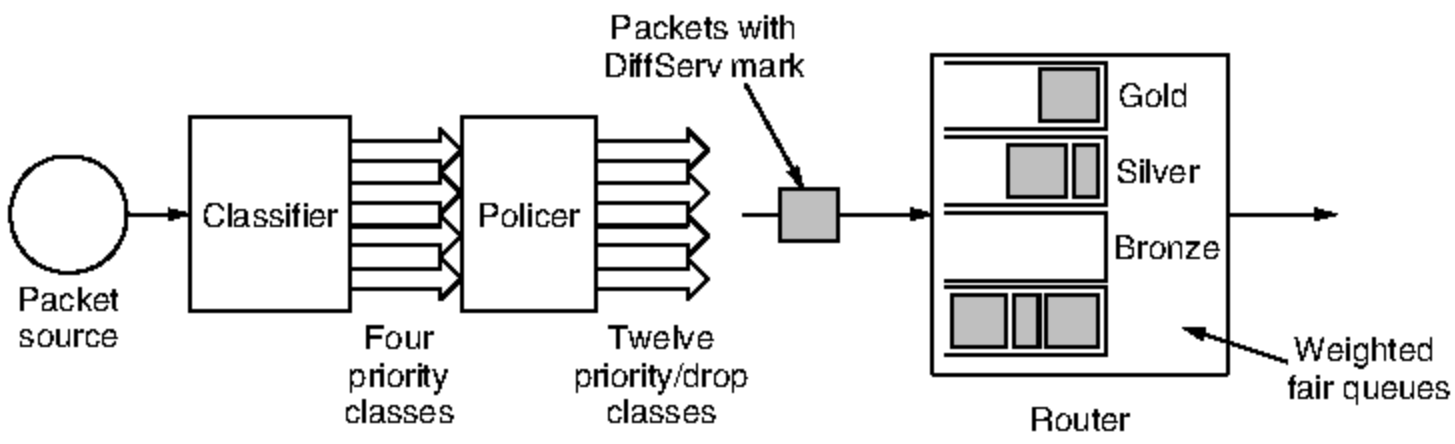


Figure 5-37. A possible implementation of assured forwarding.

weighted fair queueing for the four priority classes, with higher classes given higher weights. In this way, the higher classes will get most of the bandwidth, but the lower classes will not be starved of bandwidth entirely. For example, if the weights double from one class to the next higher class, the higher class will get twice the bandwidth. Within a priority class, packets with a higher discard class can be preferentially dropped by running an algorithm such as RED (Random Early Detection), which we saw in Sec. 5.3.5. RED will start to drop packets as congestion builds but before the router has run out of buffer space. At this stage, there is still buffer space with which to accept low discard packets while dropping high discard packets.

5.5 INTERNETWORKING

Until now, we have implicitly assumed that there is a single homogeneous network, with each machine using the same protocol in each layer. Unfortunately, this assumption is wildly optimistic. Many different networks exist, including PANs, LANs, MANs, and WANs. We have described Ethernet, Internet over cable, the fixed and mobile telephone networks, 802.11, 802.16, and more. Numerous protocols are in widespread use across these networks in every layer. In the following sections, we will take a careful look at the issues that arise when two or more networks are connected to form an **internetwork**, or more simply an **internet**.

It would be much simpler to join networks together if everyone used a single networking technology, and it is often the case that there is a dominant kind of network, such as Ethernet. Some pundits speculate that the multiplicity of technologies will go away as soon as everyone realizes how wonderful [fill in your favorite network] is. Do not count on it. History shows this to be wishful thinking. Different kinds of networks grapple with different problems, so, for example, Ethernet and satellite networks are always likely to differ. Reusing existing systems, such as running data networks on top of cable, the telephone network, and power

lines, adds constraints that cause the features of the networks to diverge. Heterogeneity is here to stay.

