THE ISIS PROJECT:
Fault-Tolerance in Large Distributed Systems

Semi-annual R&D Status Report

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607-355-9199

July 1, 1990

This work was sponsored by the Defense Advanced Research Projects Agency (DoD), under contract NAG 2-593 issued by the NASA Ames Research Center, Moffett Field, CA.

The view, opinions and findings contained in this report are those of the authors and should not be construed as an official DoD position, policy, or decision.
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The ISIS Project

This semi-annual status report covers activities of the ISIS project during the second half of 1989. Because this is our second progress report under NASA funding, we assume that the reader has some background regarding the goals and status of our effort, and focus instead on technical accomplishments during the report period and goals for the next six months. In addition to listing our most recent publications, we also cite several more general publications at the end of this report. Readers unfamiliar with our work should probably start by reading our previous semi-annual report (January 1990) and perhaps some of these cited papers.

Goals of the ISIS Effort During the Report Period

During the first six months of 1990, our project had several independent objectives:

1. At the level of the ISIS Toolkit, we undertook to complete ISIS release V2.0, containing our "bypass" communication protocols [1]. This effort was successful, and V2.0 has been released with a preliminary implementation of the bypass protocols. Performance of the system is greatly enhanced by this change, but the initial software release is limited in some respects. We are excited by this progress, and have now found several additional ways to extend the protocol suite in the most common client-server settings that arise under ISIS. We plan to make these changes to our system during the coming six months. With these changes, the bypass mechanisms will have largely replaced all other protocol implementations in ISIS, and the system should be both more scalable and capable of accommodating communication transport protocols based on hardware (ethernet and FDDI) multicast. With respect to system releases, we are close to a V2.1 release consisting of V2.0 with bug fixes, but are undecided as to when these other extensions would be released to the public.

2. The Meta project focused on the definition of the Lomita programming language during the report period. Lomita is a high level language for specifying rules that monitor sensors for conditions of interest and triggering appropriate reactions. This design has now been completed, and implementation of Lomita is underway on the Meta 2.0 platform (this defines a database interface to system instrumentation and is already operational).

3. The Deceit file system effort completed a prototype, which is now operational at Cornell. Our current plans are to make Deceit available on a
limited basis, primarily for use in two hospital information systems with which we are collaborating.

4. A long-haul communication subsystem project was completed and can now be used as part of ISIS. This effort resulted in tools for linking ISIS systems on different LANs together over long-haul communications lines.

5. Several papers were completed during the report period and are described below.

6. Visitor Robbert Van Renesse (a recent graduate of the Ameoba project of Vrije University, Amsterdam) developed Magic Lantern, a graphical tool for building application monitoring and control interfaces. Magic is being included as part of our general ISIS releases. Van Renesse is now spending six months at Bell Laboratories, but we are hoping that he will subsequently return to our group on a permanent basis.

Goals of the ISIS Effort During the Remainder of 1990

Looking to the remaining six months of 1990, our project has identified the following objectives:

1. ISIS Toolkit: This group sees a near-term need to implement the so-called "pg_client" extensions to the bypass mode protocols and to complete the implementation of our hierarchical process-group tools.

2. We have become interesting in developing a general purpose ISIS-based system resource manager. Such a tool would schedule tasks onto machines in a network setting, providing a low-level programmable interface that could be specialized for a wide variety of uses. Tools of this sort are easily built under ISIS and we see little difficulty in this undertaking, which should be of substantial value within the ISIS user community.

3. We are starting work on the design of a new version of the ISIS system, ISIS++, which would run under the Mach and Chorus kernels. Goals would be increased transparency, the ability to support a realtime toolkit, improved object-oriented interfaces, and the possibility of exploiting shared memory and higher performance multicast facilities. The current UNIX-based toolkit would still be supported on top of ISIS++.

4. We hope to achieve a working version of Lomita during this period.

5. We expect to see a working "production" version of Deceit completed during this period and in use at some number of Beta-test sites.
6. We hoping to complete papers on several topics: Experience with the ISIS Toolkit, Part I: Issues of scope and scale (Birman and Cooper), Experience with the ISIS Toolkit, Part II: Programming with process groups (Birman and Cooper), Theory of failure detection (Birman and Ricciardi), Long-haul communication programming methodology (Birman, Makpan- gou, Stephenson).

7. An effort to develop a distributed version of ML over ISIS has attracted the interest of Robert Cooper, and we are actively involved in this effort. Our hope is that it may yield insight into better programming tools for use within the system.

Publications

Appendix A contains a copy of our papers on Distributed Application Management and on the BYPASS communications protocols. As noted above, a substantial number of papers are now in the pipeline, including one on scaling through the use of hierarchical mechanisms throughout ISIS, and another on our experience programming with process groups. New topics arise constantly. We expect to release a number of these in technical report form during the fall of 1990.

The following papers give information on ISIS and META for readers unfamiliar with our work. Reprints are available from Cornell. Our project has released a much larger number of papers. E-mail reprint requests to "isis@cs.cornell.edu" or contact the project administrative aide at 607-255-9198.

1. ISIS USERS MANUAL. Kenneth P. Birman, ed Cornell University Department of Computer Science.

   This programmers manual discusses the interface presented to ISIS users who program in C, Lisp or Fortran. The current version of the manual covers ISIS V1.8; an extensive revision version is planned for March 1990, and will include discussion of architectural issues that arise in mapping large applications to the ISIS system.


   This textbook was compiled from the lecture notes used in Arctic 88 and Fingerlakes 89, advanced courses in distributed computing. Two chapters cover the ISIS approach in considerable detail and represent a good technical introduction to our work. The conclusions chapter may also be of interest to readers; it explores the general question of robustness in distributed systems.
The following list identifies publications and papers that will be released in the near future. All of these papers are intended for eventual publication in journals or conferences.


   A new scheme is presented for efficiently implementing a reliable, causally ordered multicast primitive. Intended for use in the ISIS toolkit, it offers a way to bypass the most costly aspects of ISIS while benefiting from virtual synchrony. The facility scales with bounded overhead. Performance is extremely good over a range of reliability and delivery ordering properties; with these new protocols, an application pays for the properties it needs. Moreover, users can plug in new protocols and benefit from them in the context of the remainder of the ISIS runtime environment. Speedups of more than an order of magnitude were obtained when the scheme was implemented within Isis. One conclusion is that systems built using ISIS can achieve performance competitive with the best existing multicast facilities – a finding contradicting the widespread concern that fault-tolerance may be unacceptably costly. All protocols described in the paper have been implemented and instrumented, and the code is available in ISIS V2.0.


   The ISIS Project has developed a distributed programming toolkit and a collection of higher level applications based on these tools. ISIS is now in use at more than 300 locations world-wide. Here, we discuss the lessons (and surprises) gained from this experience with the real world.


   This paper describes MTP: a reliable transport protocol that utilizes the multicast strategy of applicable lower layer network architectures. In addition to transporting data reliably and efficiently, MTP provides the client synchronization necessary for agreement on the receipt of data and the joining of the group of communicants.


   Formal methods for specifying fault-tolerant requirements are extremely important for proving a given application correct and robust. Formal
methods provide a clear and specific description and can lead to a better understanding of the problem. Moreover, a formalism enables one to quantify problems and solutions and compare their strengths. By quantifying both, minimal solutions can be fitted to problems. We believe modal logics, derived from the model of computation, provide an accessible and useful means of specifying and reasoning about fault-tolerant requirements for asynchronous systems. We developed a tense-epistemic logic based on a model of computation whose basic semantic entities are the consistent cuts of an asynchronous run. We use it to specify the safety and liveness properties of the Group Membership Problem in asynchronous systems that can experience crash failures. Phrased in terms of knowledge, these conditions quantify solutions in that there is a correspondence between message complexity and the level of knowledge attained. We use this correspondence to derive two solutions to the Group Membership Problem.


It is common to construct software systems by interconnecting non-distributed components, using remote procedure calls and stream communication channels. This paper examines the problem of distributed application management as it arises in systems having this structure. Our discussion is based on the META system: a collection of utilities and tools for constructing distributed application management software. Built using the ISIS toolkit, these include facilities for monitoring and scheduling activity in an underlying system, for dynamically reconfiguring in response to failures or load changes and for automatically restarting failed system components. A key facility is a sensor/actuator interface supporting a fault-tolerant database of realtime sensor values that can be interrogated using a realtime interval-logic query language. The set of sensors is completely extensible and may include such values as machine load, temperature readings from a thermometer, and even values of dynamically updated variables in the memory space of a user-process. The META system is available from Cornell University as part of its ISIS system distribution.


This technical report describes the new ISIS facilities for interconnecting systems running on physically separated local area networks. The facility assumes that links will normally be down and are open only periodically; it spools communication automatically and transmits in bursts when the opportunity arises. By examining the needs of typical wide area applications (scheduling and a replicated directory) and argument is made that these
facilities will be adequate for solving a wide variety of wide-area problems. The facility has been fully implemented and instrumented; it is included as part of ISIS V2.0.


ISIS and META are two distributed systems projects at Cornell University. The ISIS project, led by Ken Birman, has developed a new methodology, virtual synchrony, for writing robust distributed software.


A reactive system is characterized by a control program that interacts with an environment (or controlled program). The control program monitors the environment and reacts to significant events by sending commands to the environment. This structure is quite general. Not only are most embedded real-time systems reactive systems, but so are monitoring and debugging systems and distributed application management systems. Since reactive systems are usually long-running and may control physical equipment, fault-tolerance is vital. Our research tries to understand the principal issues of fault-tolerance in real-time reactive systems and to build tools that allow a programmer to design reliable, real-time reactive systems.


A priority inversion occurs when a low-priority task causes execution of a higher-priority task to be delayed. The possibility of priority inversion complicates the analysis of systems that use priority-based schedulers because priority inversions invalidate the assumption that a task can be delayed only by higher-priority tasks. This paper formalizes priority inversion and gives sufficient conditions as well as some new protocols for preventing priority inversions.
ISIS Publications List


ISIS Publications List Continued


89-1014 An Advanced Course on Distributed Systems Lecture notes from Attic '88, an advanced course on distributed systems. Addison-Wesley, 1989. To be published.

