Tight Bounds for *k*-Set Agreement

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Tight Bounds for k-Set Agreement

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Abstract

We prove tight bounds on the time needed to solve *k-set agreement*. In this problem, each processor starts with an arbitrary input value taken from a fixed set, and halts after choosing an output value. In every execution, at most k distinct output values may be chosen, and every processor's output value must be some processor's input value. We analyze this problem in a synchronous, message-passing model where processors fail by crashing. We prove a lower bound of $\lfloor f/k \rfloor + 1$ rounds of communication for solutions to k-set agreement that tolerate f failures, and we exhibit a protocol proving the matching upper bound. This result shows that there is an inherent tradeoff between the running time, the degree of coordination required, and the number of faults tolerated, even in idealized models like the synchronous model. The proof of this result is interesting because it is the first to apply topological techniques to the synchronous model.

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1 Introduction

Most interesting problems in concurrent and distributed computing require processors to coordinate their actions in some way. It can also be important for protocols solving these problems to tolerate processor failures, and to execute quickly. Ideally, one would like to optimize all three properties—degree of coordination, fault-tolerance, and efficiency—but in practice, of course, it is usually necessary to make tradeoffs among them. In this paper, we give a precise characterization of the tradeoffs required by studying a family of basic coordination problems called k-set agreement.

In *k-set agreement* [Cha91], each processor starts with an arbitrary input value and halts after choosing an output value. These output values must satisfy two conditions: each output value must be some processor's input value, and the set of output values chosen must contain at most k distinct values. The first condition rules out trivial solutions in which a single value is hard-wired into the protocol and chosen by all processors in all executions, and the second condition requires that the processors coordinate their choices to some degree. This problem is interesting because it defines a family of coordination problems of increasing difficulty. At one extreme, if n is the number of processors in the system, then n-set agreement is trivial: each processor simply chooses its own input value. At the other extreme, 1-set agreement requires that all processors choose the same output value, a problem equivalent to the *consensus* problem [LSP82, PSL80, FL82, FLP85, Dol82, Fis83]. Consensus is well-known to be the "hardest" problem, in the sense that all other decision problems can be reduced to it. Consensus arises in applications as diverse as on-board aircraft control [$\mathrm{W}^{+}78$], database transaction commit [BHG87], and concurrent object design [Her91]. Between these extremes, as we vary the value of k from n to 1, we gradually increase the degree of processor coordination required.

We consider this family of problems in a *synchronous*, *message-passing* model with *crash failures*. In this model, n processors communicate by sending messages over a completely connected network. Computation in this model proceeds in a sequence of rounds. In each round, processors send messages to other processors, then receive messages sent to them in the same round, and then perform some local computation and change state. This means that all processors take steps at the same rate, and that all messages take the same amount of time to be delivered. Communication is reliable, but up to f processors can fail by stopping in the middle of the protocol.

The primary contribution of this paper is a lower bound on the amount of time required to solve k-set agreement, together with a protocol for k-set agreement that proves a matching upper bound. Specifically, we prove that any protocol solving k-set agreement in this model and tolerating f failures requires $\lfloor f/k \rfloor + 1$ rounds of communication in the worst case—assuming $n \geq f + k + 1$, meaning that there are at least $k+1$ nonfaulty processors—and we prove a matching upper bound by exhibiting a protocol that solves k-set agreement in $\lfloor f/k \rfloor + 1$ rounds. Since consensus is just 1-set agreement, our lower bound implies the well-known lower bound of $f + 1$ rounds for consensus when $n \geq f + 2$ [FL82]. More important, the running time $r = \lfloor f/k \rfloor + 1$ demonstrates that there is a smooth but inescapable tradeoff among the number f of faults tolerated, the degree k of coordination achieved, and the time r the protocol must run. For a fixed value of f, Figure 1 shows that 2-set agreement can be achieved

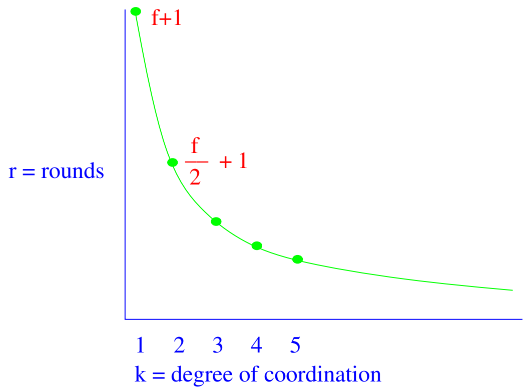

Figure 1: Tradeoff between rounds and degree of coordination.

in half the time needed to achieve consensus. In addition, the lower bound proof itself is interesting because of the geometric proof technique we use, combining ideas due to Chaudhuri [Cha91, Cha93], Fischer and Lynch [FL82], Herlihy and Shavit [HS93], and Dwork, Moses, and Tuttle [DM90, MT88].

In the past few years, researchers have developed powerful new tools based on classical algebraic topology for analyzing tasks in asynchronous models (e.g., [AR96, BG93, GK96, HR94, HR95, HS93, HS94, SZ93]).

The principal innovation of these papers is to model computations as simplicial complexes (rather than graphs) and to derive connections between computations and the topological properties of their complexes. This paper extends this topological approach in several new ways: it is the first to derive results in the synchronous model, it derives lower bounds rather than computability results, and it uses explicit constructions instead of existential arguments.

Although the synchronous model makes some strong (and possibly unrealistic) assumptions, it is well-suited for proving lower bounds. The synchronous model is a special case of almost every realistic model of a concurrent system we can imagine, and therefore any lower bound for k-set agreement in this simple model translates into a lower bound in any more complex model. For example, our lower bound holds for models that permit messages to be lost, failed processors to restart, or processor speeds to vary. Moreover, our techniques may be helpful in understanding how to prove (possibly) stricter lower bounds in more complex models. Naturally, our protocol for k-set agreement in the synchronous model does not work in more general models, but it is still useful because it shows that our lower bound is the best possible in the synchronous model.

This paper is organized as follows. In Section 2, we give an informal overview of our lower bound proof. In Section 3 we define our model of computation, and in Section 4 we define k-set agreement. In Sections 5 through 9 we prove our lower bound, and in Section 10 we give a protocol solving k-set agreement, proving the matching upper bound.

2 Overview

We start with an informal overview of the ideas used in the lower bound proof. For the remainder of this paper, suppose P is a protocol that solves k-set agreement and tolerates the failure of f out of n processors, and suppose P halts in $r < \lfloor f/k \rfloor + 1$ rounds. This means that all nonfaulty processors have chosen an output value at time r in every execution of P. In addition, suppose $n \geq f + k + 1$, which means that at least $k + 1$ processors never fail. Our goal is to consider the *global states* that occur at time r in executions of P, and to show that in one of these states there are $k + 1$ processors that have chosen $k + 1$ distinct values, violating k-set agreement. Our strategy is to consider the *local states* of processors that occur at time r in executions of P, and to investigate the combinations of these local states that occur in global states. This investigation depends on the construction of a geometric object. In this section, we use a simplified version of this object to illustrate the general ideas in our proof.

Since consensus is a special case of k-set agreement, it is helpful to review the standard proof of the $f + 1$ round lower bound for consensus [FL82, DS83, Mer85, DM90] to see why new ideas are needed for k-set agreement. Suppose that the protocol P is a consensus protocol, which means that in all executions of P all nonfaulty processors have chosen the same output value at time r. Two global states g_1 and g_2 at time r are said to be *similar* if some nonfaulty processor p has the same local state in both global states. The crucial property of similarity is that the decision value of any processor in one global state completely determines the decision value for any processor in all similar global states. For example, if all processors decide v in g_1, then certainly p decides v in g_1. Since p has the same local state in g_1 and g_2, and since p's decision value is a function of its local state, processor p also decides v in g_2. Since all processors agree with p in g_2, all processors decide v in g_2, and it follows that the decision value in g_1 determines the decision value in g_2. A *similarity chain* is a sequence of global states, $g_1, \dots, g_\ell$, such that g_i is similar to g_{i+1}. A simple inductive argument shows that the decision value in g_1 determines the decision value in g_ℓ. The lower bound proof consists of showing that all time r global states of P lie on a single similarity chain. It follows that all processors choose the same value in all executions of P, independent of the input values, violating the definition of consensus.

The problem with k-set agreement is that the decision values in one global state do not determine the decision values in similar global states. If p has the same local state in g_1 and g_2, then p must choose the same value in both states, but the values chosen by the other processors are not determined. Even if $n - 1$ processors have the same local state in g_1 and g_2, the decision value of the last processor is still not determined. The fundamental insight in this paper is that k-set agreement requires considering all "degrees" of similarity at once, focusing on the number and identity of local states common to two global states. While this seems difficult—if not impossible—to do using conventional graph theoretic techniques like similarity chains, there is a *geometric* generalization of similarity chains that provides a compact way of capturing all degrees of similarity simultaneously, and it is the basis of our proof.

A simplex is just the natural generalization of a triangle to n dimensions: for example, a 0-dimensional simplex is a vertex, a 1-dimensional simplex is an edge linking two vertices, a 2-dimensional simplex is a solid triangle, and a 3-dimensional simplex

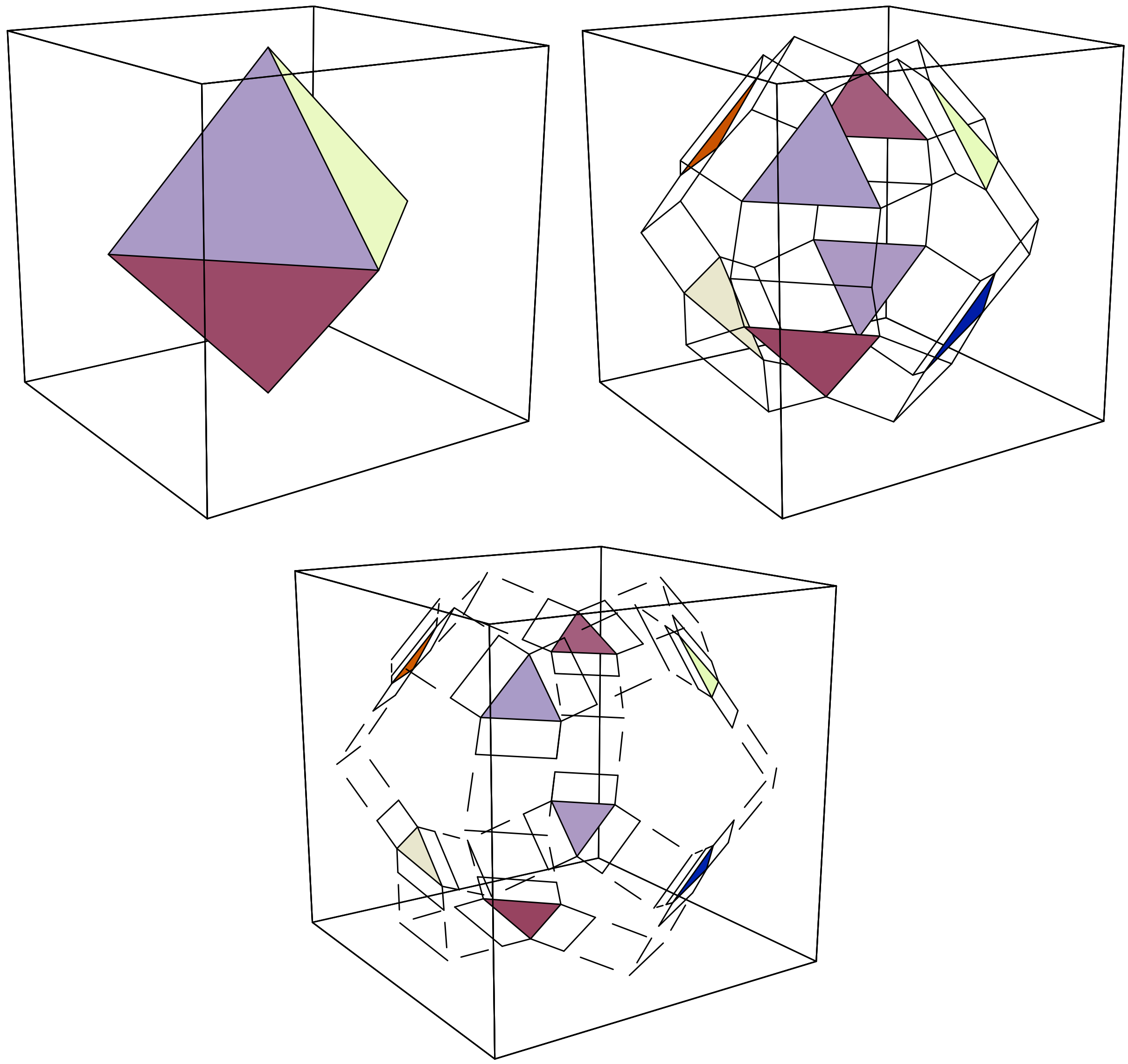

Figure 2: Global states for zero, one, and two-round protocols.