If there will always be different networks, it would be simpler if we did not need to interconnect them. This also is unlikely. Bob Metcalfe postulated that the value of a network with N nodes is the number of connections that may be made between the nodes, or N^2 (Gilder, 1993). This means that large networks are much more valuable than small networks because they allow many more connections, so there always will be an incentive to combine smaller networks.

The Internet is the prime example of this interconnection. (We will write Internet with a capital "I" to distinguish it from other internets, or connected networks.) The purpose of joining all these networks is to allow users on any of them to communicate with users on all the other ones. When you pay an ISP for Internet service, you may be charged depending on the bandwidth of your line, but what you are really paying for is the ability to exchange packets with any other host that is also connected to the Internet. After all, the Internet would not be very popular if you could only send packets to other hosts in the same city.

Since networks often differ in important ways, getting packets from one network to another is not always so easy. We must address problems of heterogeneity, and also problems of scale as the resulting internet grows very large. We will begin by looking at how networks can differ to see what we are up against. Then we shall see the approach used so successfully by IP (Internet Protocol), the network layer protocol of the Internet, including techniques for tunneling through networks, routing in internetworks, and packet fragmentation.

5.5.1 How Networks Differ

Networks can differ in many ways. Some of the differences, such as different modulation techniques or frame formats, are internal to the physical and data link layers. These differences will not concern us here. Instead, in Fig. 5-38 we list some of the differences that can be exposed to the network layer. It is papering over these differences that makes internetworking more difficult than operating within a single network.

When packets sent by a source on one network must transit one or more foreign networks before reaching the destination network, many problems can occur at the interfaces between networks. To start with, the source needs to be able to address the destination. What do we do if the source is on an Ethernet network and the destination is on a WiMAX network? Assuming we can even specify a WiMAX destination from an Ethernet network, packets would cross from a connectionless network to a connection-oriented one. This may require that a new connection be set up on short notice, which injects a delay, and much overhead if the connection is not used for many more packets.

Many specific differences may have to be accommodated as well. How do we multicast a packet to a group with some members on a network that does not

Item	Some Possibilities
Service offered	Connectionless versus connection oriented
Addressing	Different sizes, flat or hierarchical
Broadcasting	Present or absent (also multicast)
Packet size	Every network has its own maximum
Ordering	Ordered and unordered delivery
Quality of service	Present or absent; many different kinds
Reliability	Different levels of loss
Security	Privacy rules, encryption, etc.
Parameters	Different timeouts, flow specifications, etc.
Accounting	By connect time, packet, byte, or not at all

Figure 5-38. Some of the many ways networks can differ.

support multicast? The differing max packet sizes used by different networks can be a major nuisance, too. How do you pass an 8000-byte packet through a network whose maximum size is 1500 bytes? If packets on a connection-oriented network transit a connectionless network, they may arrive in a different order than they were sent. That is something the sender likely did not expect, and it might come as an (unpleasant) surprise to the receiver as well.

These kinds of differences can be papered over, with some effort. For example, a gateway joining two networks might generate separate packets for each destination in lieu of better network support for multicasting. A large packet might be broken up, sent in pieces, and then joined back together. Receivers might buffer packets and deliver them in order.

Networks also can differ in large respects that are more difficult to reconcile. The clearest example is quality of service. If one network has strong QoS and the other offers best effort service, it will be impossible to make bandwidth and delay guarantees for real-time traffic end to end. In fact, they can likely only be made while the best-effort network is operated at a low utilization, or hardly used, which is unlikely to be the goal of most ISPs. Security mechanisms are problematic, but at least encryption for confidentiality and data integrity can be layered on top of networks that do not already include it. Finally, differences in accounting can lead to unwelcome bills when normal usage suddenly becomes expensive, as roaming mobile phone users with data plans have discovered.

5.5.2 How Networks Can Be Connected

There are two basic choices for connecting different networks: we can build devices that translate or convert packets from each kind of network into packets for each other network, or, like good computer scientists, we can try to solve the