Virtually-Synchronous Communication
Based on a Weak Failure Suspector

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Abstract

Failure detectors (or, more accurately Failure Suspectors – FS) appear to be a fundamental service upon which to build fault-tolerant, distributed applications. This paper shows that a FS with very weak semantics (i.e. that delivers failure and recovery information in no specific order) suffices to implement virtually-synchronous communication (VSC) in an asynchronous system subject to process crash failures and network partitions. The VSC paradigm is particularly useful in asynchronous systems and greatly simplifies building fault-tolerant applications that mask failures by replicating processes. We suggest a three-component architecture to implement virtually-synchronous communication: 1) at the lowest level, the FS component; on top of it, 2a) a component that defines new views, and 2b) a component that reliably multicasts messages within a view. The issues covered in this paper also lead to a better understanding of the various membership service semantics proposed in recent literature.

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1 Introduction

There have recently been several papers about membership services in asynchronous systems [2, 12, 13, 17, 18, 19, 20]. A membership service is responsible for giving each process (consistent) information about the operational processes in the system. A process calls this information its view of the system processes. A membership service typically reacts to process crashes or recoveries, leading it to define a set of views. The membership services mentioned vary according to the underlying failure model considered, as well as the properties they provide with respect to the set of views delivered to each process: (e.g. whether another view may exist simultaneously, the degree of agreement among members):

- [17, 18] consider processes with crash failure semantics, excluding network partitions.
- [19, 20] consider systems in which processes may crash and the network may partition. However, despite network partitions, this membership service defines only majority views – a unique, totally-ordered sequence of views. Such a membership service is said to have linear semantics.
- The membership services described in [1, 2, 13] consider the same failure scenario as above, but only define a partial order on the views. That is, if the system is partitioned in two (or more) subnetworks then two (or more) views, one in each subnetwork, may exist concurrently.

Concurrent views offer an interesting extension to membership services, and force us to consider a further semantic distinction based on whether concurrent views are permitted to intersect. If two concurrent views may overlap, we say the membership service semantics are weak-partial, if they may not we say the semantics are strong-partial. Among those that permit concurrent views, [2] appears to be a strong-partial membership service. [13] considers both strong-partial and weak-partial membership services, and [1] and [12] consider only weak-partial membership service. These variants raise a new, pertinent question: when is a strong-partial service required, and when does a weak-partial membership service suffice. The objective of this paper is to suggest an answer to this question, by showing that a strong-partial membership service is intimately related to virtually-synchronous communication. We do not discuss when a linear membership service is required.

The idea of virtually-synchronous communication (VSC) was first introduced by Isis [3, 4]. VSC can be understood as rule for ordering message deliveries (reliable multicasts) with respect to view changes (received from the membership service). We give a precise definition.
for VSC in Section 5.4. VSC defines a powerful model for building fault-tolerant processes that mask failures by replication. It has also been argued [5] that ordering message deliveries consistently around process failures and recoveries is a fundamental part of any distributed computation; thus VSC is a vital primitive for inherently-distributed programming. Relatedly, many common distributed applications are more easily understood and solved if they can make use of VSC [21]. Finally, if the VSC abstraction we define in this paper is augmented with a majority requirement, [22] shows it is a powerful model in which transaction commit is easily (albeit probabilistically) implemented. Understanding that the VSC abstraction is more basic than the transaction abstraction gives broader insight to the problem of building fault-tolerant applications. However, we note that solving VSC is not equivalent to solving consensus [10].

Traditionally virtually-synchronous communication has been implemented with a two-component architecture: a membership service, and on top of it, multicast component. However, understanding the relationship between a membership service and virtually-synchronous communication has lead us to consider a three-component architecture, with (1) a Failure Suspector component FS delivering information about the communication topology, (2) a View Component VC defining views, and (3) a Multicast Component MC implementing virtually-synchronous communication. We divide the functionality of the traditional membership service between our FS and VC components.