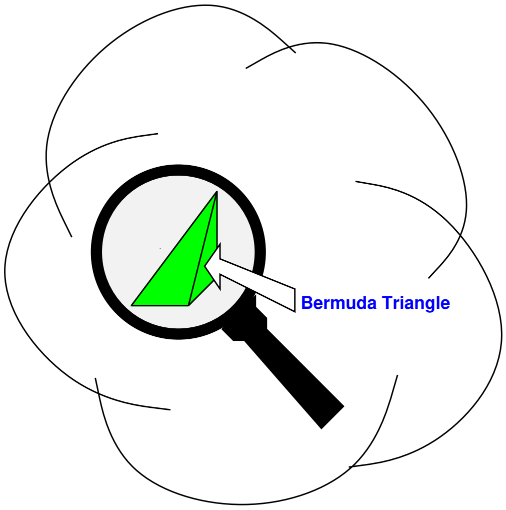

Figure 3: Global states for an r-round protocol (showing the embedded Bermuda Triangle).

is a solid tetrahedron. We can represent a global state for an n-processor protocol as an $(n-1)$-dimensional simplex [Cha93, HS93], where each vertex is labeled with a processor id and local state. If g_1 and g_2 are global states in which p_1 has the same local state, then we "glue together" the vertices of g_1 and g_2 labeled with p_1. Figure 2 shows how these global states glue together in a simple protocol in which each of three processors repeatedly sends its state to the others. Each process begins with a binary input. The first picture shows the possible global states after zero rounds: since no communication has occurred, each processor's state consists only of its input. It is easy to check that the simplices corresponding to these global states form an octahedron. The next picture shows the complex after one round. Each triangle corresponds to a failure-free execution, each free-standing edge to a single-failure execution, and so on. The third picture shows the possible global states after three rounds.

The set of global states after an r-round protocol is quite complicated (Figure 3), but it contains a well-behaved subset of global states which we call the *Bermuda Triangle B*, since all fast protocols vanish somewhere in its interior. The Bermuda Triangle (Figure 4) is constructed by starting with a large k-dimensional simplex, and *triangulating* it into a collection of smaller k-dimensional simplexes. We then label each vertex with an ordered pair (p, s) consisting of a processor identifier p and a local state s in such a way that for each simplex T in the triangulation there is a global state g consistent with the labeling of the simplex: for each ordered pair (p, s) labeling a corner of T, processor p has local state s in global state g.

To illustrate the process of labeling vertices, Figure 5 shows a simplified representation of a two-dimensional Bermuda Triangle B. It is the Bermuda Triangle for

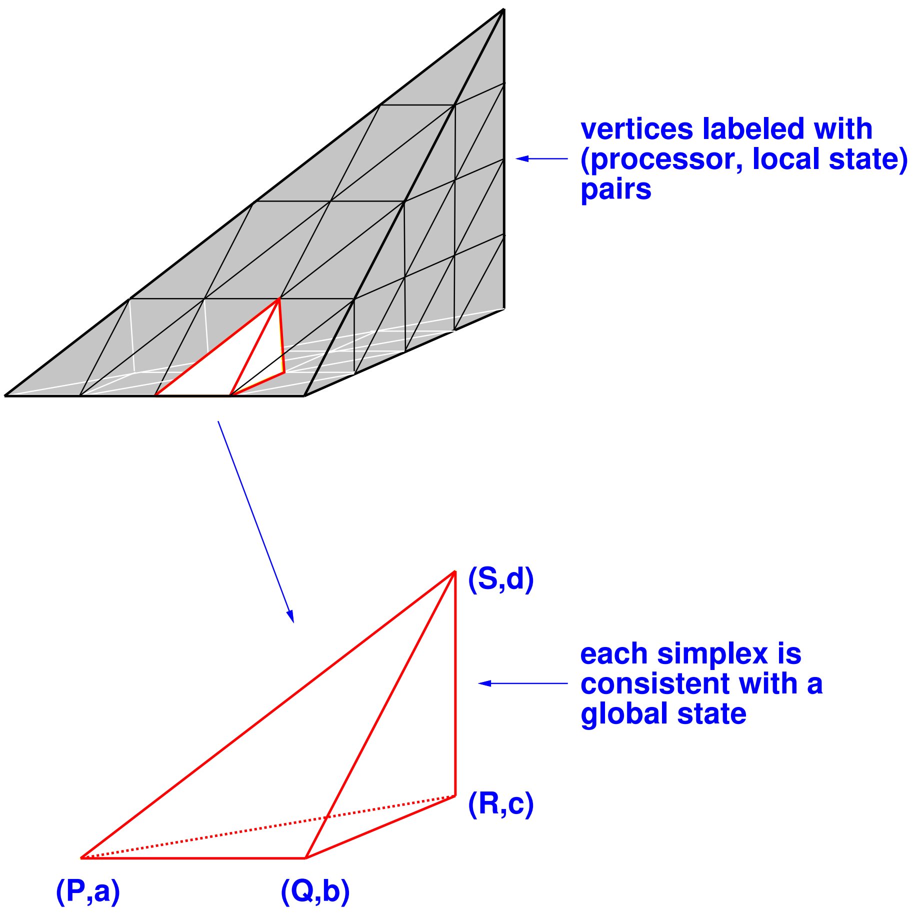

Figure 4: Bermuda Triangle with simplex representing typical global state.

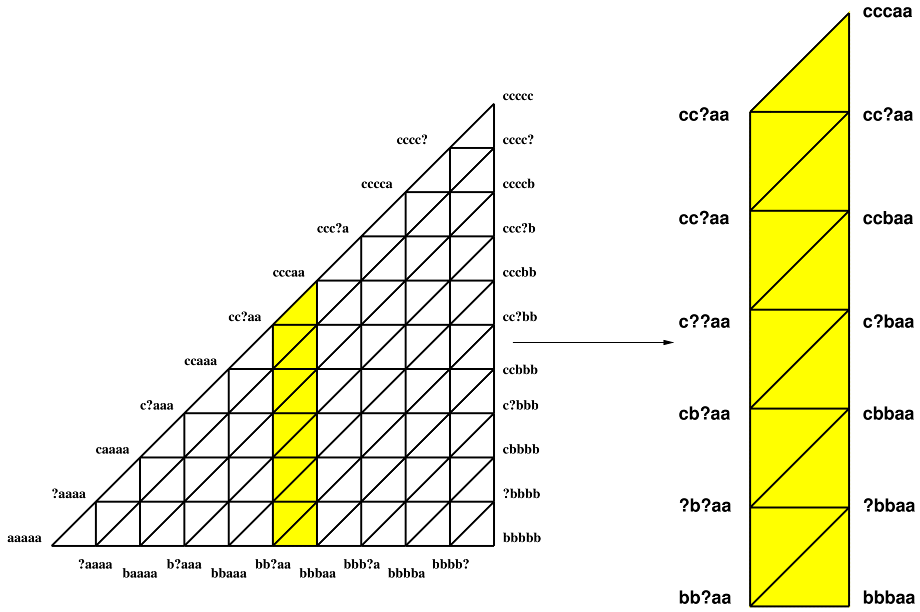

Figure 5: The Bermuda Triangle for 5 processors and a 1-round protocol for 2-set agreement.

a protocol P for 5 processors solving 2-set agreement in 1 round. We have labeled grid points with local states, but we have omitted processor ids and many intermediate nodes for clarity. The local states in the figure are represented by expressions such as $bb?aa$. Given 3 distinct input values a, b, c, we write $bb?aa$ to denote the local state of a processor p at the end of a round in which the first two processors have input value b and send messages to p, the middle processor fails to send a message to p, and the last two processors have input value a and send messages to p. In Figure 5, following any horizontal line from left to right across B, the input values are changed from a to b. The input value of each processor is changed—one after another—by first silencing the processor, and then reviving the processor with the input value b. Similarly, moving along any vertical line from bottom to top, processors' input values change from b to c.

The complete labeling of the Bermuda Triangle B shown in Figure 5—which would include processor ids—has the following property. Let (p, s) be the label of a grid point x. If x is a corner of B, then s specifies that each processor starts with the same input value, so p must choose this value if it finishes protocol P in local state s. If x is on an edge of B, then s specifies that each processor starts with one of the two input values labeling the ends of the edge, so p must choose one of these values if it halts in state s. Similarly, if x is in the interior of B, then s specifies that each processor starts with one of the three values labeling the corners of B, so p must choose one of these three values if it halts in state s.

Now let us "color" each grid point with output values (Figure 6). Given a grid point x labeled with (p, s), let us color x with the value v that p chooses in local state s

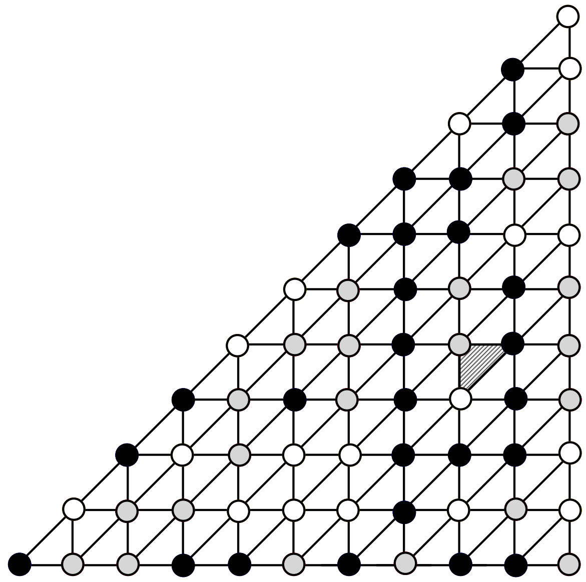

Figure 6: Sperner's Lemma.

at the end of P. This coloring of B has the property that the color of each of the corners is determined uniquely, the color of each point on an edge between two corners is forced to be the color of one of the corners, and the color of each interior point can be the color of any corner. Colorings with this property are called *Sperner colorings*, and have been studied extensively in the field of algebraic topology. At this point, we exploit a remarkable combinatorial result first proved in 1928: *Sperner's Lemma* [Spa66, p.151] states that any Sperner coloring of any triangulated k-dimensional simplex must include at least one simplex whose corners are colored with all $k + 1$ colors. In our case, however, this simplex corresponds to a global state in which $k + 1$ processors choose $k+1$ distinct values, which contradicts the definition of k-set agreement. Thus, in the case illustrated above, there is no protocol for 2-set agreement halting in 1 round.

We note that the basic structure of the Bermuda Triangle and the idea of coloring the vertices with decision values and applying Sperner's Lemma have appeared in previous work by Chaudhuri [Cha91, Cha93]. In that work, she also proved a lower bound of $\lfloor f/k \rfloor + 1$ rounds for k-set agreement, but for a very restricted class of protocols. In particular, a protocol's decision function can depend only on vectors giving partial information about which processors started with which input values, but cannot depend on any other information in a processor's local state, such as processor identities or message histories. The technical challenge in this paper is to construct a labeling of vertices with processor ids and local states that will allow us to prove a lower bound for k-set agreement for arbitrary protocols.

Our approach consists of four parts. First, we label points on the edges of B with global states. For example, consider the edge between the corner where all processors start with input value a and the corner where all processors start with b. We construct a long sequence of global states that begins with a global state in which all processors start with a, ends with a global state in which all processors start with b, and in between systematically changes input values from a to b. These changes are made so gradually, however, that for any two adjacent global states in the sequence, at most one processor can distinguish them. Second, we label each remaining point using a combination of the global states on the edges. Third, we assign nonfaulty processors to points in such a way that the processor labeling a point has the same local state in the global states labeling all adjacent points. Finally, we project each global state onto the associated nonfaulty processor's local state, and label the point with the resulting processor-state pair.

3 The Model

We use a synchronous, message-passing model with crash failures. The system consists of n processors, $p_1, \ldots, p_n$. Processors share a global clock that starts at 0 and advances in increments of 1. Computation proceeds in a sequence of *rounds*, with round r lasting from time $r-1$ to time r. Computation in a round consists of three phases: first each processor p sends messages to some of the processors in the system, possibly including itself, then it receives the messages sent to it during the round, and finally it performs some local computation and changes state. We assume that the communication network is totally connected: every processor is able to send distinct messages to every other processor in every round. We also assume that communication is reliable (although processors can fail): if p sends a message to q in round r, then the message is delivered to q in round r.

Processors follow a deterministic *protocol* that determines what messages a processor should send and what output a processor should generate. A protocol has two components: a *message component* that maps a processor's local state to the list of messages it should send in the next round, and an *output component* that maps a processor's local state to the output value (if any) that it should choose. Processors can be faulty, however, and any processor p can simply *stop* in any round r. In this case, processor p follows its protocol and sends all messages the protocol requires in rounds 1 through $r-1$, sends some subset of the messages it is required to send in round r, and sends no messages in rounds after r. We say that p is *silent* from round r if p sends no messages in round r or later. We say that p is *active* through round r if p sends all messages in round r and earlier.