Supporting Large Scale Applications on Networks of Workstations, Cooper and Birman, April 1989.


Budget

A budgetary summary for the report period has been submitted directly to Maj. Boesch using the DARPA electronic reporting format, with a copy to Jerry Yan at NASA Ames. Expenditures are in line with projections under our current operating budget.
Appendix A
Fast Causal Multicast

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Abstract

A new protocol is presented that efficiently implements a reliable, causally ordered multicast primitive and is easily extended into a totally ordered one. Intended for use in the ISIS toolkit, it offers a way to bypass the most costly aspects of ISIS while benefiting from virtual synchrony. The facility scales with bounded overhead. Measured speedups of more than an order of magnitude were obtained when the protocol was implemented within ISIS. One conclusion is that systems such as ISIS can achieve performance competitive with the best existing multicast facilities - a finding contradicting the widespread concern that fault-tolerance may be unacceptably costly.

Keywords and phrases: Distributed computing, fault-tolerance, process groups, reliable multicast, ABCAST, CBCAST, ISIS.

TR90-1105

This work was supported by the Defense Advanced Research Projects Agency (DoD) under ARPA order 6037, Contract N00140-87-C-8904 and under DARPA/NASA subcontract NAG2-593 administered by the NASA Ames Research Center. The views, opinions, and findings contained in this report are those of the authors and should not be construed as an official Department of Defense position, policy, or decision.
1 Introduction

The Isis Toolkit [BJKS88] provides a variety of tools for building software in loosely coupled distributed environments. The system has been successful in addressing problems of distributed consistency, cooperative distributed algorithms, and fault-tolerance. At the time of this writing, ISIS was in use at more than 250 locations worldwide.

Two aspects of ISIS are key to its overall approach:

- An implementation of virtually synchronous process groups.
- A collection of atomic multicast protocols with which processes and group members interact with groups.

Although ISIS supports a wide range of multicast protocols, a protocol called CBCAST accounts for the majority of communication in the system; in fact, many of the ISIS tools are little more than invocations of this communication primitive. For example, the ISIS replicated data tool uses a single (asynchronous) CBCAST to perform each update and locking operation; reads require no communication at all. A consequence is that the cost of CBCAST represents the dominant performance bottleneck in the Isis system.

The initial Isis CBCAST protocol was costly in part for structural reasons, and in part because of the protocol used. The implementation was within a protocol server, hence all CBCAST communication was via an indirect path. Independent of the cost of the protocol itself, this indirection was tremendously expensive. With respect to the protocol used, our initial implementation favored generality over specialization, permitting extremely flexible destination addressing, and using a piggybacking mechanism that achieved a desired ordering property but required a garbage collection mechanism. On the other hand, this structure seemed to be the only one capable of supporting a powerful, general set of programming tools like the ones in our toolkit: simpler protocols often simply overlook critical forms of functionality, which may explain why so few have entered widespread use. Particularly valuable to us has been the ability to support multiple, possibly overlapping process groups, and virtual synchrony [BJKS88].

The protocol we present here is based on a causal ordering protocol originally developed by Schiper [SES89]. Unlike our previous work, it assumes a preexisting virtually synchronous programming environment like the one that Isis provides, although using few of its features. Further, it supports a relatively restricted form of multicast addressing. Were our work done outside of the context of Isis, this would seriously limit its generality. In our implementation, however, messages that do not conform to these restrictions are simply routed via the old, more costly algorithm. A highly optimized multicast protocol results that bypasses the old Isis system and imposes very little overhead beyond that of the message transport layer. The majority of Isis communication satisfies the requirements of the bypass protocols and hence benefits from our work.

Our protocol uses a timestamping scheme, and in this respect resembles prior work by
Ladkin [LL86] and Peterson [PBS89]. However, our results are substantially more general. The most important differences are these:

- Peterson's Psync-based protocol can be used only in systems composed of a single process group, ours supports multiple, possibly overlapping process groups.

- Both Peterson's and Ladkin's protocols have overhead linear in the number of processes that ever participated in the application, which could be large; our overhead is bounded and small.

Like Peterson's and Ladkin's protocols, our basic protocol provides for message delivery ordering that respects causality in the sender (CBCAST), but is readily extended into a more costly protocol that provides a total delivery ordering even for concurrent invocations (ABCasts).

The bypass protocol suite lets users select the multicast properties desired for an application. Choices include a "raw" delivery service achieving extremely high performance but with minimal reliability guarantees, multicast with atomicity and FIFO delivery, and causal or total ordering. This approach permits the user to pay for just those reliability and ordering properties needed by the application.

The paper is structured as follows. Section 2 reviews the multicasting problem and defines our terminology. Sections 3 and 4 introduce our new technique. Section 5 discusses extensions of the CBCAST protocol, including the bypass ABCAST protocol. The costs of our various primitives are measured in Section 6.

2 Execution model

2.1 Basic system model

The system is composed of processes $P = \{p_1, p_2, \ldots, p_n\}$ with disjoint memory spaces. Initially, we assume that this set is static and known in advance; later we relax this assumption. Processes fail by crashing detectably (a fail-stop assumption); notification is provided by Isis in a manner described below. In many situations, processes will need to cooperate. For this purpose, they form process groups. Each such group has a name and a set of member processes; members join and leave dynamically; a failure causes a departure from all groups to which a process belongs. The members of a process group need not be identical, nor is there any limit on the number of groups to which a process may belong. The set of groups is denoted by $G = \{g_1, g_2, \ldots\}$. In typical settings, the number of groups will be large and processes will belong to several groups.

Our system model is unusual in assuming an external service that implements the process group abstraction. The interface from a process to this service will not concern us here, but the manner in which the service communicates to a process is highly relevant.
A view of a process group is a list of its members. A view sequence for $g$ is a list $\text{view}_0(g), \text{view}_1(g), ..., \text{view}_n(g)$, where

1. $\text{view}_0(g) = \emptyset$.
2. $\forall i : \text{view}_i(g) \subseteq P$, where $P$ is the set of all processes in the system.
3. $\text{view}_i(g)$ and $\text{view}_{i+1}(g)$ differ by the addition or subtraction of exactly one process.

We assume that some sort of process group service computes new views and communicates them to the members of the groups involved. Processes learn of the failure of other group members only through this view mechanism, never through any sort of direct observation.

We assume that direct communication between processes is always possible; the software implementing this is called the message transport layer. Within our protocols, processes always communicate using point-to-point and multicast messages; the latter may be transmitted using multiple point-to-point messages if no more efficient alternative is available.

The transport communication primitives must provide lossless, uncorrupted, sequenced message delivery. Our approach permits application builders to define new transport protocols, perhaps to take advantage of special hardware. Our initial implementation uses unreliable datagrams, but has an experimental protocol that exploits ethernet hardware multicast.

The execution of a process is a partially ordered sequence of events, each corresponding to the execution of an indivisible action. An acyclic event order, denoted $\prec$, reflects the dependence of events occurring at process $p$ upon one another. The event $\text{send}_p(m)$ denotes the transmission of $m$ by process $p$ to a set of 1 or more destinations $\text{dests}(m)$; the receive event is denoted $\text{recv}_p(m)$. We omit the subscript when the context is unambiguous.

If $|\text{dests}(m)| > 1$ we will assume that $\text{send}$ puts messages into all communication channels in a single action that might be interrupted by failure, but not by other $\text{send}$ or $\text{recv}$ actions. We denote by $\text{recv}_p(\text{view}_i(g))$ the event by which a process $p$ belonging to $g$ "learns" of $\text{view}_i(g)$.

We distinguish the event of receiving a message from the event of delivery, since this allows us to model protocols that delay message delivery until some condition is satisfied. The delivery event is denoted $\text{deliver}(m)$ where $\text{recv}(m) \prec \text{deliver}(m)$.

### 2.2 Properties required of multicast protocols

Although Isis makes heavy use of virtual synchrony, it will not be necessary to formalize this property for our present discussion. However, the support of virtual synchrony places several obligations on the processes in our system. First, when a process multicasts a message $m$ to group $g$, $\text{dests}(m)$ must be the current membership of $g$. Secondly, when the group view changes, all messages sent in the prior view must be "flushed" out of the system (delivered) before the new view may be used. Finally, messages must satisfy a
failure atomicity property: if a message \( m \) is delivered to any member of a group, and it stay operational, \( m \) must be delivered to all members of the group even if the sender fails before completing the transmission.

The multicast protocols that interest us here also provide delivery ordering guarantees. As in [Lam78], we define the potential causality relation for the system, \( \rightarrow \), as the transitive closure of the relation defined as follows:

1. If \( \exists p : e \rightarrow^p e' \), then \( e \rightarrow e' \)
2. \( \forall m : \text{send}(m) \rightarrow \text{recv}(m) \)

CBCAST satisfies a causal delivery property:

If \( m \) and \( m' \) are CBCAST's and \( \text{send}(m) \rightarrow \text{send}(m') \) then

\[ \forall p \in \text{dests}(m) \cap \text{dests}(m') : \text{deliver}(m) \rightarrow^p \text{deliver}(m'). \]

If two CBCAST messages are concurrent, the protocol places no constraints on their delivery ordering at overlapping destinations.

ABCAST extends the CBCAST ordering into a total one:

If \( m \) and \( m' \) are ABCAST's then either

1. \( \forall p \in \text{dests}(m) \cap \text{dests}(m') : \text{deliver}(m) \rightarrow^p \text{deliver}(m'), \) or
2. \( \forall p \in \text{dests}(m) \cap \text{dests}(m') : \text{deliver}(m') \rightarrow^p \text{deliver}(m). \)

Because the ABCAST protocol orders concurrent events, it is more costly than CBCAST; requiring synchronous solutions where the CBCAST protocol admits efficient asynchronous solutions. Birman and Joseph [BJ89] and Schmuck [Sch88] have exhibited a large class of algorithms that can be implemented using asynchronous CBCAST. Moreover, Schmuck has shown that in many settings algorithms specified in terms of ABCAST can be modified to use CBCAST without compromising correctness.