In addition to increasing our understanding of the relationship between any membership service and virtually-synchronous communication, this architecture allowed us to specify precisely the FS semantics needed to guarantee VC and MC liveness. One weakness of previous work in this area has been a lack of precise semantics for the FS part of the system. Explicitly, the paper shows:

- that virtually-synchronous communication satisfying the definition given in Section 5.4 can be implemented with a modular, three-component architecture for system models with both process crash failures and network partitions (i.e. link failures). We start with a very simple model, and from it construct a useful communication primitive for fault-tolerant, distributed applications.

- how to define concurrent views that have empty intersections. That is, how to implement strong-partial membership semantics in a system that may partition. The basic idea is to define a view as a set of pairs \((\text{proc id}, \text{proc sequence number})\).
that if we remove the MC component from the architecture (e.g. if virtually-synchronous communication is not needed), then the view component defines views that do not satisfy the empty intersection condition (i.e. giving a membership service with a weak-partial semantics).

Section 2 describes our low-level system model and the interaction of the three components. Section 3 gives a precise semantics for the failure suspector. Sections 4 and 5 sketch how to implement the VCp and MCp components, and Section 6 completes the VCp and MCp protocols. We conclude in Section 7.

2 System Model

Our low-level system model consists of an infinite name space of process identifiers, Proc = \{p1, p2, ..., \}. The name space is infinite to model infinite executions in which processes continually fail and recover. At any point in time, however, there are only a finite number of executing processes under consideration and we restrict our attention to these. For this finite set of executing processes, we assume a completely-connected network of FIFO channels. Processes communicate by passing messages over these channels, though they too may fail. The system has no global clock, and message transmission delays are unbounded. Processes fail by crashing, which we model by the local event \texttt{cras}_p. We model the recovery of a process with a new identifier. A process \( p \) may (1) send a message to another process, (2) deliver a message sent by another process \( q \), and (3) perform local computation.

A \textit{history}, \( h_p \), for process \( p \) is a sequence of events beginning with the event \texttt{start}_p and terminating, if at all, with the event \texttt{crash}_p: \( h_p = \texttt{start}_p \cdot e^1_p \cdot \ldots \cdot e^k_p \), for \( 0 \leq k \). A \textit{cut} is an \( n \)-tuple of process histories, one for each \( p \in \text{Proc} \). We assume familiarity with inter-event causality [15] and with consistent cuts [8].

Crash failures are surprisingly difficult to handle in an asynchronous system. Fischer, et.al [10] show that, because it is impossible to distinguish a crashed process from one that is just very slow, any problem requiring "all correct processes" to agree on some value cannot be solved deterministically; that is, no deterministic protocol can make progress if it must also make accurate process failure detections. One way around this is for asynchronous systems to incorporate some mechanism for \textit{suspecting} failures, as well as a means of handling failure suspicions consistently (e.g. \( p \) may suspect \( q \) faulty while \( r \) may not; perhaps \( r \) and/or \( q \) even suspect \( p \)). Our system model assumes a \textit{failure suspector} that eventually
suspects a crashed process,\textsuperscript{1} which suffices to ensure our protocols make progress. We do not require anything more of the failure suspector.