A *full-information protocol* is one in which every processor broadcasts its entire local state to every processor, including itself, in every round [PSL80, FL82, Had83]. One nice property of full-information protocols is that every execution of a full-information protocol P has a compact representation called a *communication graph* [MT88]. The communication graph $\mathcal{G}$ for an r-round execution of P is a two-dimensional two-colored graph. The vertices form an $n \times r$ grid, with processor names 1 through n labeling the vertical axis and times 0 through r labeling the horizontal axis. The node representing processor p at time i is labeled with the pair $\langle p, i\rangle$. Given any pair of processors p and q and any round i, there is an edge between $\langle p, i-1\rangle$ and $\langle q, i\rangle$ whose color determines whether p successfully sends a message to q in

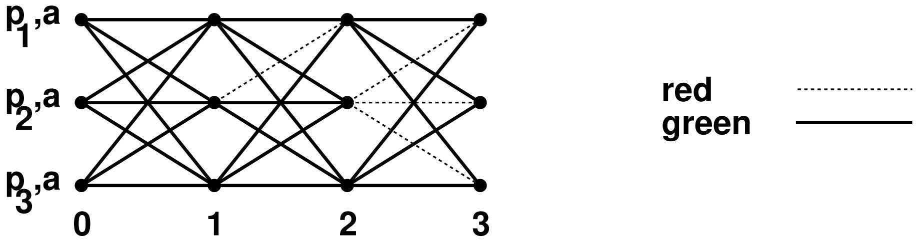

Figure 7: A three-round communication graph.

round i: the edge is green if p succeeds, and red otherwise. In addition, each node $\langle p, 0\rangle$ is labeled with p's input value. Figure 7 illustrates a three round communication graph. In this figure, green edges are denoted by solid lines and red edges by dashed lines. We refer to the edge between $\langle p, i-1\rangle$ and $\langle q, i\rangle$ as the *round i edge from p to q*, and we refer to the node $\langle p, i-1\rangle$ as the *round i node for p* since it represents the point at which p sends its round i messages. We define what it means for a processor to be *silent* or *active* in terms of communication graphs in the obvious way.

In the crash failure model, a processor is silent in all rounds following the round in which it stops. This means that all communication graphs representing executions in this model have the property that if a round i edge from p is red, then all round $j \geq i+1$ edges from p are red, which means that p is silent from round $i+1$. We assume that all communication graphs in this paper have this property, and we note that every r-round graph with this property corresponds to an r-round execution of P.

Since a communication graph $\mathcal{G}$ describes an execution of P, it also determines the global state at the end of P, so we sometimes refer to $\mathcal{G}$ as a *global communication graph*. In addition, for each processor p and time t, there is a subgraph of $\mathcal{G}$ that corresponds to the local state of p at the end round t, and we refer to this subgraph as a *local communication graph*. The local communication graph for p at time t is the subgraph $\mathcal{G}(p, t)$ of $\mathcal{G}$ containing all the information visible to p at the end of round t. Namely, $\mathcal{G}(p, t)$ is the subgraph induced by the node $\langle p, t\rangle$ and all earlier nodes reachable from $\langle p, t\rangle$ by a sequence (directed backwards in time) of green edges followed by at most one red edge. In the remainder of this paper, we use graphs to represent states. Wherever we used "state" in the informal overview of Section 2, we now substitute the word "graph." Furthermore, we defined a full-information protocol to be a protocol in which processors broadcast their local states in every round, but we now assume that processors broadcast their local communication graphs instead. In addition, we assume that all executions of a full-information protocol run for exactly r rounds and produce output at exactly time r. All local and global communication graphs are graphs at time r, unless otherwise specified.

The crucial property of a full-information protocol is that every protocol can be simulated by a full-information protocol, and hence that we can restrict attention to full-information protocols when proving the lower bound in this paper:

Lemma 1: If there is an n-processor protocol solving k-set agreement with f failures in r rounds, then there is an n-processor full-information protocol solving k-set agreement with f failures in r rounds.

4 The k-set Agreement Problem

The *k-set agreement problem* [Cha91] is defined as follows. We assume that each processor p_i has two private registers in its local state, a read-only input register and a write-only output register. Initially, p_i's input register contains an arbitrary input value from a set V containing at least $k+1$ values $v_0, \ldots, v_k$, and its output register is empty. A protocol solves the problem if it causes each processor to halt after writing an output value to its output register in such a way that

1. every processor's output value is some processor's input value, and
2. the set of output values chosen has size at most k.

5 Bermuda Triangle

In this section, we define the basic geometric constructs used in our proof that every protocol P solving k-set agreement and tolerating f failures requires at least $\lfloor f/k \rfloor + 1$ rounds of communication, assuming $n \geq f + k + 1$.

We start with some preliminary definitions. A *simplex* S is the convex hull of $k+1$ affinely-independent[1] points $x_0, \ldots, x_k$ in Euclidean space. It is a k-dimensional volume, the k-dimensional analogue of a solid triangle or tetrahedron. The points $x_0, \ldots, x_k$ are called the *vertices* of S, and k is the *dimension* of S. We sometimes call S a *k-simplex* when we wish to emphasize its dimension. A simplex F is a *face* of S if the vertices of F form a subset of the vertices of S (which means that the dimension of F is at most the dimension of S). A set of k-simplexes $S_1, \ldots, S_\ell$ is a *triangulation* of S if $S = S_1 \cup \cdots \cup S_\ell$ and the intersection of S_i and S_j is a face of each[2] for all pairs i and j. The *vertices* of a triangulation are the vertices of the S_i. Any triangulation of S induces triangulations of its faces in the obvious way.

The construction of the Bermuda Triangle is illustrated in Figure 8. Let $\mathcal{B}$ be the k-simplex in k-dimensional Euclidean space with vertices

$$(0, \ldots, 0), (N, 0, \ldots, 0), (N, N, 0, \ldots, 0), \ldots, (N, \ldots, N),$$

where N is a huge integer defined later in Section 6.3. The *Bermuda Triangle* B is a triangulation of $\mathcal{B}$ defined as follows. The vertices of B are the grid points contained in $\mathcal{B}$: these are the points of the form $x = (x_1, \ldots, x_k)$, where the x_i are integers between 0 and N satisfying $x_1 \geq x_2 \geq \cdots \geq x_k$.

Informally, the simplexes of the triangulation are defined as follows: pick any grid point and walk one step in the positive direction along each dimension (Figure 9).

[1] Points $x_0, \ldots, x_k$ are affinely independent if $x_1 - x_0, \ldots, x_k - x_0$ are linearly independent.

[2] Notice that the intersection of two arbitrary k-dimensional simplexes S_i and S_j will be a volume of some dimension, but it need not be a face of either simplex.

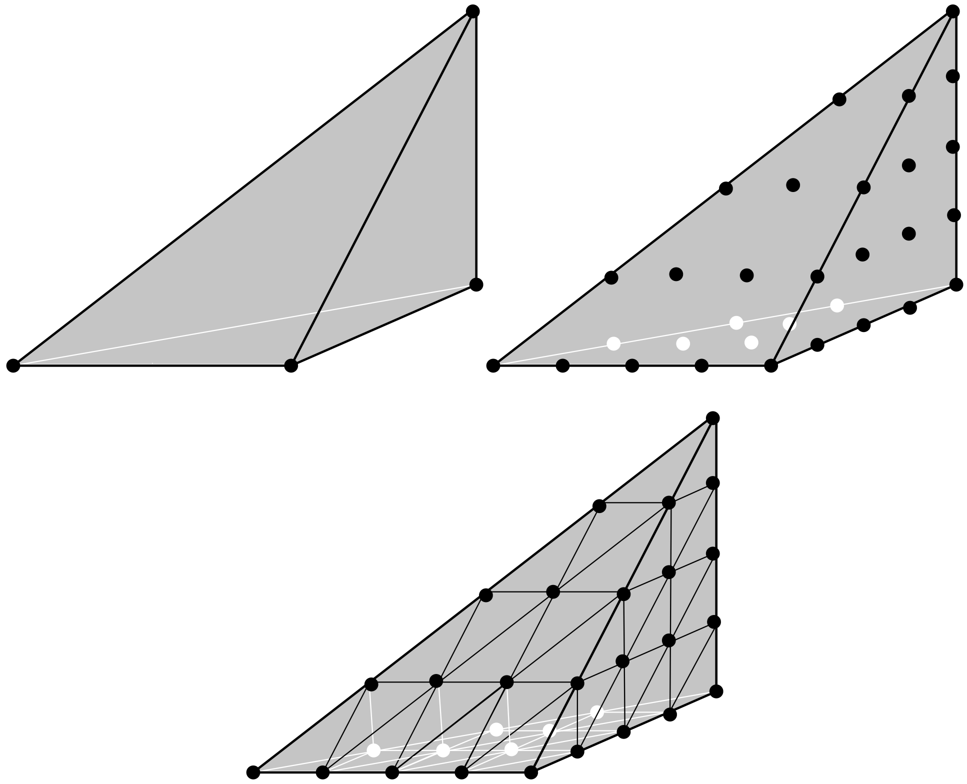

Figure 8: Construction of Bermuda Triangle.

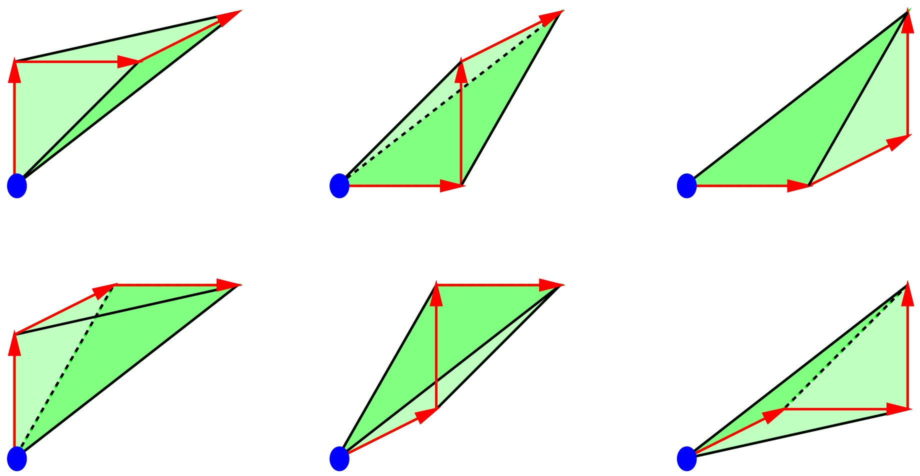

Figure 9: Simplex generation in Kuhn's triangulation.

The $k + 1$ points visited by this walk define the vertices of a simplex, and the triangulation B consists of all simplexes determined by such walks. For example, the 2-dimensional Bermuda Triangle is illustrated in Figure 5. This triangulation, known as *Kuhn's triangulation*, is defined formally as follows [Cha93]. Let $e_1, \ldots, e_k$ be the unit vectors; that is, e_i is the vector $(0, \ldots, 1, \ldots, 0)$ with a single 1 in the ith coordinate. A simplex is determined by a point y_0 and an arbitrary permutation $f_1, \ldots, f_k$ of the unit vectors $e_1, \ldots, e_k$: the vertices of the simplex are the points $y_i = y_{i-1} + f_i$ for all $i > 0$. When we list the vertices of a simplex, we always write them in the order $y_0, \ldots, y_k$ in which they are visited by the walk.

For brevity, we refer to the vertices of $\mathcal{B}$ as the *corners* of B. The "edges" of $\mathcal{B}$ are partitioned to form the edges of B. More formally, the triangulation B induces triangulations of the one-dimensional faces (line segments connecting the vertices) of $\mathcal{B}$, and these induced triangulations are called the *edges* of B. The simplexes of B are called *primitive simplexes*.

Each vertex of B is labeled with an ordered pair $(p, \mathcal{L})$ consisting of a processor id p and a local communication graph $\mathcal{L}$. As illustrated in the overview in Section 2, the crucial property of this labeling is that if S is a primitive simplex with vertices $y_0, \ldots, y_k$, and if each vertex y_i is labeled with a pair $(q_i, \mathcal{L}_i)$, then there is a global communication graph $\mathcal{G}$ such that each q_i is nonfaulty in $\mathcal{G}$ and has local communication graph $\mathcal{L}_i$ in $\mathcal{G}$. Constructing this labeling is the subject of the next three sections. We first assign global communication graphs $\mathcal{G}$ to vertices in Section 6, then we assign processors p to vertices in Section 7, and then we assign ordered pairs $(p, \mathcal{L})$ to vertices in Section 8, where $\mathcal{L}$ is the local communication graph of p in $\mathcal{G}$.