The protocols presented here all assume that processes only multicast to groups that they are members of, and that all multicasts are to the full membership of a single group.

For demonstrating liveness, we will assume that any message sent by a process is eventually received unless the sender or destination fails, and that failures are eventually reported by ISIS.

### 3 The CBCAST bypass protocol

This section presents two basic CBCAST protocols for use within a single process group with fixed membership. Both use timestamps to delay messages that arrive out of causal order. The section that follows extends these schemes and then merges them to obtain a single solution for use with multiple, dynamic process groups.
3.1 Timestamping protocols

We begin by describing two protocols for assigning timestamps to messages and for comparing timestamps. The protocols are standard except in one respect: whereas most timestamping protocols count arbitrary "events", the ones defined here count only send events.

3.2 Logical time

The first timestamping protocol is based on one introduced by [Lam78], called the logical clock protocol. Each process $p$ maintains an unbounded local counter, $LT(p)$, which it initializes to zero. For each event $send(m)$ at $p$, $p$ sets $LT(p) = LT(p) + 1$. Messages are timestamped with the sender's incremented counter. A process $p$ receiving a message with timestamp $LT(m)$ sets $LT(p) = max(LT(p), LT(m))$. As in [Lam78], one can show that if $send(m) \rightarrow send(m')$ then $LT(m) < LT(m')$. The converse, however, does not hold: the protocol may order messages that were sent concurrently.

Note that the $LT$ counter for a process is updated at the recv event, as opposed to the deliver event, for an incoming message. We make use of this property in the development below.

3.3 Vector time

A second timestamping protocol is based on the substitution of vector times for the local counters in the logical time protocol. Vector times were proposed originally in [Mar84]; other researchers have also used them [Fid88, Mat89, LL86, Sch88]; our use of them is motivated by a protocol presented in [SES89]. In comparison with logical times, this protocol has the advantage of representing $\rightarrow$ precisely.

A vector time for a process $p_i$, denoted $VT(p_i)$, is a vector of length $n$ (where $n = |P|$), indexed by process-id.

1. When $p_i$ starts execution, $VT(p_i)$ is initialized to zeros.

2. For each event $send(m)$ at $p_i$, $VT(p_i)[i]$ is incremented by 1.

3. Each message sent by process $p_i$ is timestamped with the incremented value of $VT(p_i)$.

4. When process $p_j$ delivers a message $m$ from $p_i$ containing $VT(m)$, $p_j$ modifies its vector time in the following manner:

$$\forall k \in 1..n : VT(p_j)[k] = max(VT(p_j)[k], VT(m)[k])$$

Rules for comparing vector times are:
1. \( VT_1 \leq VT_2 \iff \forall i: VT_1[i] \leq VT_2[i] \)

2. \( VT_1 < VT_2 \iff VT_1 \leq VT_2 \) and \( \exists i: VT_1[i] < VT_2[i] \)

Notice that in contrast to the rule for \( LT(p) \), \( VT(p) \) is updated at the \emph{deliver} event for an incoming message. We will make use of this distinction below.

It can be shown that given messages \( m \) and \( m' \), \( send(m) \rightarrow send(m') \iff VT(m) < VT(m') \) \[\text{MatS9,Fid88}\]: vector timestamps represent causality precisely. This constitutes the fundamental property of vector times, and the primary reason for our interest in such times as opposed to logical ones.

### 3.4 Causal message delivery

Recall that if processes communicate using \texttt{CBCAST}, all messages must be delivered in an order consistent with causality. Suppose that a set of processes \( P \) communicate using only broadcasts to the full set of processes in the system; that is, \( \forall m : dests(m) = P \). This hypothesis is unrealistic, but Section 4 will adapt the resulting protocol to a settings with multiple process groups.\(^1\) We now develop two \emph{delivery protocols} by which each process \( p \) receives messages sent to it, delays them if necessary, and then delivers them such that:

If \( send(m) \rightarrow send(m') \) then \( deliver(m) \rightarrow deliver(m') \).

#### 3.4.1 LT protocol

Our first solution to the problem is based on logical clocks; and is referred to as the LT protocol from here on. It is related to other solutions that have appeared in the literature \[\text{Lam78,CASD86}\] and will be used as a building block later on. The basic technique will be to delay a message until messages with at least as large a timestamp has been received from every other process in the system. However, since this would only work if every process sends an infinite stream of multicasts, a channel flushing mechanism is introduced to avoid potentially unbounded delays.

Say that the channel from process \( p_j \) to \( p_i \) has been \emph{flushed at time} \( LT(m) \) if \( p_i \) will never receive a message \( m' \) from \( p_j \) with \( LT(m') < LT(m) \). Flushing can be achieved by \emph{pinging}. To ping a channel, \( p_i \) sends \( p_j \) a timestamped inquiry message \( ink \), but without first incrementing \( LT(p_i) \). On receiving an inquiry \( p_j \), as usual, sets \( LT(p_j) = \max(LT(p_j), LT(ink)) \) and replies with an ack message containing \( LT(p_j) \), without modifying \( LT(p_j) \). On receiving the ack \( p_i \), as usual, sets \( LT(p_i) = \max(LT(p_i), LT(ack)) \).

If no new messages are being multicast, pinging advances \( LT(p_i) \) and \( LT(p_j) \) to the same value.

The protocol is as follows:

\(^1\)This hypothesis is actually used only in the VT delivery protocol.
1. Before sending message \( m \), process \( p_i \) increments \( LT(p_i) \) and then timestamps \( m \).

2. On receiving message \( m \), process \( p_j \) sets \( LT(p_j) = \max(LT(p_j), LT(m)) \). Then, \( p_j \) delays \( m \) until for all \( k \neq i \), the channel between \( p_j \) and \( p_k \) has been flushed for time \( LT(m) \). \( p_j \) does not delay messages received from itself.

3. If \( m \) has the minimum timestamp among messages satisfying (2), \( m \) may be delivered.

To prove that causal delivery is achieved, consider two messages such that \( send(m_1) \rightarrow send(m_2) \), and hence \( LT(m_1) < LT(m_2) \). There are two cases:

1. **The same process sends \( m_1 \) and \( m_2 \).** Because communication is FIFO, \( m_1 \) will be received before \( m_2 \), and because \( LT(m_1) < LT(m_2) \), condition 3 guarantees that \( m_1 \) will be delivered before \( m_2 \).

2. **Different processes send \( m_1 \) and \( m_2 \).** According to condition 2, \( m_2 \) can only be delivered when all channels have been flushed for \( LT(m_2) \). As communication is FIFO, and \( LT(m_1) < LT(m_2) \), it follows that \( m_1 \) has been received. Condition 3 then guarantees that \( m_1 \) will be delivered before \( m_2 \).

The communication cost, however, is high: \( 2n - 3 \) messages may be needed to flush channels for every message delivered, hence to multicast one message, \( O(n^2) \) messages could be transmitted. For infrequent multicasting, this cost may well be tolerable; the overhead would be unacceptable if incurred frequently. However, in place of pinging, processes can periodically multicast their logical timestamps to all other group members. Receipt of such a multicast flushes the channels: at worst, a received message will be delayed until the recipient has multicast its timestamp and all other processes have done a subsequent timestamp multicast. The overhead of the protocol can now be tuned for a given environment.\(^2\)

### 3.4.2 VT protocol

A much cheaper solution can be derived using vector timestamps; we will refer to this as the VT protocol. The idea is basically the same as in the LT protocol, but because \( VT(m)[k] \) indicates precisely how many multicasts by process \( p_k \) precede \( m \), a recipient of \( m \) will know precisely how long \( m \) must be delayed prior to delivery; namely, until

\(^2\)Readers familiar with the \( \Delta-T \) real-time protocols of \cite{CASD86} will note the similarity between that protocol and this version of ours. Clock synchronization (on which the \( \Delta-T \) scheme is based) is normally done using periodic multicasts \cite{ST87}. This modification recalls suggestions made in \cite{Lam78}, and makes logical clocks behave like weakly synchronized physical clocks. Clock synchronization algorithms with good message complexity are known, hence substitution of a \( \Delta-T \) based protocol for the logical clock-based protocol in our "combined" algorithm, below, is an intriguing direction for future study.
VT(m)[k] messages have been delivered from pk. Since → is an acyclic order accurately represented by the vector time, the resulting delivery order is causal and deadlock free. The protocol is as follows:

1. Before sending m, process pi increments VT(pi)[i] and timestamps m.
2. On reception of message m sent by pi and timestamped with VT(m), process pj ≠ pi delays m until
   \[ \begin{align*}
   VT(m)[i] &= VT(p_j)[i] + 1 \\
   \forall k \neq i : VT(m)[k] &\leq VT(p_j)[k]
   \end{align*} \]
   Process pj need not delay messages received from itself.
3. When a message m is delivered, VT(p_j)[i] is incremented (this is simply the vector time update protocol from Section 3.3).

Step 2 is the key to the protocol. This guarantees that any message m' transmitted causally before m (and hence with VT(m') < VT(m)) will be delivered at pj before m is delivered. An example in which this rule is used to delay delivery of a message appears in Figure 1.

![Diagram](image)

**Figure 1:** Using the VT rule to delay message delivery

The correctness of the protocol will be proved in two stages. We first show that causality is never violated (safety) and then we demonstrate that the protocol never delays a message indefinitely (liveness).

**Safety.** Consider the actions of a process pj that receives two messages m₁ and m₂ such that send(m₁) → send(m₂).
Case 1. $m_1$ and $m_2$ are both transmitted by the same process $p_i$. Recall that we assumed a lossless, sequenced communication system, hence $p_j$ receives $m_1$ before $m_2$. By construction, $VT(m_1) < VT(m_2)$, hence under step 2, $m_2$ can only be delivered after $m_1$ has been delivered.

Case 2. $m_1$ and $m_2$ are transmitted by two distinct processes $p_i$ and $p_j$. We will show by induction on the messages received by process $p_j$ that $m_2$ cannot be delivered before $m_1$. Assume that $m_1$ has not been delivered and that $p_j$ has received $k$ messages.