Each process has three components that interact to implement the virtually-synchronous communication primitive for application-layer processes (Figure 1). The Failure Suspector (\(\text{FS}_p\)) is at the lowest level and notifies both the Multicast Component (\(\text{MC}_p\)), and the View Component (\(\text{VC}_p\)) about suspected changes in the communication topology. Such changes arise from actual process and link failures, as well as high processor loads and heavy network traffic (indistinguishable from true failures). \(\text{VC}_p\) defines \(p\)'s current view, \(\text{View}_p()\), an approximation of the set of processes with which \(p\) can communicate, and sends \(\text{View}_p()\) to \(\text{MC}_p\). \(\text{MC}_p\) is responsible for reliably multicasting application-layer messages until it receives an accessibility-change notification from \(\text{FS}_p\). These notifications signal a suspected change in the communication topology and the attendant need to alter \(\text{View}_p()\). However, neither \(\text{MC}_p\) nor \(\text{VC}_p\) can do this naively since virtually-synchronous communication requires that members of \(\text{View}_p()\) that also accompany \(p\) to its next view receive the same set of messages that were multicast within \(\text{View}_v()\) (We make this definition precise in Section 5). To ensure this, \(\text{MC}_p\) delivers all outstanding multicasts, and does not issue new multicasts except to forward those that have been only partially delivered. \(\text{View}_p()\) is safely terminated when all messages multicast in it are delivered at all sites that \(\text{MC}_p\) believes non-faulty. When \(\text{MC}_p\) detects this condition (Section 4) it informs \(\text{VC}_p\), which then determines a new view for \(\text{MC}_p\) from the accessibility notifications it received from \(\text{FS}_p\).

Section 3 describes the properties our Failure Suspector components must satisfy. These are weak yet reasonable requirements, and are easily implemented in any asynchronous system. Section 4 discusses \(\text{VC}_p\), and Section 5 discusses \(\text{MC}_p\). These components execute protocols

\textsuperscript{1}This can easily be implemented with time-outs.
to detect global properties [8, 16].

3 The Failure Suspector

Given process p, $F_S_p$ emits a sequence of $\text{not-comm}(q)$ and $\text{comm}(r)$ suspicion messages to $MC_p$ and $VC_p$. Since the system is asynchronous we cannot guarantee the accuracy or timeliness of these suspicions; the most we can require is that $F_S_p$ eventually suspects true crashes and recoveries. This is not unreasonable. It is known that fault-tolerant protocols in asynchronous systems cannot make progress if they are required to make accurate failure determinations. Our approach introduces an inaccurate failure suspector to gain liveness. On the other hand, we cannot require $F_S_p$ to suspect all periods of transient inaccessibility – a network partition may repair before it is noticed.

Since, in theory, $F_S_p$ may suspect processes arbitrarily, we have divorced $F_S_p$ implementation from the problem at hand. In a real system, $F_S_p$ might take cues from the underlying communication layer, the operating system, response delays, and so forth.\footnote{For example, to detect failures $F_S_p$ could query a process, deeming it inaccessible if it does not respond in a timely fashion (inaccurate, but satisfying the requirement). We might put the onus on a process to announce its recovery.}

On every consistent cut c, $F_S_p$ maintains two non-intersecting sets, $\text{CommSet}_p(c)$ and $\text{NotCommSet}_p(c)$. When $F_S_p$ suspects $q \in \text{CommSet}_p(c)$, q is removed from $\text{CommSet}_p(c)$ and is thereafter a member of $\text{NotCommSet}_p(c)$. Whenever these sets change, $F_S_p$ notifies $VC_p$ and $MC_p$ by emitting the appropriate $\text{comm}()$ or $\text{not-comm}()$ messages.

We have a reciprocity condition for (perceived) partitions, as well. To model the nature of network partitions, we require eventual reciprocity of inaccessibility suspicions. That is, if $F_S_p$ suspects $q$ then eventually either $F_S_q$ suspects $p$ or $q$ fails.

A logical formula holds on a consistent cut. The membership of an \textit{indexical set} of processes depends on when it is considered. In our model, 'when' translates to consistent cuts, the only physically-realizable instances. We use the following formulas and indexical sets to specify the behavior of $F_S_p$.

- $\text{NOTCOMM}_p(q)$ holds on $c$ if $q \in \text{NotCommSet}_p(c)$

- $\text{COMM}_p(q)$ holds along $c$ if $q \in \text{CommSet}_p(c)$

- $\text{DOWN}_q$ holds along $c = (h_1, \ldots, h_q, \ldots, h_n)$ if $\text{crash}_q$ is the last event in $h_q$
• UP<sub>q</sub> holds along \( c = (h_1, \ldots, h_q, \ldots, h_n) \) if crash<sub>q</sub> is not an event in \( h_q \).