6 Graph Assignment

In this section, we label each vertex of B with a global communication graph. Actually, for expository reasons, we augment the definition of a communication graph and label vertices of B with these augmented communication graphs instead. Constructing this labeling involves several steps. We define operations on augmented communication graphs that make minor changes in the graphs, and we use these operations to construct long sequences of graphs. Then we label vertices along edges of B with graphs from these sequences, and we label interior vertices of B by performing a merge of the graphs labeling the edges.

6.1 Augmented Communication Graphs

We extend the definition of a communication graph to make the processor assignment in Section 7 easier to describe. We augment communication graphs with tokens, and place tokens on the graph so that if processor p fails in round i, then there is a token on the node $\langle p, j-1 \rangle$ for processor p in some earlier round $j \leq i$ (Figure 10). In this sense, every processor failure is "covered" by a token, and the number of processors failing in the graph is bounded from above by the number of tokens. In the next few sections, when we construct long sequences of these graphs, tokens will be moved between adjacent processors within a round, and used to guarantee that processor failures

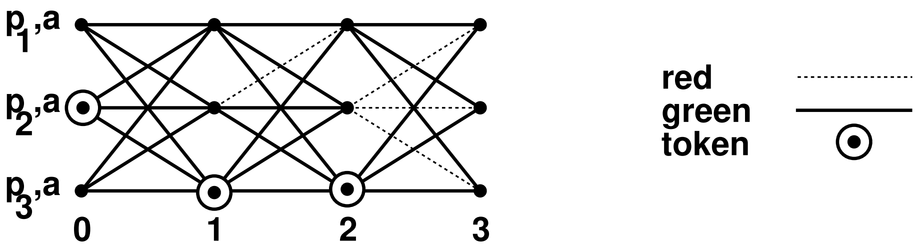

Figure 10: Three-round communication graph with one token per round.

in adjacent graphs change in a orderly fashion. For every value of ℓ, we define graphs with exactly ℓ tokens placed on nodes in each round, but we will be most interested in the two cases with ℓ equal to 1 and k.

For each value $\ell > 0$, we define an *ℓ-graph* $\mathcal{G}$ to be a communication graph with tokens placed on the nodes of the graph that satisfies the following conditions for each round i, $1 \leq i \leq r$:

1. The total number of tokens on round i nodes is exactly ℓ.
2. If a round i edge from p is red, then there is a token on a round $j \leq i$ node for p.
3. If a round i edge from p is red, then p is silent from round $i + 1$.

We say that p is *covered by a round i token* if there is a token on the round i node for p, we say that p is *covered in round i* if p is covered by a round $j \leq i$ token, and we say that p is *covered* in a graph if p is covered in any round. Similarly, we say that a round i edge from p is covered if p is covered in round i. The second condition says every red edge is covered by a token, and this together with the first condition implies that at most ℓr processors fail in an ℓ-graph. We often refer to an ℓ-graph as a *graph* when the value of ℓ is clear from context or unimportant. We emphasize that the tokens are simply an accounting trick, and have no meaning as part of the global or local state in the underlying communication graph.

We define a *failure-free* ℓ-graph to be an ℓ-graph in which all edges are green, and all round i tokens are on processor p_1 in all rounds i.

6.2 Graph operations

We now define four operations on augmented graphs that make only minor changes to a graph. In particular, the only change an operation makes is to change the color of a single edge, to change the value of a single processor's input, or to move a single token between adjacent processors within the same round. The operations are defined as follows (see Figure 11):

1. *delete*(i, p, q): This operation changes the color of the round i edge from p to q to red, and has no effect if the edge is already red. This makes the delivery of the

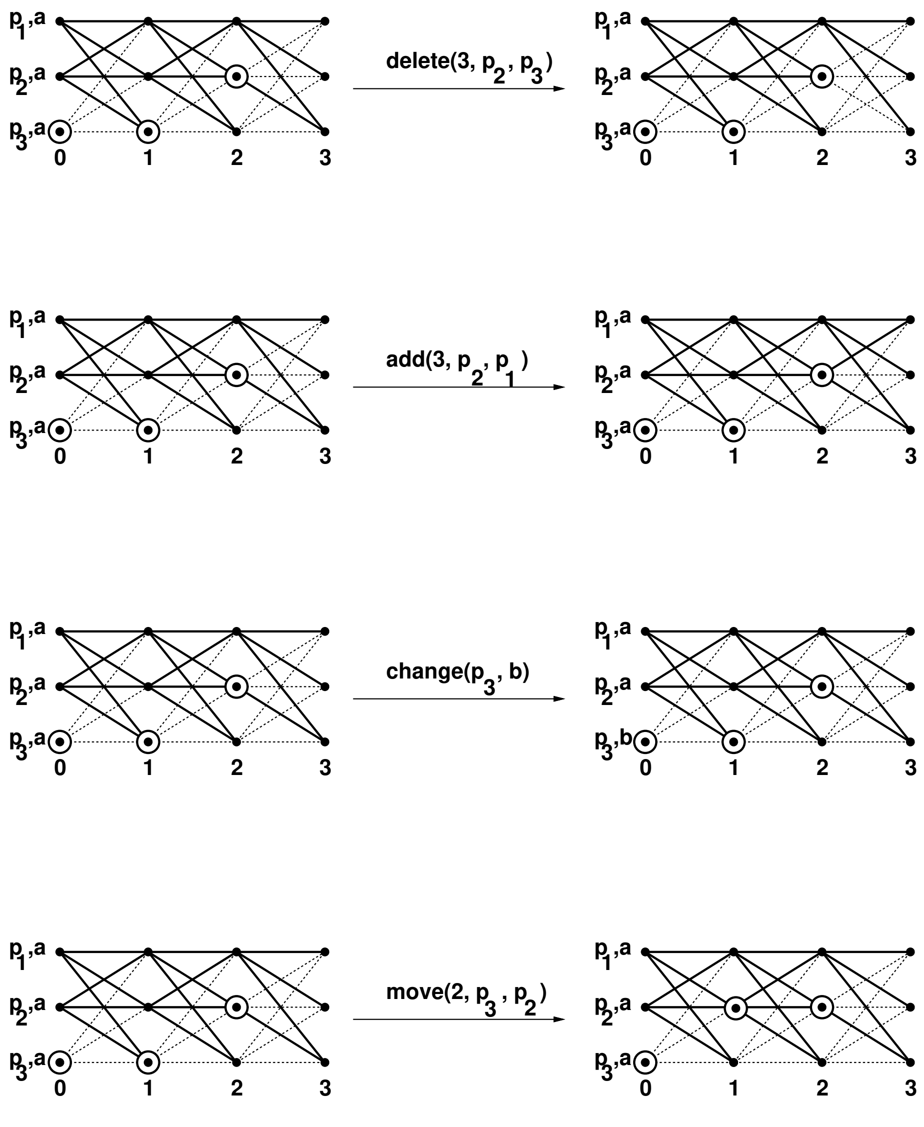

Figure 11: Operations on augmented communication graphs.

round i message from p to q unsuccessful. It can only be applied to a graph if p and q are silent from round $i+1$, and p is covered in round i.

2. *add*(i, p, q): This operation changes the color of the round i edge from p to q to green, and has no effect if the edge is already green. This makes the delivery of the round i message from p to q successful. It can only be applied to a graph if p and q are silent from round $i+1$, processor p is active through round $i-1$, and p is covered in round i.

3. *change*(p, v): This operation changes the input value for processor p to v, and has no effect if the value is already v. It can only be applied to a graph if p is silent from round 1, and p is covered in round 1.

4. *move*(i, p, q): This operation moves a round i token from $\langle p, i-1\rangle$ to $\langle q, i-1\rangle$, and is defined only for adjacent processors p and q (that is, $\{p, q\} = \{p_j, p_{j+1}\}$ for some j). It can only be applied to a graph if p is covered by a round i token, and all red edges are covered by other tokens.

It is obvious from the definition of these operations that they preserve the property of being an ℓ-graph: if $\mathcal{G}$ is an ℓ-graph and τ is a graph operation, then $\tau(\mathcal{G})$ is an ℓ-graph. We define *delete*, *add*, and *change* operations on communication graphs in exactly the same way, except that the condition "p is covered in round i" is omitted.

6.3 Graph sequences

We now define a sequence $\sigma[v]$ of graph operations that can be applied to any failure-free graph $\mathcal{G}$ to transform it into the failure-free graph $\mathcal{G}[v]$ in which all processors have input v. We want to emphasize that the sequences $\sigma[v]$ differ only in the value v. For this reason, we define a parameterized sequence $\sigma[\mathrm{V}]$ with the property that for all values v and all graphs $\mathcal{G}$, the sequence $\sigma[v]$ transforms $\mathcal{G}$ to $\mathcal{G}[v]$. In general, we define a *parameterized sequence* $\sigma[\mathrm{X}_1, \ldots, \mathrm{X}_\ell]$ to be a sequence of graph operations with free variables $\mathrm{X}_1, \ldots, \mathrm{X}_\ell$ appearing as parameters to the graph operations in the sequence.

Given a graph $\mathcal{G}$, let *red*$(\mathcal{G}, p, m)$ and *green*$(\mathcal{G}, p, m)$ be graphs identical to $\mathcal{G}$ except that all edges from p in rounds $m, \ldots, r$ are red and green, respectively. We define these graphs only if

1. p is covered in round m in $\mathcal{G}$,

2. all faulty processors are silent from round m (or earlier) in $\mathcal{G}$, and

3. and all tokens are on p_1 in rounds $m+1, \ldots, r$ in $\mathcal{G}$.

In addition, we define the graph *green*$(\mathcal{G}, p, m)$ only if

4. p is active through round $m-1$ in $\mathcal{G}$.

These restrictions guarantee that if $\mathcal{G}$ is an ℓ-graph and *red*$(\mathcal{G}, p, m)$ and *green*$(\mathcal{G}, p, m)$ are defined, then *red*$(\mathcal{G}, p, m)$ and *green*$(\mathcal{G}, p, m)$ are both ℓ-graphs.

In the case of ordinary communication graphs, a result by Moses and Tuttle [MT88] implies that there is a "similarity chain" of graphs between $\mathcal{G}$ and $red(\mathcal{G}, p, m)$ and between $\mathcal{G}$ and $green(\mathcal{G}, p, m)$. In their proof—a refinement of similar proofs by Dwork and Moses [DM90] and others—the sequence of graphs they construct has the property that each graph in the chain can be obtained from the preceding graph by applying a sequence of the *add*, *delete*, and *change* graph operations defined above. The same proof works for augmented communication graphs, provided we insert *move* operations between the *add*, *delete*, and *change* operations to move tokens between nodes appropriately. With this modification, we can prove the following. Let $faulty(\mathcal{G})$ be the set of processors that fail in $\mathcal{G}$.

Lemma 2: For every processor p, round m, and set π of processors, there are sequences $silence_\pi(p, m)$ and $revive_\pi(p, m)$ such that for all graphs $\mathcal{G}$:

1. If $red(\mathcal{G}, p, m)$ is defined and $\pi = faulty(\mathcal{G})$, then
$$silence_\pi(p, m)(\mathcal{G}) = red(\mathcal{G}, p, m).$$

2. If $green(\mathcal{G}, p, m)$ is defined and $\pi = faulty(\mathcal{G})$, then
$$revive_\pi(p, m)(\mathcal{G}) = green(\mathcal{G}, p, m).$$

Proof: We proceed by reverse induction on m. Suppose $m = r$. Define

$$\begin{aligned} silence_\pi(p, r) &= delete(r, p, p_1) \cdots delete(r, p, p_n) \\ revive_\pi(p, r) &= add(r, p, p_1) \cdots add(r, p, p_n). \end{aligned}$$

For part 1, let $\mathcal{G}$ be any graph and suppose $red(\mathcal{G}, p, r)$ is defined. For each i with $0 \leq i \leq n$, let $\mathcal{G}_i$ be the graph identical to $\mathcal{G}$ except that the round r edges from p to $p_1, \ldots, p_i$ are red. Since $red(\mathcal{G}, p, r)$ is defined, condition 1 implies that p is covered in round r in $\mathcal{G}$. For each i with $1 \leq i \leq n$, it follows that $\mathcal{G}_{i-1}$ is really a graph, and $delete(r, p, p_i)$ can be applied to $\mathcal{G}_{i-1}$ and transforms it to $\mathcal{G}_i$. Since $\mathcal{G} = \mathcal{G}_0$ and $\mathcal{G}_n = red(\mathcal{G}, p, r)$, it follows that $silence_\pi(p, r)$ transforms $\mathcal{G}$ to $red(\mathcal{G}, p, r)$. For part 2, let $\mathcal{G}$ be any graph and suppose $green(\mathcal{G}, p, r)$ is defined. The proof of this part is the direct analogue of the proof of part 1. The only difference is that since we are coloring round r edges from p green instead of red, we must verify that p is active through round $r - 1$ in $\mathcal{G}$, but this follows immediately from condition 4.