Observe first that $send(m_1) \to send(m_2)$, hence $VT(m_1) < VT(m_2)$ (basic property of vector times). In particular, if we consider the field corresponding to process $p_i$, the sender of $m_1$, we have
\[ VT(m_1)[i] \leq VT(m_2)[i] \]  

(1)

Base case. The first message delivered by $p_j$ cannot be $m_2$. Recall that if no messages have been delivered to $p_j$, then $VT(p_j)[i] = 0$. However, $VT(m_1)[i] > 0$ (because $m_1$ is sent by $p_i$), hence $VT(m_2)[i] > 0$. By application of step 2 of the protocol, $m_2$ cannot be delivered by $p_j$.

Inductive step. Suppose $p_j$ has received $k$ messages, none of which is a message $m$ such that $send(m_1) \to send(m)$. If $m_1$ has not yet been delivered, then
\[ VT(p_j)[i] < VT(m_1)[i] \]  

(2)

This follows because the only way to assign a value to $VT(p_j)[i]$ greater than $VT(m_1)[i]$ is to deliver a message from $p_i$ that was sent subsequent to $m_1$, and such a message would be causally dependent on $m_1$. From relations 1 and 2 it follows that
\[ VT(p_j)[i] < VT(m_2)[i] \]

By application of step 2 of the protocol, the $k + 1$'st message delivered by $p_j$ cannot be $m_2$. □

Liveness. Suppose that there exists a broadcast message $m$ sent by process $p_i$ that can never be delivered to process $p_j$. Step 2 implies that either:
\[ VT(m)[i] \neq VT(p_j)[i] + 1, \text{ or} \]
\[ \exists k \neq i : VT(m)[k] > VT(p_j)[k] \]

and that $m$ was not transmitted by process $p_j$. We consider these cases in turn.

1. $VT(m)[i] \neq VT(p_j)[i] + 1$, that is, $m$ is not the next message to be delivered from $p_i$ from $p_j$. Since all messages are multicast to all processes and channels are lossless and sequenced, it follows that there must be some message $m'$ sent by $p_i$ that $p_j$ received previously, has not yet delivered, and with $VT(m')[i] = VT(p_j)[i] + 1$. If $m'$ is also delayed, it must be under the other case.
2. \(\exists k \neq i : VT(m)[k] > VT(p_i)[k]\). Let \(n = VT(m)[k]\). The \(n^\text{th}\) transmission of process \(p_k\), must be some message \(m' \rightarrow m\) that has either not been received at \(p_j\), or was received and is delayed. Under the hypothesis that all messages are sent to all processes, \(m'\) was already multicast to \(p_j\). Since the communication system eventually delivers all messages, we may assume that \(m'\) has been received by \(p_j\). The same reasoning that was applied to \(m\) can now be applied to \(m'\). The number of messages that must be delivered before \(m\) is finite and \(>\) is acyclic, hence this leads to a contradiction. \(\square\)

4 Extensions to the basic protocol

Neither of the protocols in Section 3 is suitable for use in a virtually synchronous setting with multiple process groups and dynamically changing group views. This section first extends the simple VT CBCAST protocol of Section 3.4.2 into one suitable for use with multiple but static process groups, but arrives at a protocol subject to a significant constraint on what we call the communication structure of the system. Then, we show how to combine the protocol with other mechanisms, notably the LT CBCAST protocol of Section 3.4.1, to overcome this limitation. We arrive at a powerful, general solution.

4.1 Transmission limited to within a single process group

The first extension to the VT protocol is concerned with processes that multicast only within a single process group at a time. This problem is clearly trivial if process groups don't overlap, a property that can be deduced at runtime (see Section 4.4.4). On the other hand, we have assumed that overlap will not be uncommon. Such scenarios motivate the series of changes to the algorithm presented in this section and the ones that follow.

The first change is concerned with processes that belong to multiple groups, e.g. a process \(p_i\) belongs to groups \(g_a\) and \(g_b\), and multicasts only within groups. Multicasts sent by \(p_i\) to \(g_a\) must be distinguished from those to \(g_b\), since a process \(p_j\) belonging to \(g_b\) and not to \(g_a\) that receives a message with \(VT(m)[j] = k\) will otherwise have no way to determine how many of these \(k\) messages were sent to \(g_b\) and hence precede \(m\) causally. This leads us to extend the single VT clock to multiple VT clocks; \(VT_a\) is the logical clock associated with group \(g_a\), and \(VT_a[i]\) thus counts multicasts by process \(p_i\) to group \(g_a\). Processes maintain VT clocks for each group in the system, and attach all the VT clocks to every message that they multicast.

The next change is to step 2 of the VT protocol. Suppose that process \(p_j\) receives a message \(m\) sent in group \(g_a\) with sender \(p_i\), and that \(p_j\) also belongs to groups \(\{g_1, \ldots, g_n\} \equiv G_j\). Step 2 can be replaced by the following rule:

\[\text{Clearly, if } p_i \text{ is not a member of } g_a, \text{ then } VT_a[i] = 0, \text{ thus allowing a sparse representation of the timestamp. For clarity, we will continue to represent each timestamp } VT_a \text{ as a vector of length } n, \text{ with a special entry } \ast \text{ for each process that is not a member of } g_a.\]
On reception of message $m$ from $p_i \neq p_j$, sent in $g_a$, process $p_j$ delays $m$ until

2.1' $VT_a(m)[i] = VT_a(p_j)[i] + 1$, and

2.2' $\forall k : (p_k \in g_a \land k \neq i) : VT_a(m)[k] \leq VT(p_j)[k]$, and

2.3' $\forall g : (g \in G_j) : VT_g(m) \leq VT_g(p_j)$.

As above, $p_j$ does not delay messages received from itself.

Figure 2 illustrates the application of this rule in an example with four processes into groups identified as $p_1...p_4$. Processes $p_1$, $p_2$ and $p_3$ belong to group $G_1$, and processes $p_2$, $p_3$ and $p_4$ to group $G_2$. Notice that $m_2$ and $m_3$ are delayed at $p_3$, because it is a member of $G_1$ and must receive $m_1$ first. However, $m_2$ is not delayed at $p_4$, because $p_4$ is not a member of $G_1$. And $m_3$ is not delayed at $p_2$, because $p_2$ has already received $m_1$ and it was the sender of $m_2$.

![Diagram of message sent within process groups](image)

Figure 2: Messages sent within process groups. $G_1 = \{p_1, p_2, p_3\}$ and $G_2 = \{p_2, p_3, p_4\}$

The proof of Section 3 adapts without difficulty to this new situation; we omit the nearly identical argument. One can understand the modified VT protocol in intuitive terms. By ignoring the vector timestamps for certain groups in step 2.3', we are asserting that there is no need to be concerned that any undelivered message from these groups could causally precede $m$. But, the ignored entries correspond to groups to which $p_j$ does not belong, and it was assumed that all communication is done within groups.

4.2 Use of partial vector timestamps

Until the present, we have associated with each message a vector time or vector times having a total size determined by the number of processes and groups comprising the
application. Although such a constraint arises in many published CBCAST protocols, the resulting vector sizes would rapidly grow to dominate message sizes. A substantial reduction in the number of vector timestamps that each process must maintain and transmit is possible in the case of certain communication patterns, which are defined precisely below. Even if communication does not always follow these patterns, our new solution can form the basis of other slightly more costly solutions which are also described below.

Define the communication structure of a system to be an undirected graph \( CG = (G, E) \) where the nodes, \( G \), correspond to process groups and edge \( (g_1, g_2) \) belongs to \( E \) iff there exists a process \( p \) belonging to both \( g_1 \) and \( g_2 \). If the graph so obtained has no biconnected component \(^4\) containing more than \( k \) nodes, we will say that the communication structure of the system is \( k \)-bounded. In a \( k \)-bounded communication structure, the length of the largest simple cycle is \( k \). \(^5\) A 0-bounded communication structure is a tree (we neglect the uninteresting case of a forest). Clearly, such a communication structure is acyclic.

Notice that causal communication cycles can arise even if \( CG \) is acyclic. For example, in figure 2, message \( m_1, m_2, m_3 \) and \( m_4 \) form a causal cycle spanning both \( g_1 \) and \( g_2 \). However, the acyclic structure restricts such communication cycles in a useful way—such cycles will either be simple cycles of length 2, or complex cycles.

Below, we demonstrate that it is unnecessary to transport all vector timestamps on each message in the \( k \)-bounded case. If a given group is in a biconnected component of size \( k \), processes in this group need only to maintain and transmit timestamps for other groups in this biconnected component. We can also show that they need to maintain at least these timestamps. As a consequence, if the communication structure is acyclic, processes need only maintain the timestamps for the groups to which they belong.

We proceed to the proof of our main result in stages. First we address the special case of an acyclic communication structure.

**Lemma 1:** If a system has an acyclic communication structure, each process in the system only maintains and multicasts the VT timestamps of groups to which it belongs.

Notice that under this lemma, the overhead on a message is limited by the size and number of groups to which a process belongs.

We wish to show that if message \( m_1 \) is sent (causally) before message \( m_k \), then \( m_1 \) will be delivered before \( m_k \) at all overlapping sites. Consider the chain of messages below.

\[
\begin{align*}
m_1 & \rightarrow m_2 \rightarrow m_3 \rightarrow \ldots \rightarrow m_{k-1} \rightarrow m_k \\
p_1 & \rightarrow p_2 \rightarrow p_3 \rightarrow \ldots \rightarrow p_{k-1} \rightarrow p_k \\
g_1 & \rightarrow g_2 \rightarrow g_3 \rightarrow \ldots \rightarrow g_{k-1} \rightarrow g_k
\end{align*}
\]

This schema signifies that process \( p_1 \) multicasts message \( m_1 \) to group \( g_1 \), that process \( p_2 \) first receives message \( m_1 \) as a member of group \( g_1 \) and then multicasts \( m_2 \) to \( g_2 \),

\(^4\)Two vertices are in the same biconnected component of a graph if there is a path between them after any other vertex has been removed.