Suppose $m < r$ and the induction hypothesis holds for $m + 1$. Define $\pi' = \pi \cup \{p\}$ and define

$$\begin{aligned} set(m+1, p_i) &= move(m+1, p_1, p_2) \cdots move(m+1, p_{i-1}, p_i) \\ reset(m+1, p_i) &= move(m+1, p_i, p_{i-1}) \cdots move(m+1, p_2, p_1). \end{aligned}$$

The *set* function moves the token from p_1 to p_i and the *reset* function moves the token back from p_i to p_1.

Define $block(m, p, p_i)$ to be $delete(m, p, p_i)$ if $p_i \in \pi'$, and otherwise

$$silence_{\pi'}(p_i, m+1) \quad \begin{array}{c} set(m+1, p_i) \\ delete(m, p, p_i) \\ reset(m+1, p_i). \end{array} \quad revive_{\pi' \cup \{p_i\}}(p_i, m+1)$$

Define $unblock(m, p, p_i)$ to be $add(m, p, p_i)$ if $p_i \in \pi'$, and otherwise

$$silence_{\pi'}(p_i, m+1) \quad \begin{array}{c} set(m+1, p_i) \\ add(m, p, p_i) \\ reset(m+1, p_i). \end{array} \quad revive_{\pi' \cup \{p_i\}}(p_i, m+1)$$

Finally, define

$$\begin{aligned} block(m, p) &= block(m, p, p_1) \cdots block(m, p, p_n) \\ unblock(m, p) &= unblock(m, p, p_1) \cdots unblock(m, p, p_n) \end{aligned}$$

and define

$$\begin{aligned} silence_\pi(p, m) &= silence_\pi(p, m+1)\, block(m, p) \\ revive_\pi(p, m) &= silence_\pi(p, m+1)\, unblock(m, p)\, revive_{\pi'}(p, m+1). \end{aligned}$$

For part 1, let $\mathcal{G}$ be any graph, and suppose $red(\mathcal{G}, p, m)$ is defined and $\pi = faulty(\mathcal{G})$. Since $red(\mathcal{G}, p, m)$ is defined, the graph $red(\mathcal{G}, p, m+1)$ is also defined, and the induction hypothesis for $m+1$ states that $silence_\pi(p, m+1)$ transforms $\mathcal{G}$ to $red(\mathcal{G}, p, m+1)$. We now show that $block(m, p)$ transforms $red(\mathcal{G}, p, m+1)$ to $red(\mathcal{G}, p, m)$, and we will be done. For each i with $0 \le i \le n$, let $\mathcal{G}_i$ be the graph identical to $\mathcal{G}$ except that p is silent from round $m+1$ and the edges from p to $p_1, \ldots, p_i$ are red in $\mathcal{G}_i$. Since $red(\mathcal{G}, p, m)$ is defined, condition 1 implies that p is covered in round m in $\mathcal{G}$. For each i with $0 \le i \le n$, it follows that $\mathcal{G}_i$ really is a graph and that $\pi' = faulty(\mathcal{G}_i)$. Since $red(\mathcal{G}, p, m+1) = \mathcal{G}_0$ and $\mathcal{G}_n = red(\mathcal{G}, p, m)$, it is enough to show that $block(m, p, p_i)$ transforms $\mathcal{G}_{i-1}$ to $\mathcal{G}_i$ for each i with $1 \le i \le n$. The proof of this fact depends on whether $p_i \in \pi'$, so we consider two cases.

Consider the easy case with $p_i \in \pi'$. We know that p is covered in round m in $\mathcal{G}_{i-1}$ since it is covered in $\mathcal{G}$ by condition 1. We know that p is silent from round $m+1$ in $\mathcal{G}_{i-1}$ since it is silent in $\mathcal{G}_0 = red(\mathcal{G}, p, m+1)$. We know that p_i is silent from round $m+1$ in $\mathcal{G}_{i-1}$ since $p_i \in \pi'$ implies (assuming that p_i is not just p again) that p_i fails in $\mathcal{G}$, and hence is silent from round $m+1$ in $\mathcal{G}$ by condition 2. This means that $block(m, p, p_i) = delete(m, p, p_i)$ can be applied to $\mathcal{G}_{i-1}$ to transform $\mathcal{G}_{i-1}$ to $\mathcal{G}_i$.

Now consider the difficult case when $p_i \notin \pi'$. Let $\mathcal{H}_{i-1}$ and $\mathcal{H}_i$ be graphs identical to $\mathcal{G}_{i-1}$ and $\mathcal{G}_i$, except that a single round $m+1$ token is on p_i in $\mathcal{H}_{i-1}$ and $\mathcal{H}_i$. Condition 3 guarantees that all round $m+1$ tokens are on p_1 in $\mathcal{G}$, and hence in $\mathcal{G}_{i-1}$ and $\mathcal{G}_i$, so $\mathcal{H}_{i-1}$ and $\mathcal{H}_i$ really are graphs. In addition, $set(m+1, p_i)$ transforms $\mathcal{G}_{i-1}$ to $\mathcal{H}_{i-1}$, and $reset(m+1, p_i)$ transforms $\mathcal{H}_i$ to $\mathcal{G}_i$. Let $\mathcal{I}_{i-1}$ and $\mathcal{I}_i$ be identical to $\mathcal{H}_{i-1}$ and $\mathcal{H}_i$ except that p_i is silent from round $m+1$ in $\mathcal{I}_{i-1}$ and $\mathcal{I}_i$. Processor p_i is covered in round $m+1$ in $\mathcal{H}_{i-1}$ and $\mathcal{H}_i$, so $\mathcal{I}_{i-1}$ and $\mathcal{I}_i$ really are graphs. In fact, p_i does not fail in $\mathcal{G}$ since $p_i \notin \pi'$, so p_i is active through round m in $\mathcal{I}_{i-1}$ and $\mathcal{I}_i$, so $\mathcal{I}_{i-1} = red(\mathcal{H}_{i-1}, p_i, m+1)$ and $\mathcal{H}_i = green(\mathcal{I}_i, p_i, m+1)$. The inductive hypothesis for $m+1$

states that $silence_{\pi'}(p_i, m+1)$ transforms $\mathcal{H}_{i-1}$ to $\mathcal{I}_{i-1}$, and $revive_{\pi'\cup\{p_i\}}(p_i, m+1)$ transforms $\mathcal{I}_i$ to $\mathcal{H}_i$. Finally, notice that the only difference between $\mathcal{I}_{i-1}$ and $\mathcal{I}_i$ is the color of the round m edge from p to p_i. Since p is covered in round m and p and p_i are silent from round $m+1$ in both graphs, we know that $delete(m, p, p_i)$ transforms $\mathcal{I}_{i-1}$ to $\mathcal{I}_i$. It follows that $block(m, p, p_i)$ transforms $\mathcal{G}_{i-1}$ to $\mathcal{G}_i$, and we are done.

For part 2, let $\mathcal{G}$ be any graph and suppose $green(\mathcal{G}, p, m)$ is defined and $\pi = faulty(\mathcal{G})$. Since $green(\mathcal{G}, p, m)$ is defined, let $\mathcal{G}' = green(\mathcal{G}, p, m)$. Now let $\mathcal{H}$ and $\mathcal{H}'$ be graphs identical to $\mathcal{G}$ and $\mathcal{G}'$ except that p is silent from round $m+1$ in $\mathcal{H}$ and $\mathcal{H}'$. Since $green(\mathcal{G}, p, m)$ is defined, processor p is covered in round m in $\mathcal{G}$ by condition 1 and hence in $\mathcal{G}'$, so $\mathcal{H}$ and $\mathcal{H}'$ really are graphs. In addition, since $green(\mathcal{G}, p, m)$ is defined, processor p is active through round $m-1$ in $\mathcal{G}$ by condition 4, so processor p is active through round m in $\mathcal{G}'$ and $\mathcal{H}'$. This means that $green(\mathcal{H}', p, m+1)$ is defined, and in fact we have $\mathcal{H} = red(\mathcal{G}, p, m+1)$ and $\mathcal{G}' = green(\mathcal{H}', p, m+1)$. The induction hypothesis for $m+1$ states that $silence_{\pi}(p, m+1)$ transforms $\mathcal{G}$ to $\mathcal{H}$ and that $revive_{\pi'}(p, m+1)$ transforms $\mathcal{H}'$ to $\mathcal{G}'$. To complete the proof, we need only show that $unblock(m, p)$ transforms $\mathcal{H}$ to $\mathcal{H}'$. The proof of this fact is the direct analogue of the proof in part 1 that $block(m, p)$ transforms $red(\mathcal{G}, p, m+1)$ to $red(\mathcal{G}, p, m)$. The only difference is that since we are coloring round m edges from p with green instead of red, we must verify that p is active through round $m-1$ in the graphs $\mathcal{H}_i$ analogous to $\mathcal{G}_i$ in the proof of part 1, but this follows immediately from condition 4. $\square$

Given a graph $\mathcal{G}$, let $\mathcal{G}_i[v]$ be a graph identical to $\mathcal{G}$, except that processor p_i has input v. Using the preceding result, we can transform $\mathcal{G}$ to $\mathcal{G}_i[v]$.

Lemma 3: For each i, there is a parameterized sequence $\sigma_i[\mathrm{V}]$ with the property that for all values v and failure-free graphs $\mathcal{G}$, the sequence $\sigma_i[v]$ transforms $\mathcal{G}$ to $\mathcal{G}_i[v]$.

Proof: Define

$$\begin{aligned} set(1, p_i) &= move(1, p_1, p_2) \cdots move(1, p_{i-1}, p_i) \\ reset(1, p_i) &= move(1, p_i, p_{i-1}) \cdots move(1, p_2, p_1) \end{aligned}$$

and define

$$\sigma_i[\mathrm{V}] = set(1, p_i) silence_{\emptyset}(p_i, 1) change(p_i, \mathrm{V}) revive_{\{p_i\}}(p_i, 1) reset(1, p_i)$$

where $\emptyset$ denotes the empty set. Now consider any value v and any failure-free graph $\mathcal{G}$, and let $\mathcal{G}' = \mathcal{G}_i[v]$. Since $\mathcal{G}$ and $\mathcal{G}'$ are failure-free graphs, all round 1 tokens are on p_1, so let $\mathcal{H}$ and $\mathcal{H}'$ be graphs identical to $\mathcal{G}$ and $\mathcal{G}'$ except that a single round 1 token is on p_i in $\mathcal{H}$ and $\mathcal{H}'$. We know that $\mathcal{H}$ and $\mathcal{H}'$ are graphs, and that $set(1, p_i)$ transforms $\mathcal{G}$ to $\mathcal{H}$ and $reset(1, p_i)$ transforms $\mathcal{H}'$ to $\mathcal{G}'$. Since p_i is covered in $\mathcal{H}$ and $\mathcal{H}'$, let $\mathcal{I}$ and $\mathcal{I}'$ be identical to $\mathcal{H}$ and $\mathcal{H}'$ except that p_i is silent from round 1. We know that $\mathcal{I}$ and $\mathcal{I}'$ are graphs, and it follows by Lemma 2 that $silence_{\emptyset}(p_i, 1)$ transforms $\mathcal{H}$ to $\mathcal{I}$ and that $revive_{\{p_i\}}(p_i, 1)$ transforms $\mathcal{I}'$ to $\mathcal{H}'$. Finally, notice that $\mathcal{I}$ and $\mathcal{I}'$ differ only in the input value for p_i. Since p_i is covered and silent from round 1 in both graphs, the operation $change(p_i, v)$ can be applied to $\mathcal{I}$ and transforms it to $\mathcal{I}'$. Stringing all of this together, it follows that $\sigma_i[v]$ transforms $\mathcal{G}$ to $\mathcal{G}' = \mathcal{G}_i[v]$. $\square$

By concatenating such operation sequences, we can transform $\mathcal{G}$ into $\mathcal{G}[v]$ by changing processors' input values one at a time:

Lemma 4: Let $\sigma[\mathrm{V}] = \sigma_1[\mathrm{V}] \cdots \sigma_n[\mathrm{V}]$. For every value v and failure-free graph $\mathcal{G}$, the sequence $\sigma[v]$ transforms $\mathcal{G}$ to $\mathcal{G}[v]$.

Now we can define the parameter N used in defining the shape of B: N is the length of the sequence $\sigma[\mathrm{V}]$, which is exponential in r.

6.4 Graph merge

Speaking informally, we will use each sequence $\sigma[v_i]$ of graph operations to generate a sequence of graphs, and we will use this sequence of graphs to label vertices along the edge of B in the ith dimension. Then we will label vertices in the interior of B by performing a "merge" of the graphs on the edges in the different dimensions.

The *merge* of a sequence $\mathcal{H}_1, \ldots, \mathcal{H}_k$ of graphs is a graph defined as follows:

1. an edge e is colored red if it is red in any of the graphs $\mathcal{H}_1, \ldots, \mathcal{H}_k$, and green otherwise, and

2. an initial node $\langle p, 0\rangle$ is labeled with the value v_i where i is the maximum index such that $\langle p, 0\rangle$ is labeled with v_i in $\mathcal{H}_i$, or v_0 if no such i exists, and

3. the number of tokens on a node $\langle p, i\rangle$ is the sum of the number of tokens on the node in the graphs $\mathcal{H}_1, \ldots, \mathcal{H}_k$.

The first condition says that a message is missing in the resulting graph if and only if it is missing in any of the merged graphs. To understand the second condition, notice that for each processor p_j there is a integer s_j with the property that p_j's input value in changed to v_i by the s_jth operation appearing in $\sigma[v_i]$. Now choose a vertex $x = (x_1, \ldots, x_k)$ of B, and imagine walking from the origin to x by walking along the first dimension to $(x_1, 0, \ldots, 0)$, then along the second dimension to $(x_1, x_2, 0, \ldots, 0)$, and so forth. In each dimension i, processor p_j's input is changed from v_{i-1} to v_i after s_j steps in this dimension. Since $x_1 \geq x_2 \geq \cdots \geq x_k$, there is a final dimension i in which p_j's input is changed to v_i, and never changed again. The second condition above is just a compact way of identifying this final value v_i.

Lemma 5: Let $\mathcal{H}$ be the merge of the graphs $\mathcal{H}_1, \ldots, \mathcal{H}_k$. If $\mathcal{H}_1, \ldots, \mathcal{H}_k$ are 1-graphs, then $\mathcal{H}$ is a k-graph.

Proof: We consider the three conditions required of a k-graph in turn. First, there are k tokens in each round of $\mathcal{H}$ since there is 1 token in each round of each graph $\mathcal{H}_1, \ldots, \mathcal{H}_k$. Second, every red edge in $\mathcal{H}$ is covered by a token since every red edge in $\mathcal{H}$ corresponds to a red edge in one of the graphs $\mathcal{H}_j$, and this edge is covered by a token in $\mathcal{H}_j$. Third, if there is a red edge from p in round i in $\mathcal{H}$, then there is a red from p in round i of one of the graphs $\mathcal{H}_j$. In this graph, p is silent from round $i+1$, so the same is true in $\mathcal{H}$. Thus, $\mathcal{H}$ is a k-graph. $\square$

6.5 Graph assignments

Now we can define the assignment of graphs to vertices of B. For each value v_i, let $\mathcal{F}_i$ be the failure-free 1-graph in which all processors have input v_i. Let $x = (x_1, \ldots, x_k)$ be an arbitrary vertex of B. For each coordinate x_j, let σ_j be the prefix of $\sigma[v_j]$ consisting of the first x_j operations, and let $\mathcal{H}_j$ be the 1-graph resulting from the application of σ_j to $\mathcal{F}_{j-1}$. This means that in $\mathcal{H}_j$, some set $p_1, \ldots, p_i$ of adjacent processors have had their inputs changed from v_{j-1} to v_j. The graph $\mathcal{G}$ labeling x is defined to be the merge of $\mathcal{H}_1, \ldots, \mathcal{H}_k$. We know that $\mathcal{G}$ is a k-graph by Lemma 5, and hence that at most $rk \leq f$ processors fail in $\mathcal{G}$.

Remember that we always write the vertices of a primitive simplex in a canonical order $y_0, \ldots, y_k$. In the same way, we always write the graphs labeling the vertices] of a primitive simplex in the canonical order $\mathcal{G}_0, \ldots, \mathcal{G}_k$, where $\mathcal{G}_i$ is the graph labeling y_i.

6.6 Graphs on a simplex

The graphs labeling the vertices of a primitive simplex have some convenient properties. For this section, fix a primitive simplex S, and let $y_0, \ldots, y_k$ be the vertices of S and let $\mathcal{G}_0, \ldots, \mathcal{G}_k$ be the graphs labeling the corresponding vertices of S. Our first result says that any processor that is uncovered at a vertex of S is nonfaulty at all vertices of S.

Lemma 6: If processor q is not covered in the graph labeling a vertex of S, then q is nonfaulty in the graph labeling every vertex of S.

Proof: Let $y_0 = (a_1, \ldots, a_k)$ be the first vertex of S. For each i, let σ_i and $\sigma_i\tau_i$ be the prefixes of $\sigma[v_i]$ consisting of the first a_i and a_i+1 operations, and let $\mathcal{H}_i$ and $\mathcal{H}'_i$ be the result of applying σ_i and $\sigma_i\tau_i$ to $\mathcal{F}_{i-1}$. For each i, we know that the graph $\mathcal{G}_i$ labeling the vertex y_i of S is the merge of graphs $\mathcal{I}^i_1, \ldots, \mathcal{I}^i_k$ where $\mathcal{I}^i_j$ is either $\mathcal{H}_j$ or $\mathcal{H}'_j$. Suppose q is faulty in $\mathcal{G}_i$. Then q must be faulty in some graph $\mathcal{I}^i_j$ in the sequence of graphs $\mathcal{I}^i_1, \ldots, \mathcal{I}^i_n$ merged to form $\mathcal{G}_i$, so q must fail in one of the graphs $\mathcal{H}_j$ or $\mathcal{H}'_j$. Since σ_j and $\sigma_j\tau_j$ are prefixes of $\sigma[v_j]$, it is easy to see from the definition of $\sigma[v_j]$ that the fact that q fails in one of the graphs $\mathcal{H}_j$ and $\mathcal{H}'_j$ implies that q is covered in both graphs. Since one of these graphs is contained in the sequence of graphs merged to form $\mathcal{G}_a$ for each a, it follows that q is covered in each $\mathcal{G}_a$. This contradicts the fact that q is uncovered in a graph labeling a vertex of S. □

Our next result shows that we can use the bound on the number of tokens to bound the number of processors failing at any vertex of S.

Lemma 7: If F_i is the set of processors failing in $\mathcal{G}_i$ and $F = \cup_i F_i$, then $|F| \leq rk \leq f$.

Proof: If $q \in F$, then $q \in F_i$ for some i and q fails in $\mathcal{G}_i$, so q is covered in every graph labeling every vertex of S by Lemma 6. It follows that each processor in F is covered in each graph labeling S. Since there are at most rk tokens to cover processors in any graph, there are at most rk processors in F. □

We have assigned graphs to S, and now we must assign processors to S. A *local processor labeling* of S is an assignment of distinct processors $q_0, \ldots, q_k$ to the vertices $y_0, \ldots, y_k$ of S so that q_i is uncovered in $\mathcal{G}_i$ for each y_i. A *global processor labeling* of B is an assignment of processors to vertices of B that induces a local processor labeling at each primitive simplex. The final important property of the graphs labeling S is that if we use a processor labeling to label S with processors, then S is consistent with a single global communication graph. The proof of this requires a few preliminary results.

Lemma 8: If $\mathcal{G}_{i-1}$ and $\mathcal{G}_i$ differ in p's input value, then p is silent from round 1 in $\mathcal{G}_0, \ldots, \mathcal{G}_k$. If $\mathcal{G}_{i-1}$ and $\mathcal{G}_i$ differ in the color of an edge from q to p in round t, then p and q are silent from round $t+1$ in $\mathcal{G}_0, \ldots, \mathcal{G}_k$.

Proof: Suppose the two graphs $\mathcal{G}_{i-1}$ and $\mathcal{G}_i$ labeling vertices y_{i-1} and y_i differ in the input to p at time $t = 0$ or in the color of an edge from q to p in round t. The vertices differ in exactly one coordinate j, so $y_{i-1} = (a_1, \ldots, a_j, \ldots, a_k)$ and $y_i = (a_1, \ldots, a_j + 1, \ldots, a_k)$. For each ℓ, let σ_ℓ be the prefix of $\sigma[v_\ell]$ consisting of the first a_ℓ operations, and let $\mathcal{H}^0_\ell$ be the result of applying σ_ℓ to $\mathcal{F}_{\ell-1}$. Furthermore, in the special case of $\ell = j$, let $\sigma_j \tau_j$ be the prefix of $\sigma[v_j]$ consisting of the first $a_j + 1$ operations, and let $\mathcal{H}^1_j$ be the result of applying $\sigma_j \tau_j$ to $\mathcal{F}_{j-1}$.

We know that $\mathcal{G}_{i-1}$ is the merge of $\mathcal{H}^0_1, \ldots, \mathcal{H}^0_j, \ldots, \mathcal{H}^0_k$, and that $\mathcal{G}_i$ is the merge of $\mathcal{H}^0_1, \ldots, \mathcal{H}^1_j, \ldots, \mathcal{H}^0_k$. If $\mathcal{H}^0_j$ and $\mathcal{H}^1_j$ are equal, then $\mathcal{G}_{i-1}$ and $\mathcal{G}_i$ are equal. Thus, $\mathcal{H}^0_j$ and $\mathcal{H}^1_j$ must differ in the input to p at time $t = 0$ or in the color of an edge between q and p in round t, exactly as $\mathcal{G}_{i-1}$ and $\mathcal{G}_i$ differ. Since $\mathcal{H}^0_j$ and $\mathcal{H}^1_j$ are the result of applying σ_j and $\sigma_j \tau_j$ to $\mathcal{F}_{j-1}$, this change at time t must be caused by the operation τ_j. It is easy to see from the definition a graph operation like τ_j that (1) if τ_j changes p's input value, then p is silent from round 1 in $\mathcal{H}^0_j$ and $\mathcal{H}^1_j$, and (2) if τ_j changes the color of an edge from q to p in round t, then p and q are silent from round $t+1$ in $\mathcal{H}^0_j$ and $\mathcal{H}^1_j$. Consequently, the same is true in the merged graphs $\mathcal{G}_{i-1}$ and $\mathcal{G}_i$. $\square$

Lemma 9: If $\mathcal{G}_{i-1}$ and $\mathcal{G}_i$ differ in the local communication graph of p at time t, then p is silent from round $t + 1$ in $\mathcal{G}_0, \ldots, \mathcal{G}_k$.

Proof: We proceed by induction on t. If $t = 0$, then the two graphs must differ in the input to p at time 0, and Lemma 8 implies that p is silent from round 1 in the graphs $\mathcal{G}_0, \ldots, \mathcal{G}_k$ labeling the simplex. Suppose $t > 0$ and the inductive hypothesis holds for $t - 1$. Processor p's local communication graph at time t can differ in the two graphs for one of two reasons: either p hears from some processor q in round t in one graph and not in the other, or p hears from some processor q in both graphs but q has different local communication graphs at time $t - 1$ in the two graphs. In the first case, Lemma 8 implies that p is silent from round $t+1$ in the graphs $\mathcal{G}_0, \ldots, \mathcal{G}_k$. In the second case, the induction hypothesis for $t - 1$ implies that q is silent from round t in the graphs $\mathcal{G}_0, \ldots, \mathcal{G}_k$. In particular, q is silent in round t in $\mathcal{G}_{i-1}$ and $\mathcal{G}_i$, contradicting the assumption that p hears from q in round t in both graphs, so this case can't happen. $\square$

Lemma 10: If p sends a message in round r in any of the graphs $\mathcal{G}_0, \ldots, \mathcal{G}_k$, then p has the same local communication graph at time $r-1$ in all of the graphs $\mathcal{G}_0, \ldots, \mathcal{G}_k$.

Proof: If p has different local communication graphs at time $r-1$ in two of the graphs $\mathcal{G}_0, \ldots, \mathcal{G}_k$, then there are two adjacent graphs $\mathcal{G}_{i-1}$ and $\mathcal{G}_i$ in which p has different local communication graphs at time $r-1$. By Lemma 9, p is silent in round r in all of the graphs $\mathcal{G}_0, \ldots, \mathcal{G}_k$, contradicting the assumption that p sent a round r message in one of them. □

Finally, we can prove the crucial property of primitive simplexes in the Bermuda Triangle:

Lemma 11: Given a local processor labeling, let $q_0, \ldots, q_k$ be the processors labeling the vertices of S, and let $\mathcal{L}_i$ be the local communication graph of q_i in $\mathcal{G}_i$. There is a global communication graph $\mathcal{G}$ with the property that each q_i is nonfaulty in $\mathcal{G}$ and has the local communication graph $\mathcal{L}_i$ in $\mathcal{G}$.

Proof: Let Q be the set of processors that send a round r message in any of the graphs $\mathcal{G}_0, \ldots, \mathcal{G}_k$. Notice that this set includes the uncovered processors $q_0, \ldots, q_k$, since Lemma 6 says that these processors are nonfaulty in each of these graphs. For each processor $q \in Q$, Lemma 10 says that q has the same local communication graph at time $r-1$ in each graph $\mathcal{G}_0, \ldots, \mathcal{G}_k$.