\(^5\)The nodes of a simple cycle (other than the starting node) are distinct; a complex cycle may contain arbitrary repeated nodes.
and so forth. In general, \( g_i \) may be the same as \( g_j \) for \( i \neq j \) and \( p_i \) and \( p_j \) may be the same even for \( i \neq j \) (in other words, the processes \( p_i \) and the groups \( g_i \) are not necessarily all different). Let the term message chain denote such a sequence of messages, and let the notation \( m_i \xrightarrow{p} m_j \) mean that \( p \) transmits \( m_j \) using a timestamp \( VT(m_j) \) that directly reflects the transmission of \( m_i \). For example, say that \( m_i \) was the \( k \)'th message transmitted by process \( p_i \) in group \( g_a \). We will write \( m_i \xrightarrow{p} m_j \) iff \( VT_a(p_j)[i] \geq k \) and consequently \( VT_a(m_j)[i] \geq k \). Our proof will show that if \( m_i \rightarrow m_j \) and the destinations of \( m_i \) and \( m_j \) overlap, then \( m_i \xrightarrow{p} m_j \), where \( p_j \) is the sender of \( m_j \).

We now note some simple facts about this message chain that we will use in the proof. Recall that a multicast to a group \( g_a \) can only be performed by a process \( p_i \) belonging to \( g_a \). Also, since the communication structure is acyclic, processes can be members of at most two groups. Since \( m_k \) and \( m_1 \) have overlapping destinations, and \( p_2 \), the destination of \( m_1 \), is a member of \( g_1 \) and \( g_2 \), then \( g_k \), the destination of the final broadcast, is either \( g_1 \) or \( g_2 \). Since \( CG \) is acyclic, the message chain \( m_1 \ldots m_k \) simply traverses part of a tree reversing itself at one or more distinguished groups. We will denote such a group \( g_r \). Although causality information is lost as a message chain traverses the tree, we will show that when the chain reverses itself at some group \( g_r \), the relevant information will be "recovered" on the way back.

Proof of Lemma 1: The proof is by induction on \( l \), the length of the message chain \( m_1 \ldots m_k \). Recall that we must show that if \( m_1 \) and \( m_k \) have overlapping destinations, they will be delivered in causal order at all such destinations, i.e \( m_1 \) will be delivered before \( m_k \).

Base case. \( l = 2 \). Here, causal delivery is trivially achieved, since \( p_k \equiv p_2 \) must be a member of \( g_1 \) and \( m_k \) will be transmitted with \( g_1 \)'s timestamp. It will therefore be delivered correctly at any overlapping destinations.

Inductive step. Suppose that our algorithm delivers all pairs of causally related messages correctly if there is a message chain between them of length \( l < k \). We show that causality is not violated for message chains where \( l = k \).

Consider a point in the causal chain where it reverses itself. We represent this by \( m_{r-1} \rightarrow m_r \rightarrow m_{r+1} \rightarrow m_{r+2} \), where \( m_{r-1} \) and \( m_{r+1} \) are sent in \( g_{r-1} \equiv g_{r+1} \) by \( p_r \) and \( p_{r+1} \) respectively, and \( m_r \) and \( m_{r+1} \) are sent in \( g_r \) by \( p_r \) and \( p_{r+1} \). Note that \( p_r \) and \( p_{r+1} \) are members of both groups. This is illustrated in Figure 3. Now, \( m_r \) will not be delivered at \( p_{r+1} \) until \( m_r \) has been delivered there, since they are both broadcast in \( G_r \). We now have \( m_{r-1} \rightarrow m_r \xrightarrow{p_r} m_{r+1} \rightarrow m_{r+2} \). We have now established a message chain between \( m_1 \) and \( m_k \) where \( l < k \). So, by the induction hypothesis, \( m_1 \) will be delivered before \( m_k \) at any overlapping destinations, which is what we set out to prove. 

We now proceed to prove the main theorem.
Theorem 1: Each process $p_i$ in a system needs only to maintain and multicast the VT timestamps of groups in the biconnected components of $CG$ to which $p_i$ belongs.

Proof: As with Lemma 1, our proof will focus on the message chain that established a causal link between the sending of two messages with overlapping destinations. This sequence may contain simple cycles of length up to $k$, where $k$ is the size of the largest biconnected component of $CG$. Consider the simple cycle illustrated below, contained in some arbitrary message chain.

$$m_1 \rightarrow \ldots \rightarrow m_c \rightarrow m_{c+1}$$

Now, since $p_1$, $p_2$ and $p_3$ are all in groups in a simple cycle of $CG$, all the groups are in the same biconnected component of $CG$, and all processes on the message chain will maintain and transmit the timestamps of all the groups. In particular, when $m_c$ arrives at $p_3$, it will carry a copy of $VT_{g_1}$ that indicates that $m_1$ was sent. This means that $m_c$ will not be delivered at $p_3$ until $m_1$ has been delivered there. So $m_{c+1}$ will not be transmitted by $p_3$ until $m_1$ has been delivered there. Thus $m_1 \rightarrow p_3 \rightarrow m_{c+1}$. We may repeat this process for each simple cycle of length greater than 2 in the causal chain, reducing it to a chain within one group. We now apply Lemma 1, completing the proof. □

Theorem 1 shows us what timestamps are sufficient in order to assure correct delivery of messages. Are all these timestamps in fact necessary? It turns out that the answer is yes. It is easy to show that if a process that is a member of a group within a biconnected com-
ponent of $CG$ does not maintain a VT timestamp for some other group in $CG$, causality may be violated. We therefore state without formal proof:

**Theorem 2:** If a system uses the VT protocol to maintain causality, it is both necessary and sufficient for a process $p$ to maintain and transmit those VT timestamps corresponding to groups in the biconnected component of $CG$ to which $p$ belongs.

### 4.3 Extensions to arbitrary communication structures

In general, managing information concerning the biconnected components of $CG$ may be difficult, especially in a dynamic environment. We believe that the most practical use of the above result is in the acyclic case, since a process can conservatively determine that it is not in any cycle by observing that the group of which it is a member overlaps with at most one other group—a completely local test (but see also Section 4.4.4). Consequently, although all our results generalize, the remainder of the paper focuses on the acyclic solution, and we initially implemented only the acyclic solution in ISIS. In this section, we give two protocols that work in more general communication structures. The first protocol does not use any knowledge about the communication structure, but it sometimes imposes delays on message multicasting. The second protocol does use knowledge about the communication structure, but does not impose delays on message multicasting. We then extend both protocols to arbitrary dynamic communication structures.

#### 4.3.1 Conservative solution

Our first solution is denoted the *conservative protocol*. Each multicast $m$ is followed by a second multicast $\text{terminate}(m)$ signifying that $m$ has reached all of its destinations. The sender of a multicast will normally know when to send the $\text{terminate}$ as a side-effect of the protocol used to overcome packet loss. The $\text{terminate}$ message may sent as a separate multicast, but it can also be piggybacked on the next CBCAST sent to the same group. A $\text{terminate}$ message is not itself terminated.

We will say that a group is *active for process $p$*, if:

1. $p$ is the initiator of a multicast to $g$ that has not terminated, or
2. $p$ has received an unterminated multicast to $g$, or
3. $p$ has delayed the local delivery of a multicast to $g$ (sent by some other process $p'$).

Note that this is a local property; i.e. process $p$ may compute whether or not it is active for some group $g$ by examining its local state. The *conservative multicast rule* states that a process $p$ may multicast to group $g$ iff $g$ is the only active group for process $p$ or $p$ has no active groups. Multicasts are sent using the VT protocol, as usual. Notice that this rule imposes a delay only when two causally successive messages are sent to different
groups. The conservative solution could be inefficient, but yields a correct VT protocol. However, the overhead it imposes could be substantial if processes multicast to several different groups in quick succession, and it is subject to potential starvation (this can, however, be overcome).

The conservative solution will work correctly even if group membership changes dynamically.

For brevity, we omit the correctness proof of this solution. The key point is that if \( p \) multicasts \( m \) to \( g_2 \) after \( g_1 \) has ceased to be active, then there are no undelivered multicasts \( m' \) in \( g_1 \) s.t. \( m' \rightarrow m \). This can be demonstrated by showing that if \( g_1 \) is no longer active and \( m' \rightarrow m \), then \( m' \) has terminated.

### 4.3.2 Excluded Groups

Assume that \( CG \) contains cycles, but that some mechanism has been used to select a subset of edges \( X \) such that \( CG' = (G, E - X) \) is known to be acyclic. We extend our solution to use the acyclic VT protocol for most communication within groups. If there is some \( g' \) such that \( (g, g') \in X \) we will say that group \( g \) is an excluded group and some multicasts to or from \( g \) will be done using one of the protocols described below.

Keeping track of excluded groups could be difficult; however it is easy to make pessimistic estimates (and we will derive an protocol that works correctly with such pessimistic estimates). For example, in Isis, a process \( p \) might assume that it is in an excluded group if there is more than one other neighboring group. This is a safe assumption; any group in a cycle in \( CG \) will certainly have two neighboring groups. This subsection and the two that follow develop solutions for arbitrary communication structures, assuming that some method such as the previous is used to safely identify excluded groups.

### 4.3.3 Combining the VT and LT protocols

Recall the LT multicast protocol presented in Section 3. The protocol was inefficient, but required that only a single timestamp be sent on each message. Here, we run the LT and VT protocols simultaneously, piggybacking on each message both LT and VT timestamps, and apply a unified version of the LT and VT delivery schemes on receipt. The LT timestamp is not incremented on every broadcast; it is only incremented on certain broadcasts as described below. This greatly reduces the number of extra messages that would be induced by the basic LT algorithm.

Say that \( m \) is to be multicast by \( p \) to group \( g \). We say that \( p \) is not safe in \( g \) if:

- The last message \( p \) received was from some other group \( g' \).
- Either \( g \) or \( g' \) is an excluded group.
Our protocol rule is simple; on sending, if process \( p \) is not safe in group \( g \), \( p \) will increment both its' LT timesamp and its' VT timestamp before multicasting a message to \( g \). Otherwise, it will just increment its' VT timestamp. A message is delivered when it is deliverable according to both the LT delivery rule and the VT delivery rule.