Let $\mathcal{H}$ be the global communication graph underlying any one of these graphs. Notice that each processor $q \in Q$ is active through round $r-1$ in $\mathcal{H}$. To see this, notice that since q sends a message in round r in one of the graphs labeling S, it sends all messages in round $r-1$ in that graph. On the other hand, if q fails to send a message in round $r-1$ in $\mathcal{H}$, then the same is true for the corresponding graph labeling S. Thus, there are adjacent graphs $\mathcal{G}_{i-1}$ and $\mathcal{G}_i$ labeling S where p sends a round $r-1$ message in one and not in the other. Consequently, Lemma 8 says q is silent in round r in all graphs labeling S, but this contradicts the fact that q does send a round r message in one of these graphs.

Now let $\mathcal{G}$ be the global communication graph obtained from $\mathcal{H}$ by coloring green each round r edge from each processor $q \in Q$, unless the edge is red in one of the local communication graphs $\mathcal{L}_0, \ldots, \mathcal{L}_k$ in which case we color it red in $\mathcal{G}$ as well. Notice that since the processors $q \in Q$ are active through round $r-1$ in $\mathcal{H}$, changing the color of a round r edge from a processor $q \in Q$ to either red or green is acceptable, provided we do not cause more that f processors to fail in the process. Fortunately, Lemma 7 implies that there are at least $n - rk \geq n - f$ processors that do not fail in any of the graphs $\mathcal{G}_0, \ldots, \mathcal{G}_k$. This means that there is a set of $n-f$ processors that send to every processor in round r of every graph $\mathcal{G}_i$, and in particular that the round r edges from these processors are green in every local communication graph $\mathcal{L}_i$. It follows that for at least $n-f$ processors, all round r edges from these processors are green in $\mathcal{G}$, so at most f processors fail in $\mathcal{G}$.

Each processor q_i is nonfaulty in $\mathcal{G}$, since q_i is nonfaulty in each $\mathcal{G}_0, \ldots, \mathcal{G}_k$, meaning each edge from q_i is green in each $\mathcal{G}_0, \ldots, \mathcal{G}_k$ and $\mathcal{L}_0, \ldots, \mathcal{L}_k$, and therefore in $\mathcal{G}$. In addition, each processor q_i has the local communication graph $\mathcal{L}_i$ in $\mathcal{G}$. To see this, notice that $\mathcal{L}_i$ consists of a round r edge from p_j to q_i for each j, and the local communication graph for p_j at time $r-1$ if this edge is green. This edge is green in $\mathcal{L}_i$

if and only if it is green in $\mathcal{G}$. In addition, if this edge is green in $\mathcal{L}_i$, then it is green in $\mathcal{G}_i$. In this case, Lemma 10 says that p_j has the same local communication graph at time $r-1$ in each graph $\mathcal{G}_0, \ldots, \mathcal{G}_k$, and therefore in $\mathcal{G}$. Consequently, q_i has the local communication graph $\mathcal{L}_i$ in $\mathcal{G}$. □

7 Processor Assignment

What Lemma 11 at the end of the preceding section tells us is that all we have left to do is to construct a global processor labeling. In this section, we show how to do this. We first associate a set of "live" processors with each communication graph labeling a vertex of B, and then we choose one processor from each set to label vertices of B.

7.1 Live processors

Given a graph $\mathcal{G}$, we construct a set of $c = n - rk \geq k + 1$ uncovered (and hence nonfaulty) processors. We refer to these processors as the *live* processors in $\mathcal{G}$, and we denote this set by *live*$(\mathcal{G})$. These live sets have one crucial property: if $\mathcal{G}$ and $\mathcal{G}'$ are two graphs labeling adjacent vertices, and if p is in both *live*$(\mathcal{G})$ and *live*$(\mathcal{G}')$, then p has the same rank in both sets. As usual, we define the *rank* of p_i in a set R of processors to be the number of processors $p_j \in R$ with $j \leq i$.

Given a graph $\mathcal{G}$, we now show how to construct *live*$(\mathcal{G})$. This construction has one goal: if $\mathcal{G}$ and $\mathcal{G}'$ are graphs labeling adjacent vertices, then the construction should minimize the number of processors whose rank differs in the sets *live*$(\mathcal{G})$ and *live*$(\mathcal{G}')$. The construction of *live*$(\mathcal{G})$ begins with the set of all processors, and removes a set of rk processors, one for each token. This set of removed processors includes the covered processors, but may include other processors as well. For example, suppose p_i and p_{i+1} are covered with one token each in $\mathcal{G}$, but suppose p_i is uncovered and p_{i+1} is covered by two tokens in $\mathcal{G}'$. For simplicity, let's assume these are the only tokens on the graphs. When constructing the set *live*$(\mathcal{G})$, we remove both p_i and p_{i+1} since they are both covered. When constructing the set *live*$(\mathcal{G}')$, we remove p_{i+1}, but we must also remove a second processor corresponding to the second token covering p_{i+1}. Which processor should we remove? If we choose a low processor like p_1, then we have changed the rank of a low processor like p_2 from 2 to 1. If we choose a high processor like p_n, then we have change the rank of a high processor like p_{n-1} from $n-3$ to $n-2$. On the other hand, if we choose to remove p_i again, then no processors change rank. In general, the construction of *live*$(\mathcal{G})$ considers each processor p in turn. If p is covered by m_p tokens in $\mathcal{G}$, then the construction removes m_p processors by starting with p, working down the list of remaining processors smaller than p, and then working up the list of processors larger than p if necessary.

Specifically, given a graph $\mathcal{G}$, the *multiplicity* of p is the number m_p of tokens appearing on nodes for p in $\mathcal{G}$, and the *multiplicity* of $\mathcal{G}$ is the vector $m = \langle m_{p_1}, \ldots, m_{p_n} \rangle$. Given the multiplicity of $\mathcal{G}$ as input, the algorithm given in Figure 12 computes *live*$(\mathcal{G})$. In this algorithm, processor p_i is denoted by its index i. We refer to the ith iteration of the main loop as the ith step of the construction. This construction has two obvious properties:

$S \leftarrow \{1, \ldots, n\}$
for each $i = 1, \ldots, n$
  count $\leftarrow 0$
  for each $j = i, i-1, \ldots, 1, i+1, \ldots, n$
    if count = m_i then break
    if $j \in S$ then
      $S \leftarrow S - \{j\}$
      count $\leftarrow$ count + 1
live$(\mathcal{G}) \leftarrow S$

Figure 12: The construction of *live*$(\mathcal{G})$.

Lemma 12: If $i \in$ *live*$(\mathcal{G})$ then

1. *i is uncovered*: $m_i = 0$
2. *room exists under i*: $\sum_{j=1}^{i-1} m_j \leq i - 1$

Proof: Suppose $i \in$ *live*$(\mathcal{G})$. For part 1, if $m_i > 0$ then i will be removed by step i if it has not already removed by an earlier step, contradicting $i \in$ *live*$(\mathcal{G})$. For part 2, notice that steps 1 through $i-1$ remove a total of $\sum_{j=1}^{i-1} m_j$ values. If this sum is greater than $i-1$, then it is not possible for all of these values to be contained in $1, \ldots, i-1$, so i will be removed within the first $i-1$ steps, contradicting $i \in$ *live*$(\mathcal{G})$. □

The assignment of graphs to the corners of a simplex has the property that once p becomes covered on one corner of S, it remains covered on the following corners of S:

Lemma 13: If p is uncovered in the graphs $\mathcal{G}_i$ and $\mathcal{G}_j$, where $i < j$, then p is uncovered in each graph $\mathcal{G}_i, \mathcal{G}_{i+1}, \ldots, \mathcal{G}_j$.

Proof: If p is covered in $\mathcal{G}_\ell$ for some ℓ between i and j, then p is uncovered in $\mathcal{G}_{\ell-1}$ and covered in $\mathcal{G}_\ell$ for some ℓ between i and j. Since $\mathcal{G}_{\ell-1}$ and $\mathcal{G}_\ell$ are on adjacent vertices of the simplex, the sequences of graphs merged to construct them are of the form $\mathcal{H}_1, \ldots, \mathcal{H}_m, \ldots, \mathcal{H}_k$ and $\mathcal{H}_1, \ldots, \mathcal{H}'_m, \ldots, \mathcal{H}_k$, respectively, for some m. Since p is uncovered in $\mathcal{G}_{\ell-1}$ and covered in $\mathcal{G}_\ell$, it must be that p is uncovered in $\mathcal{H}_m$ and covered in $\mathcal{H}'_m$. Notice, however, that $\mathcal{H}'$ is used in the construction of each graph $\mathcal{G}_\ell, \mathcal{G}_{\ell+1}, \ldots, \mathcal{G}_j$. This means that p is covered in each of these graphs, contradicting the fact that p is uncovered in $\mathcal{G}_j$. □

Finally, because token placements in adjacent graphs on a simplex differ in at most the movement of one token from one processor to an adjacent processor, we can use the preceding lemma to prove the following:

Lemma 14: If $p \in$ *live*$(\mathcal{G}_i)$ and $p \in$ *live*$(\mathcal{G}_j)$, then p has the same rank in *live*$(\mathcal{G}_i)$ and *live*$(\mathcal{G}_j)$.

Proof: Assume without loss of generality that $i < j$. Since $p \in$ *live*$(\mathcal{G}_i)$ and $p \in$ *live*$(\mathcal{G}_j)$, Lemma 12 implies that p is uncovered in the graphs $\mathcal{G}_i$ and $\mathcal{G}_j$, and Lemma 13 implies that p is uncovered in each graph $\mathcal{G}_i, \mathcal{G}_{i+1}, \ldots, \mathcal{G}_j$. Since token placements in adjacent graphs differ in at most the movement of one token from one processor to an adjacent processor, and since p is uncovered in all of these graphs, this means that the number of tokens on processors smaller than p is the same in all of these graphs. Specifically, the sum $\sum_{\ell=1}^{p-1} m_\ell$ of multiplicities of processors smaller than p is the same in $\mathcal{G}_i, \mathcal{G}_{i+1}, \ldots, \mathcal{G}_j$. In particular, Lemma 12 implies that this sum is the same value $s \leq p-1$ in $\mathcal{G}_i$ and $\mathcal{G}_j$, so p has the same rank $p-s$ in *live*$(\mathcal{G}_i)$ and *live*$(\mathcal{G}_j)$. □

7.2 Processor labeling

We now choose one processor from each set *live*$(\mathcal{G})$ to label the vertex with graph $\mathcal{G}$. Given a vertex $x = (x_1, \ldots, x_k)$, we define

$$\mathit{plane}(x) = \sum_{i=1}^{k} x_i \pmod{k+1}$$

.

Lemma 15: If x and y are distinct vertices of the same simplex, then *plane*$(x) \neq$ *plane*(y).

Proof: Since x and y are in the same simplex, we can write $y = x + f_1 + \cdots + f_j$ for some distinct unit vectors $f_1, \ldots, f_j$ and some $1 \leq j \leq k$. If $x = (x_1, \ldots, x_k)$ and $y = (y_1, \ldots, y_k)$, then the sums $\sum_{i=1}^{k} x_i$ and $\sum_{i=1}^{k} y_i$ differ by exactly j. Since $1 \leq j \leq k$ and since planes are defined as sums modulo $k+1$, we have *plane*$(x) \neq$ *plane*(y). □

We define a global processor labeling π as follows: given a vertex x labeled with a graph $\mathcal{G}$, we define π to map x to the processor having rank *plane*(x) in *live*$(\mathcal{G})$.

Lemma 16: The mapping π is a global processor labeling.

Proof: First, it is clear that π maps each vertex x labeled with a graph $\mathcal{G}_x$ to a processor q_x that is uncovered in $\mathcal{G}_x$. Second, π maps distinct vertices of a simplex to distinct processors. To see this, suppose that both x and y are labeled with p, and let $\mathcal{G}_x$ and $\mathcal{G}_y$ be the graphs labeling x and y. We know that the rank of p in *live*$(\mathcal{G}_x)$ is *plane*(x) and that the rank of p in *live*$(\mathcal{G}_y)$ is *plane*(y), and we know that p has the same rank in *live*$(\mathcal{G}_x)$ and *live*$(\mathcal{G}_y)$ by Lemma 14. Consequently, *plane*(x) = *plane*(y), contradicting Lemma 15. □

We label the vertices of B with processors according to the processor labeling π.

8 Ordered Pair Assignment

Finally, we assign ordered pairs $(p, \mathcal{L})$ of processor ids and local communication graphs to vertices of B. Given a vertex x labeled with processor p and graph $\mathcal{G}$, we label x

with the ordered pair $(p, \mathcal{L})$ where $\mathcal{L}$ is the local communication graph of p in $\mathcal{G}$. The following result is a direct consequence of Lemmas 11 and 16. It says that the local communication graphs of processors labeling the corners of a vertex are consistent with a single global communication graph.

Lemma 17: Let $q_0, \ldots, q_k$ and $\mathcal{L}_0, \ldots, \mathcal{L}_k$ be the processors and local communication graphs labeling the vertices of a simplex. There is a global communication graph $\mathcal{G}$ with the property that each q_i is nonfaulty in $\mathcal{G}$ and has the local communication graph $\mathcal{L}_i$ in $\mathcal{G}$.