Notice that the pinging overhead of the LT protocol is incurred only when logical clock values actually change, which is to say only on communication within two different groups in immediate succession, where one of the groups is excluded. That is, if process \( p \) executes for a period of time using the VT protocol and receives only messages that leave \( LT(p) \) unchanged, \( p \) will ping each neighbor processes at most once. Clocks will rapidly stabilize at the maximum existing LT value and pinging will then cease.

**Theorem 3:** The combined VT-LT protocol will always deliver messages correctly in arbitrary communication structures.

**Proof:** Consider an arbitrary message chain where the first and last messages have overlapping destinations. Without loss of generality, we will assume that \( g_1...g_k \) are distinct. We wish to show that the last message will be delivered after the first at all such destinations.

\[
\begin{align*}
& m_1 \quad m_2 \quad m_3 \quad \ldots \quad m_k-1 \quad m_k \\
& p_1 \rightarrow p_2 \rightarrow p_3 \rightarrow \ldots \rightarrow p_k \rightarrow p_{k+1} \\
& g_1 \quad g_2 \quad g_3 \quad \ldots \quad g_k-1 \quad g_k
\end{align*}
\]

If none of \( g_1...g_i \) is an excluded group, then, by Lemma 1, \( m_1 \) will be delivered before \( m_k \) at all overlapping destinations. Now, if some group \( g_i \) is excluded, two cases arise - either the last group, \( g_k \) is excluded, or some other group is excluded. If \( g_k \) is excluded, then \( p_k \) will increment its LT timestamp at some point between delivering \( m_{k-1} \) and sending \( m_k \). If some other group \( g_i \) is excluded, \( i < k \), then \( p_{k+1} \) will increment its LT timestamp between delivering \( m_k \) and sending \( m_{k+1} \). So the LT timestamp of \( m_k \) will always be greater than the LT timestamp of \( m_1 \), and \( m_k \) will be delivered after \( m_1 \) at all overlapping destinations. \( \square \)

### 4.4 Dynamic membership changes

We now consider the issue of dynamic group membership changes when using the combined protocol. This raises several issues that are addressed in turn: virtually synchronous addressing when joins occur, initializing VT timestamps, atomicity when failures occur, and the problem of detecting properties of \( CG \) at runtime, such as when a process determines that its' group adjoins at most on one other and hence always uses the acyclic VT protocol.
4.4.1 Joins

To achieve virtually synchronous addressing when group membership changes while multicasts are active, we introduce the notion of flushing the communication in a process group. Consider a process group \( g \) in group view \( \text{view}_i(g) \). Say that a new view \( \text{view}_{i+1}(g) \) now becomes defined. There are two cases: \( \text{view}_{i+1}(g) \) could reflect the addition of a new process, or it could reflect the departure (or failure) of a member. Assume initially that view changes are always due to adding new processes (we handle failures in Section 4.4.3). We will flush communication by having all the processes in \( \text{view}_{i+1}(g) \) send a message "flush \( i+1 \) of \( g \)" to all other members. During the period after sending such messages and before receiving such a flush message from all members of \( \text{view}_{i+1}(g) \) a process will accept and deliver messages but will not initiate new multicasts.

Because communication is FIFO, if process \( p \) has received a flush message from all members of \( g \) under view \( i+1 \), it will first have received any messages sent in view \( i \). It follows that all communication sent prior to and during the flush event was done using \( VT \) timestamps corresponding to \( \text{view}_i(g) \), and that all communication subsequent to installing the new view is sent using \( VT \) timestamps for \( \text{view}_{i+1}(g) \). This establishes that multicasts will be virtually synchronous in the sense of Section 2.

4.4.2 Initializing VT fields

Say that process \( p_j \) is joining group \( g_a \). Then \( p_j \) will need to obtain the current \( VT \) values for other group members. Because \( p_j \) participates in the flush protocol, this can be achieved by having each process include its \( VT \) value in the flush message. \( p_j \) will initialize \( VT_a[i] \) with the value it receives in the flush message from \( p_i \); \( p_j \) initializes \( VT_a[j] \) to 0.

4.4.3 Failure atomicity

What about the case where some member of \( g \) fails during an execution? \( \text{view}_{i+1}(g) \) will now reflect the departure of some process. Assume that process \( p_j \) has received a message \( m \) that was multicast by process \( p_i \). If \( p_i \) now fails before completing its multicast, there may be some third process \( p_k \) that has not yet received a copy of \( m \). To solve this problem, \( p_j \) must retain a copy of all delivered messages, transmitting a copy of messages initiated by \( p_i \) to other members of \( \text{view}(g) \) if \( p_i \) fails. Processes identify and reject duplicates.

Multicasting now becomes the same two-phase protocol needed to implement the conservative rule. The terminate message indicates which messages may be discarded; it can be sent as a separate message or piggybacked on some other multicast. On receiving \( \text{view}_k(g) \) indicating that \( p_i \) failed, \( p_j \) runs this protocol:

1. Close the channel to \( p_i \).
2. For any unterminated multicast \( m \) initiated by \( p_i \), send a copy of \( m \) to all processes in \( \text{view}_k(g) \) (duplicates are discarded on reception).

3. Send a flush message to all processes in \( \text{view}_k(g) \).

4. Simulate receipt of flush and ack messages from \( p_i \) as needed by the channel and view flush protocols, and treat any message being sent to \( p_i \) as having been delivered in the conservative protocol (Section 4.3.1).

5. After receiving flush messages from all processes in \( \text{view}_k(g) \), discard any messages delayed pending on a message from \( p_i \).

6. \( p_j \) ceases to maintain \( VT_g[i] \).

Step 2 ensures atomicity and step 4 prevents deadlock in the VT, LT and the conservative protocol. Step 5 relates to chains of messages \( m_1 \rightarrow m_2 \) where a copy of \( m_2 \) has been received but \( m_1 \) was lost in a failure; this can only happen if every process that received \( m_1 \) has failed (otherwise a copy of \( m_1 \) would have been received prior to receipt of the flush message). In such a situation, \( m_2 \) will never have been deliverable and hence can be discarded.

This touches on an important issue. Consider a chain of communication that arises external to a process group but dependent on a multicast within that group. Earlier, we showed that causal delivery is assured by the acyclic VT protocol, but this assumed that multicasts would not be lost. Instead, say that processes \( p_1 \) and \( p_2 \) belong to group \( g_1 \) and that process \( p_2 \) also belongs to \( g_2 \). \( p_1 \) multicasts \( m_1 \) to \( g_1 \); \( p_2 \) receives \( m_1 \) and multicasts \( m_2 \) to \( g_2 \). Now, if \( p_1 \) and \( p_2 \) both fail, it may be that \( m_1 \) is lost but that \( m_2 \) is received by the members of \( g_1 \cap g_2 \) that are still operational.

Several cases now arise, all troubling. Consider a process \( q \) that receives \( m_2 \). If \( q \) receives \( m_2 \) after running the failure protocol, it will discard it under step 5. If \( q \) receives \( m_2 \) after running the failure protocol, however, it will have discarded the \( VT \) field corresponding to \( p_1 \). \( m_2 \) will not be delayed pending receipt of \( m_1 \) and hence will ultimately be delivered, violating causality. (\( q \) cannot discard \( m_2 \) because it may have been delivered elsewhere.) We thus see that both causality and atomicity could be violated by an unfortunate sequence of failures coincident with a particular pattern of communication, and that the system will be unable to detect that this has occurred.

One way to avoid this problem is to require that processes always use the conservative rule of Section 4.3.1, even if the communication structure is known to be acyclic. In our example, this would prevent \( p_2 \) from communicating in \( g_2 \) until \( m_1 \) reached its destinations. Recall that step 4 of the protocol given above prevents the conservative rule from blocking when failures occur.

An alternative is to accept some risk and operate the system unsafely. For example, a process might be permitted to initiate a multicast to group \( g \) only if all of its own
multicasts to other groups have been delivered to at least one other destination process; this yields a protocol tolerant of any single failure.  

Given a 1-resilient protocol, the sequence of events that could cause causal delivery to be violated seems quite unlikely. A $k$-resilient protocol can be built by also delaying receivers; for large $k$, this reverts to the conservative approach.

We believe that even for a 1-resilient protocol, the scenario in question (two failures that occur in sequence simultaneously with a particular pattern of communication) is extremely improbable. The odds of such a sequence occurring is probably outweighed by the risk of a software bug or hardware problem that would cause causality to be violated for some mundane reason, like corruption of a timestamp or data structure.

Our initial implementation of bypass CBCAST uses the conservative solution between all groups; i.e. all groups are excluded. The VT protocol is used for communication within a group. This version of ISIS is thus immune to the causality and atomicity problems cited above, but incurs a high overhead if processes multicast to a series of groups in quick succession, which is not uncommon. Our plan is to modify the implementation to use the more optimistic protocols in a 1-resilient manner, but to provide application designers with a way to force the system into a completely safe mode of operation if desired. It should be noted that limitations such as this are common in distributed systems; a review of such problems is included in [BJ89]. We are not alone in advocating a "safe enough" solution in order to increase performance.

4.4.4 Dynamic communication graphs

A minor problem arises in applications having the following special structure:

1. The combined VT-LT protocol is in use.

2. Processes may leave groups other than because of failures (in ISIS, this is uncommon but possible).

3. Such a process may later join other groups.

Earlier, it was suggested that a process might observe that the (single) group to which it belongs is adjacent to just one other group, and conclude that it cannot be part of a cycle. In this class of applications, this rule may fail.