9 Sperner's Lemma

We now state Sperner's Lemma, and use it to prove a lower bound on the number of rounds required to solve k-set agreement.

Notice that the corners of B are points of the form $c_i = (N, \ldots, N, 0, \ldots, 0)$ with i indices of value N for $0 \leq i \leq k$. For example, $c_0 = (0, \ldots, 0)$, $c_1 = (N, 0, \ldots, 0)$, and $c_k = (N, \ldots, N)$. Informally, a Sperner coloring of B assigns a color to each vertex so that each corner vertex c_i is given a distinct color w_i, each vertex on the edge between c_i and c_j is given either w_i or w_j, and so on.

More formally, let S be a simplex and let F be a face of S. Any triangulation of S induces a triangulation of F in the obvious way. Let T be a triangulation of S. A *Sperner coloring* of T assigns a color to each vertex of T so that each corner of T has a distinct color, and so that the vertices contained in a face F are colored with the colors on the corners of F, for each face F of T. Sperner colorings have a remarkable property: at least one simplex in the triangulation must be given all possible colors.

Lemma 18 (Sperner's Lemma): If B is a triangulation of a k-simplex, then for any Sperner coloring of B, there exists at least one k-simplex in B whose vertices are all given distinct colors.

Let P be the protocol whose existence we assumed in the previous section. Define a coloring χ_P of B as follows. Given a vertex x labeled with processor p and local communication graph $\mathcal{L}$, color x with the value v that P requires processor p to choose when its local communication graph is $\mathcal{L}$. This coloring is clearly well-defined, since P is a protocol in which all processors chose an output value at the end of round r. We will now expand the argument sketched in the introduction to show that χ_P is a Sperner coloring.

We first prove a simple claim. Recall that $\mathcal{B}$ is the simplex whose vertices are the corner vertices $c_0, \ldots, c_k$, and that B is a triangulation of $\mathcal{B}$. Let $\mathcal{F}$ be some face of $\mathcal{B}$ *not* containing the corner c_i, and let F denote the triangulation of $\mathcal{F}$ induced by B. We prove the following technical statement about vertices in F.

Claim 19: If $x = (x_1, \ldots, x_k)$ is a vertex of a face F not containing c_i, then

1. if $i = 0$, then $x_1 = N$,
2. if $0 < i < k$, then $x_{i+1} = x_i$, and

3. if $i = k$, then $x_k = 0$.

Proof: Each vertex x of B can be expressed using *barycentric coordinates* with respect to the corner vertices: that is, $x = \alpha_0 c_0 + \cdots + \alpha_k c_k$, where $0 \le \alpha_j \le 1$ for $0 \le j \le k$ and $\sum_{i=0}^{k} \alpha_i = 1$. Since x is a vertex of a face F not containing the corner c_i, it follows that $\alpha_i = 0$. We consider the three cases.
Case 1: $i = 0$. Each corner $c_1, \ldots, c_k$ has the value N in the first position. Since $\alpha_0 = 0$, the value in the first position of $\alpha_0 c_0 + \cdots + \alpha_k c_k$ is $(\alpha_1 + \cdots + \alpha_k)N = N$.
Case 2: $0 < i < k$. Each corner $c_0, \ldots, c_{i-1}$ has 0 in positions i and $i+1$, and each corner $c_{i+1}, \ldots, c_k$ has N in positions i and $i+1$. Since $\alpha_i = 0$, the linear combination $\alpha_0 c_0 + \cdots + \alpha_k c_k$ will have the same value $(\alpha_{i+1} + \cdots + \alpha_k)N$ in positions i and $i+1$. Thus, $x_i = x_{i+1}$.
Case 3: $i = k$. Each corner $c_0, \ldots, c_{k-1}$ has 0 in position k. Since $\alpha_k = 0$, the value in the kth position of $\alpha_0 c_0 + \cdots + \alpha_k c_k$ is 0. Thus, $x_k = 0$. □

Lemma 20: If P is a protocol for k-set agreement tolerating f faults and halting in $r \le \lfloor f/k \rfloor$ rounds, then χ_P is a Sperner coloring of B.

Proof: We must show that χ_P satisfies the two conditions of a Sperner coloring.

For the first condition, consider any corner vertex c_i. Remember that c_i was originally labeled with the 1-graph $\mathcal{F}_i$ describing a failure-free execution in which all processors start with input v_i, and that the local communication graph $\mathcal{L}$ labeling c_i is a subgraph of $\mathcal{F}_i$. Since the k-set agreement problem requires that any value chosen by a processor must be an input value of some processor, all processors must chose v_i in $\mathcal{F}_i$, and it follows that the vertex c_i must be colored with v_i. This means that each corner c_i is colored with a distinct value v_i.

For the second condition, consider any face F of B, and let us prove that vertices in F are colored with the colors on the corners of F. Equivalently, suppose that c_i is not a corner of F, and let us prove that no vertex in F is colored with v_i.

Consider the global communication graph $\mathcal{G}$ originally labeling x, and the graphs $\mathcal{H}_1, \ldots, \mathcal{H}_k$ used in the merge defining $\mathcal{G}$. The definition of this merge says that the input value labeling a node $\langle p, 0 \rangle$ in $\mathcal{G}$ is v_m where m is the maximum m such that $\langle p, 0 \rangle$ is labeled with v_m in $\mathcal{H}_m$, or v_0 if no such m exists. Again, we consider three cases. In each case, we show that no processor in $\mathcal{G}$ has the input value v_i.

Suppose $i = 0$. Since $x_1 = N$ by Claim 19, we know that $\mathcal{H}_1 = \mathcal{F}_1$, where the input value of every processor is v_1. By the definition of the merge operation, it follows immediately that no processor in $\mathcal{G}$ can have input value v_0.

Suppose $1 < i < k$. Again, $x_{i+1} = x_i$ by Claim 19. Now, $\mathcal{H}_i$ is the result of applying σ_i, the first x_i operations of $\sigma[v_i]$, to the graph $\mathcal{F}_{i-1}$. Similarly, $\mathcal{H}_{i+1}$ is the result of applying σ_{i+1}, the first x_{i+1} operations of $\sigma[v_{i+1}]$, to the graph $\mathcal{F}_i$. Since $x_{i+1} = x_i$, both σ_i and σ_{i+1} are of the same length, and it follows that σ_i contains an operation of the form *change*(p, v_i) if and only if σ_{i+1} contains an operation of the form *change*(p, v_{i+1}). This implies that for any processor, either its input value is v_{i-1} in $\mathcal{H}_i$ and v_i in $\mathcal{H}_{i+1}$, or its input value is v_i in $\mathcal{H}_i$ and v_{i+1} in $\mathcal{H}_{i+1}$. In both cases, v_i is not the input value of this processor.

Suppose $i = k$. Since $x_k = 0$ by Claim 19, we know that $\mathcal{H}_k = \mathcal{F}_{k-1}$, where the input value of every processor is v_{k-1}. By the definition of *merge*, it follows immediately that no processor in $\mathcal{G}$ can have input value v_k.

Therefore, we have shown that if x is a vertex of a face F of B, and c_i is not a corner vertex of F, then the communication graph $\mathcal{G}$ corresponding to x contains no processor with input value v_i. Therefore, by the agreement condition, the value chosen at this vertex cannot be v_i, and it follows that x is assigned a color other than v_i. So, x must be colored by a color v_j such that c_j is a corner vertex of F. Since c_j is colored v_j, the second condition of Sperner's Lemma holds. So χ_P is a Sperner coloring. $\square$

Sperner's Lemma guarantees that some primitive simplex is colored by $k + 1$ distinct values, and this simplex corresponds to a global state in which $k + 1$ processors choose $k + 1$ distinct values, contradicting the definition of k-set agreement:

Theorem 21: If $n \geq f + k + 1$, then no protocol for k-set agreement can halt in fewer than $\lfloor f/k \rfloor + 1$ rounds.

Proof: Suppose P is a protocol for k-set agreement tolerating f faults and halting in $r \leq \lfloor f/k \rfloor$ rounds, and consider the corresponding Bermuda Triangle B. Lemma 20 says that χ_P is a Sperner coloring of B, so Sperner's Lemma 18 says that there is a simplex S whose vertices are colored with $k + 1$ distinct values $v_0, \ldots, v_k$. Let $q_0, \ldots, q_k$ and $\mathcal{L}_0, \ldots, \mathcal{L}_k$ be the processors and local communication graphs labeling the corners of S. By Lemma 17, there exists a communication graph $\mathcal{G}$ in which q_i is nonfaulty and has local communication graph $\mathcal{L}_i$. This means that $\mathcal{G}$ is a time r global communication graph of P in which each q_i must choose the value v_i. In other words, $k + 1$ processors must choose $k + 1$ distinct values, contradicting the fact that P solves k-set agreement in r rounds. $\square$

10 Protocol

An optimal protocol P for k-set agreement is given in Figure 13. In this protocol, processors repeatedly broadcast input values and keep track of the least input value received in a local variable *best*. Initially, a processor sets *best* to its own input value. In each of the next $\lfloor f/k \rfloor + 1$ rounds, the processor broadcasts the value of *best* and then sets *best* to the smallest value received in that round from any processor (including itself). In the end, it chooses the value of *best* as its output value.

To prove that P is an optimal protocol, we must prove that, in every execution of P, processors halt in $r = \lfloor f/k \rfloor + 1$ rounds, every processor's output value is some processor's input value, and the set of output values chosen has size at most k. The first two statements follow immediately from the text of the protocol, so we need only prove the third. For each time t and processor p, let $best_{p,t}$ be the value of *best* held by p at time t. For each time t, let $Best(t)$ be the set of values $best_{q_1,t}, \ldots, best_{q_\ell,t}$ where the processors $q_1, \ldots, q_\ell$ are the processors active through time t. Notice that $Best(0)$ is the set of input values, and that $Best(r)$ is the set of chosen output values. Our first observation is that the set $Best(t)$ never increases from one round to the next.

Lemma 22: $Best(t) \supseteq Best(t + 1)$ for all times t.

$best \leftarrow input\_value$;

for each round 1 through $\lfloor f/k \rfloor + 1$ do
  broadcast $best$;
  receive values $b_1, \ldots, b_\ell$ from other processors;
  $best \leftarrow \min\{b_1, \ldots, b_\ell\}$;

choose $best$.

Figure 13: An optimal protocol P for k-set agreement.

Proof: If $b \in Best(t+1)$, then $b = best_{p,t+1}$ for some processor p active through round $t+1$. Since $best_{p,t+1}$ is the minimum of the values $b_1, \ldots, b_\ell$ sent to p by processors during round $t+1$, we know that $b = best_{q,t}$ for some processor q that is active through round t. Consequently, $b \in Best(t)$. □

We can use this observation to prove that the only executions in which many output values are chosen are executions in which many processors fail. We say that a processor *p fails before time t* if there is a processor q to which p sends no message in round t (and p may fail to send to q in earlier rounds as well).

Lemma 23: If $|Best(t)| = d+1$, then at least dt processors fail before time t.

Proof: We proceed by induction on t. The case of $t = 0$ is immediate, so suppose that $t > 0$ and that the induction hypothesis holds for $t-1$. Since $|Best(t)| = d+1$ and since $Best(t) \subseteq Best(t-1)$ by Lemma 22, it follows that $|Best(t-1)| \geq d+1$, and the induction hypothesis for $t-1$ implies that there is a set S of $d(t-1)$ processors that fail before time $t-1$. It is enough to show that there are an additional d processors not contained in S that fail before time t.

Let $b_0, \ldots, b_d$ be the values of $Best(t)$ written in increasing order. Let q be a processor with $best_{q,t}$ set to the largest value b_d at time t, and for each value b_i let q_i be a processor that sent b_i in round $t-1$. The processors $q_0, \ldots, q_d$ are distinct since the values $b_0, \ldots, b_d$ are distinct, and these processors do not fail before time $t-1$ since they send a message in round t, so they are not contained in S. On the other hand, the processors $q_0, \ldots, q_{d-1}$ sending the small values $b_0, \ldots, b_{d-1}$ in round $t-1$ clearly did not send their values to the processor q setting $best_{q,t}$ to the large value b_d, or q would have set $best_{q,t}$ to a smaller value. Consequently, these d processors $q_0, \ldots, q_{d-1}$ fail in round t and hence fail before time t. □

Since $Best(r)$ is the set of output values chosen by processors at the end of round $r = \lfloor f/k \rfloor + 1$, if $k+1$ output values are chosen, then Lemma 23 says that at least kr processors fail, which is impossible since $f < kr$. Consequently, the set of output values chosen has size at most k, as we are done.

Theorem 24: The protocol P solves k-set agreement in $\lfloor f/k \rfloor + 1$ rounds.

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