To see this, suppose that a process $p$ belongs to group $g_1$, then leaves $g_1$ and joins $g_2$. If there was no period during which $p$ belonged to both $g_1$ and $g_2$, $p$ would use the acyclic VT protocol for all communication in both $g_1$ and $g_2$. Yet, it is clear that $p$ represents a path by which messages sent in $g_2$ could be causally dependent upon messages $p$ received

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$^6$When using a transport facility that exploits physical multicast such a message will most often have reached all of its destinations.
in $g_1$, leading to a cyclic message chain that traverses $g_1$ and $g_2$. This creates a race condition under which violations of the causal delivery ordering could result.

This problem can be overcome in the following manner. Associate with each group a counter of the number of other groups to which it has ever been adjacent; this requires only a trivial extension of the flush protocol. Moreover, say that even after a process $p$ leaves a group $g_1$, it reports itself as a one-time member of $g_1$. If $p$ joins some group $g_2$, the adjacency count for $g_2$ will now reflect its prior membership, and if a causal chain could possibly arise, multicasts will be under the exclusion rule. Clearly, this solution is conservative and could be costly. On the other hand, say that it is known that all multicasts terminate within some time delay $\sigma$. Then one could decrement the adjacency counter for a group after a delay of $\sigma$ time units without risk. In Isis, a reasonable value of $\sigma$ would be on the order of 2-3 seconds.

We have developed more sophisticated solutions to this problem, but omit these because the issue only arises in a small class of applications, and the methods and their proofs are complex.

4.4.5 Recap of the extended protocol

In presenting our algorithm as a basic scheme to which a series of extensions and modifications were made, we may have obscured the overall picture. We conclude the section with a brief summary of the protocol as we intend to use it in Isis.

The protocol we ultimately plan to use in Isis is the acyclic VT solution combined with the LT protocol. This protocol piggybacks an LT timestamp and a list of VT timestamps on each message, one VT vector for each group to which the sender of the message belongs. In addition to the code for delaying messages upon reception, the protocol implements the channel- and view-flush and terminate algorithms.

Under most conditions the Isis system will be operated conservatively, excluding groups adjacent to more than one neighboring group. As noted above, neighboring groups can be counted by piggybacking information on the view-flush protocol. Looking to the future, we expect to develop Isis subsystems that will have special $a$-priori knowledge of the communication structure. These subsystems will make use of an Isis system call $pg\_exclude(gname, TRUE/FALSE)$ to indicate the exclusion status of groups. We currently have no plans to develop sophisticated communication topology algorithms for Isis.

The initial Isis implementation consists of the VT scheme and the conservative rule, together with the view-flush and terminate protocols. We expect to add the LT extension shortly; the necessary code is small compared to what is already running.
5 Other communication requirements

In this section we consider some minor extensions of the protocol for other common communication requirements.

5.1 A Bypass ABCAST protocol

Readers may wonder if the bypass CBCAST protocol can be extended into a fast ABCAST mechanism. ABCAST is a totally ordered communication protocol: all destinations receive an ABCAST message in a single, globally fixed order.

The answer to this question depends on the semantics one associates with ABCAST addressing. One way to define ABCAST is to say that two ABCAST's to the same logical address will be totally ordered, but to make no guarantees about ordering for ABCAST messages sent to different addresses. A more powerful alternative is to say that regardless of the destination processes, if two ABCAST's overlap at some set of destinations, they are delivered in the same order. Although Isis currently supports the latter approach, it is far easier to implement a bypass ABCAST with the weaker delivery semantics; the resulting protocol resembles the one in [CM84]. This is in contrast with bypass CBCAST, which always achieves causal ordering.

Associated with each view $\text{view}_i(g)$ of a process group $g$ will be a token holder process, $\text{token}(g) \in \text{view}_i(g)$. If the holder fails, the token is automatically reassigned to a live group member using any well-known, deterministic rule. Assume that each message $m$ is uniquely identified by $\text{uid}(m)$.

To ABCAST $m$, a process holding the token uses CBCAST to transmit $m$. If the sender is not holding the token, the ABCAST is done in stages:

1. The sender CBCAST's a needs-order message containing $m$.\footnote{It might appear cheaper to forward such a message directly to the token holder. However, for a moderately large messages such a solution will double the IO done by the token holder, creating a likely bottleneck, while reducing the IO load on other destinations only to a minor degree.} Processes receiving this message delay delivery of $m$.

2. If a process holding the token receives a needs-order message, it CBCAST's a sets-order message giving a list of one or more messages, identified by $\text{uid}$, and the order in which to deliver them, which it may chose arbitrarily. If desired, a new token holder may also be specified in this message.

3. On receipt of a sets-order, a process notes the new token holder and delivers delayed messages in the specified order.

4. On detection of the failure of the token holder, after completing the flush protocol, all processes sort pending ABCAST's and deliver them in any consistent order.
This protocol is essentially identical to the replicated data protocol proved correct in [BJ89, Sch88]. Step 4 is correct because the flush ensures that any set-order messages will have been delivered atomically, hence all processes will have the same enqueued messages which they deliver immediately before installing the new view.

The cost of doing a bypass ABCAST depends on the locations where multicasts originate and frequency with which the token is moved. If multicasts tend to originate at the same process repeatedly, then once the token is moved to that site, the cost is one CBCAST per ABCAST. If they originate randomly and the token is not moved, the cost is \(1 + 1/k\) CBCAST's per ABCAST, where we assume that one set-order message is sent for ordering purposes once for every \(k\) ABCAST's. This represents a major improvement over the existing Isis ABCAST protocol. However, because bypass ABCAST achieves a weaker form of ordering, it might require changes to existing Isis applications. We have not yet decided whether to make it the default.

5.2 Point-to-point messages

Early in the paper, we asserted that asynchronous CBCAST is the dominant protocol used in Isis. Point-to-point messages, arising from replies to multicast requests and and RPC interactions, are also common. In both cases, causal delivery is desired. Here, we consider the case of point-to-point messages sent by a process \(p\) within a group \(G\) to which \(p\) belongs.

A straightforward way to incorporate point-to-point messages into our VT protocol is to require that they be acknowledged and to inhibit the sending of new multicasts during the period between when such a message is transmitted and when the acknowledgement is received (in the case of an RPC, the reply is the acknowledgement). The recipient is not inhibited, and need not keep a copy of the message. A point-to-point message is timestamped using the sender's logical and vector times, and delivered using the corresponding delivery algorithms, but neither timestamp is incremented prior to transmission. In effect, point-to-point messages are treated as events internal to the processes involved.

The argument in favor of this method is that a single point-to-point RPC is fast and the cost is unaffected by the size of the system. Although one can devise more complex methods that eliminate the period of inhibited multicasting, problems of fault-tolerance render them less desirable.

5.3 Subset multicasts

Some Isis applications form large process groups but require the ability to multicast to subsets of the total membership. Our protocol is easily extended into one supporting subset multicast, and our initial Isis implementation supports this as an option. When enabled, a VT vector timestamp of length \(sn\) is needed for a group with \(s\) senders and \(n\) members.
For example, a stock brokerage might support a quote dissemination service with two or three transmitters and hundreds of potential recipients. Rather than form a subgroup for each stock (costly approach if there are many stocks), each multicast could be sent to exactly those group members interested in a given quote. We omit the details of the subset multicast extension.

6 Performance and transport protocol selection

In this section, we discuss the performance of our protocol. We show that the performance of the bypass protocol will be largely dominated by the performance of the underlying layer that is simply concerned with moving data from one site to others. We discuss the design of some alternatives for this layer, which we are currently implementing.

6.1 Complexity and overhead of the protocol

Implementation of the bypass protocol was straightforward in Isis, requiring less than 1300 lines of code out of the total of 52,000 in the protocol layer of the system. Extensions to support the LT protocol will add little additional code. Initial measurements of performance demonstrate a five to tenfold speedup over the prior Isis protocols.

Our protocol has an overhead of both space and messages transmitted. The size of a message will be increased by the vector time fields it carries; as noted above, the number of such vectors is determined by the total cardinality of the groups to which the sender belongs directly, and hence will be small. The number of overhead messages sent will depend on the number of non-piggybacked terminate messages sent by the conservative protocol and, when implemented, the frequency of LT pinging. In Isis, LT pinging is expected to be rare and terminate messages are always piggybacked on a subsequent CBCAST unless communication in a group quiesces. (As noted before, LT overhead can be bounded using a periodic protocol, if necessary).

We believe that latency, especially when the sender of a multicast must delay before continuing computation, is the most critical and yet unappreciated form of overhead. Delays of this form are extremely noticeable. In many systems, there is only one active computation at a given instant in time, or a single computation that holds a lock or other critical resource. Delaying the sender of a multicast may thus have the effect of shutting down the entire system. In contrast, the delay between when a message is sent and when it reaches a remote destination is less relevant to performance. The sender may be delayed in two ways: if the transmission protocol itself is computationally costly, or if a self-addressed multicast cannot be delivered promptly because it is unsafe to do so. Defined in this sense, our method imposes latency on the sender of a multicast only in the conservative protocol, and only when a process switches from multicasting in one group to another, or needs to communicate in one group after receiving in another. Otherwise, the protocol is totally asynchronous. Latency on the transport side is less critical. The
dominant source of transport latency is LT pinging, and we plan to quantify this effect by instrumenting Isis and using simulations.

6.2 Implementation

An interesting feature of the bypass facility is that it assumes very little about communication between processes, and communicates in an extremely regular manner. Specifically, the protocol we ended with sends or multicasts only within groups to which a sending process belongs, and requires only that inter-process communication be sequenced and lossless. The idea of providing an interface by which the bypass multicast protocols could run over a lower-layer protocol provided by the application appealed to us, and as part of the Isis implementation of bypass CBCAST and ABCAST, we included an interface permitting this type of extension. We call this lower layer the multicast transport protocol. A multicast transport protocol simply delivers messages reliably, in FIFO order, to the groups or processes addressed.

When no special hardware for multicasting is available, the basic Isis multicast transport protocol is based on UDP (unreliable datagrams). When multicasting hardware is available, Isis can switch to an experimental multicast transport protocol that takes advantage of such hardware. The remainder of this section details the design, performance and overhead of these multicast transport protocols (in time, size, and messages exchanged per multicast).

6.3 Overhead imposed by the basic VT Protocol

This section breaks down the costs we see in terms of various components of the overhead (create a lightweight task, do the I/O, select system call, create the packets, reconstruct them on reception). Figure 4 breaks down the basic CPU costs of sending and receiving messages in our implementation. These figures are preliminary and will be revised. These figures are for the combined protocol, but they do not reflect higher level delays that might be imposed by infrequent events such as LT pinging or the view flush. Our figures were derived on a pair of SUN 3/60's doing continuous null RPC's from one to the other. The RPC request was sent in a CBCAST; the result returned in a CBCAST reply packet. A new lightweight task was created at the receiver to field each RPC request. An Isis message is fairly complex and allows scatter/gather and arbitrary user-defined and system-checked types. Since no attempt has been made to optimize message data structures for the simple case of a null RPC, this accounts for a large part of the time spent in the messaging/task layer of the system.

The main conclusion from these measurements is that the CBCAST algorithms we derive in this paper are quite inexpensive. Most of the time that a message spends in transit is spent in the lower layers of the system. Clearly, the cost of UNIX messaging is beyond our control, but a great deal can be said about multicast transport.
6.4 Multicast transport protocol selection

The basic Isis multicast transport protocol is designed around a point-to-point model. Each process in a group maintains a two-way reliable data stream with each other process in the group. Whenever possible, acknowledgement information is piggybacked on other packets, such as replies to an RPC or multicast. These streams are maintained independently of each other; for brevity, we omit discussion of such details as flow control and failure detection. This scheme has several advantages; it is relatively easy to understand, as it is based on a well-known communication model. Since it is built on top of unreliable datagrams, it can be easily implemented on any network that provides this service. It has, however, several disadvantages - in particular, it does not scale well. The processing and network transmission costs of communicating with a group rise linearly with the number of processors in the group. In addition, as the number of processes in a group increases, a process sending to the group may experience congestion at the network interface as many acknowledgement or reply packets arrive more or less simultaneously from the other other processes in the group.

We have therefore investigated the design of other multicast transport protocols. An ideal multicast transport protocol would have the following features:

- It would be independent of network topology, but able to take advantage of features of particular networks - e.g. a broadcast subnet.

- The cost of sending a message would be independent of the number of recipients of
It would work efficiently for both small and large messages.

It would have low overhead, latency and high throughput.

It is also important to note that frequently a multicast may give rise to many replies directed to the original sender. We call such an occurrence a convergecast. This can lead to congestion at the original multicast sender, with many of the replies being lost. To avoid this, a multicast transport protocol should have some sort of mechanism for co-ordinating and reliably delivering multicast replies. Similar considerations may apply to acknowledgements; however acknowledgements need not be as timely as replies - the multicast transport protocol has more freedom to delay them.

Generally speaking, a reliable multicast transport mechanism will be used in two distinct modes. In the first, stream mode, one process will multicast a large amount of data to the group before another process wishes to reply. Multicasting is continuous. This usage could arise in, for example, a trading system, where the transport mechanism is being used to disseminate quotes to trading stations. Another example is a replicated file system where a client workstation is writing a file to a group of file servers. In rpc mode, many processes multicast replicated rpc's to the group, where each rpc contains relatively little data, and is much more likely to actually require a reply. Multicasts are not continuous, but bursty. This could arise in maintaining and querying a distributed database or maintaining the state of a distributed game. Note that the application using the multicast transport protocol can provide hints as to which mode it thinks it is operating in. Intermediate modes of usage can of course arise; we do not expect them to be common.

Reliable multicast transport protocols may be divided into two classes; those based on positive acknowledgements, and those based on negative acknowledgements. Many previous proposals for reliable multicast transport protocols have been based on negative acknowledgements, including [KTHB89,AHL89,CM84]. (Some of these protocols, in addition to providing reliable transport, also provide transport ordering properties.) This is because the designers of these protocols believed that a positive acknowledgement from each receiving site would be expensive. We do not believe that this is so.

If a process group is largely communicating in rpc mode, reply messages will be converging at the sender in any case. These reply messages can carry positive acknowledgements. In addition, if there are many of these reply messages, they should be scheduled by some mechanism to avoid congestion and message loss at the multicast sender. On the other hand, if a group is largely communicating in stream mode, the issue of flow control becomes very important. The sender can’t send data faster than the slowest process in the group can receive it; in order to avoid packet loss, there will be flow control packets coming back to the sender from each other process in the group. Again, these packets may carry positive acknowledgments, and again, they must be scheduled in order to avoid congestion problems. The protocol has more flexibility in scheduling these packets than in scheduling reply packets, since they do not contain data that needs to be delivered to the higher level.

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There are several possible mechanisms for scheduling packets that are converging on the same destination. One scheme is for the original sender to schedule the packets; it will decide how many concurrent acknowledgments or replies it (and the network) can handle. It then schedules each group of acknowledgements. This scheme involves some extra work by the sender; it has the advantage that the sender can control the rate at which the packets come back depending on whether or not his client is waiting for replies.

Other methods involve the receivers co-operating to ensure that they don’t send too many packets to the sender at once. One such method basically involves passing one or several tokens around the group, with the holder of a token having the right to send reply or acknowledgement packets to the original sender. If the replies or acknowledgements are small, they can be put on the token itself, which is returned to the sender when it is full. The main problem with this scheme is that the acknowledgement or reply may take a long time to return to the original sender of a message. This can be overcome by using large window sizes, or by using a large enough number of tokens. Another problem is that the overhead of receiving a message is higher, because an acknowledgement token must be received and transmitted also. This can be overcome by having one token acknowledge several messages, and by piggybacking the acknowledgement token wherever possible. A third problem is that the loss of one acknowledgement packet may cause a message to be retransmitted to multiple destinations. We believe that the extra overhead is acceptable, since packet loss should be rare.

Another receiver-scheduled method for handling acknowledgements or replies is simply to have each acknowledgement be returned at some random time by the recipients. This scheme has been extensively analyzed by [Dan89]; the main problem is that in order to avoid congestion at the original sender, the interval from which the random delays must be picked is very long. It is also of course possible to combine several of the above schemes; for example, acknowledgements could be sender-scheduled in small groups; individual acknowledgements within each group could be further randomly delayed.

We are implementing multicast transport protocols with several of the convergecast-avoidance scheduling strategies described above, and will experiment with them as alternatives to the basic ISIS multicast transport protocol. Our implementations are based on the multicast UDP software of [Dee88], which provides a logical unreliable multicast across internets independently of whether the underlying networks support physical multicast. Full details of the design and implementation of these protocols will be found in [Ste90]. We will include performance measurements for the bypass CBCAST and ABCAST protocols running over these transport protocols in the final version of the paper.
7 Related Work

There has been a great deal of work on multicast primitives. CBCAST-like primitives are described in [BJ87, PBS89, VRB89, SES89, LL86] As noted earlier, our work is most closely related to that of Ladkin and Peterson. Both of these efforts stopped at essentially the point we reached in Section 3 arriving a protocols that would perform well within a single small group, but subject to severe drawbacks in systems with large numbers of processes and of overlapping, dynamically changing process groups. Pragmatic considerations stemming from our desire to use the protocol in ISIS motivated us to take our protocol considerably further. We believe the resulting work to be interesting from a theoretical perspective. Viewed from a practical perspective, a causal multicast protocol that scales well and imposes little overhead under typical conditions certainly represents a valuable advance.

ABCAST-like primitives are reported in [CM84, BJ87, GMS89, PGM85]. Our ABCAST protocol is motivated by the Chang-Maxemchuck solution [CM84], but is simpler and faster because it can be expressed in terms of a virtually synchronous bypass CBCAST. In particular, our protocol avoids the potentially lengthy delays required by the Chang-Maxemchuck approach prior to committing a message delivery ordering. We believe this argues strongly for a separation of concerns in particular, a decoupling process group management from the communication primitive itself.

We note that of the many protocols described in the literature, very few have been implemented, and many have potentially unbounded overhead or postulate knowledge about the system communication structure that might be complex to deduce. This makes direct performance comparisons difficult, since many published protocols give performance estimates based on simulations or measure dedicated implementations on bare hardware. We are confident that the Isis bypass communication suite gives performance fully competitive with any alternative. The ability to extend the transport layer will enable the system to remain competitive even in settings with novel architectures or special communication hardware.

The ability to run the bypass protocols over new transport protocols raises questions for future investigation. For example, one might run bypass CBCAST over a transport layer with known realtime properties. Depending on the nature of these properties, such a composed protocol could satisfy both sets of properties simultaneously, or could favor one over the other. For example, the delay of flushing channels suggests that realtime and virtual synchrony properties are fundamentally incompatible, but this still leaves open the possibility of supporting a choice between weakening the realtime guarantees to ensure that the system will be virtually synchronous and weakening virtual synchrony to ensure that realtime deadlines are always respected. For many applications, such a choice could lead to an extremely effective, tuned solution. Pursuing this idea, we see the Isis system gradually evolving into a more modular structure composed of separable facilities for group view management, enforcing causality, transporting data, and so forth.
For a particular setting, one would select just those facilities actually needed. Such a compositional programming style has been advocated by others, notably Larry Peterson in his research on the Psync system.

8 Conclusions

We have presented a new scheme, the bypass protocol, for efficiently implementing a reliable, causally ordered multicast primitive. Intended for use in the Isis toolkit, it offers a way to bypass the most costly aspects of Isis while benefiting from virtual synchrony. The bypass protocol is inexpensive, yields high performance, and scales well. Measured speedups of more than an order of magnitude were obtained when the protocol was implemented within Isis. Our conclusion is that systems such as Isis can achieve performance competitive with the best existing multicast facilities - a finding contradicting the widespread concern that fault-tolerance may be unacceptably costly.

Acknowledgements

The authors are grateful to Keith Marzullo, who suggested we search for the necessary and sufficient conditions of Section 4.2, and made many other useful comments. Tushar Chandra found several bugs in our proofs. Mark Wood, Robert Cooper and Shivakant Misra made extensive suggestions concerning the protocols and presentation, which we greatly appreciate.

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