Non-triviality Conditions for FS<sub>p</sub>

**Crashes** If \( q \) crashes, then eventually either \( p \) crashes or FS<sub>p</sub> suspects \( q \) is unreachable:

\[
\text{DOWN}_q \Rightarrow \Diamond \left( \text{NOTCOMM}_p(q) \lor \text{DOWN}_p \right)
\]

**Recoveries** If \( q \) begins executing and is reachable, then eventually either \( p \) crashes or FS<sub>p</sub> suspects \( q \) is reachable:

\[
\text{UP}_q \Rightarrow \Diamond \left( \text{COMM}_p(q) \lor \text{DOWN}_p \right)
\]

**Reciprocity** If FS<sub>p</sub> suspects \( q \) is inaccessible, then, if \( q \) does not crash, it eventually suspects \( p \) is inaccessible:

\[
\text{NOTCOMM}_p(q) \Rightarrow \Diamond \left( \text{DOWN}_q \lor \text{NOTCOMM}_q(p) \right)
\]

This is an artifact of \( p \) suspecting \( q \): since \( p \) ceases communicating with \( q \), \( p \) is, in fact, inaccessible to \( q \).

Propagation Conditions for FS<sub>p</sub>

Finally, we require failure suspectors to *gossip* among themselves.

**Inaccessibility Propagation** If FS<sub>p</sub> believes, on cut \( c \), it cannot communicate with \( q \) then it tries to propagate this belief to every FS<sub>r</sub> for \( r \in \text{CommSet}_p(c) \):

\[
\text{NOTCOMM}_p(q) \Rightarrow \Diamond \left( \text{NOTCOMM}_r(q) \lor \text{NOTCOMM}_r(p) \right)
\]

**Accessibility Propagation** If FS<sub>p</sub> believes, along \( c \), it can communicate with \( q \) then it tries to propagate this belief to every FS<sub>r</sub> for \( r \in \text{CommSet}_p(c) \):

\[
\text{COMM}_p(q) \Rightarrow \Diamond \left( \text{COMM}_r(q) \lor \text{NOTCOMM}_r(p) \right)
\]
3.1 Related work

Before discussing the other components, we discuss the relation between this and other work. In [7], Chandra and Toueg solve Distributed Consensus in an asynchronous system using a Failure Suspector, $W$, that satisfies certain (weak) requirements. [6] further shows that $W$ is the weakest suspector that can be used to solve Distributed Consensus. While we do not consider consensus in this paper we said in the Introduction that adding a majority requirement to the VSC abstraction, gives a simple, probabilistic solution to transaction commit.

Since there are no fundamental differences between solving consensus and atomic commit problem, how are both approaches related (we will not, hereafter, distinguish consensus from atomic commit)?

First it should be clear that our Failure Suspector is not weaker than $W$. More important, [7] also places a majority requirement on processes before $W$ can be used to solve consensus. To relate the two approaches, consider a generalization of consensus:

- suppose consensus is to be solved more than once, and let $consensus(i)$, for $i > 0$, be the $i^{th}$ instance of the consensus problem;
- let Proc be the initial set of processes that solve $consensus(1)$;
- $consensus(i + 1)$ begins only after $consensus(i)$ has been solved;
- for $consensus(i)$, $i > 1$, the processes chose their initial state randomly from the set $\{0, 1\}$.

In [7], $consensus(i)$ (for each $i$) would be solved by the same static set of processes Proc. The majority requirement to solve $consensus(i)$ is thus similar to a static voting scheme in the context of handling replicated data [11]. This is because [7] consider that failure suspicions are never stable: a process $p$ believing $failed(q)$ can always change its mind.

In contrast, in the VSC model, failure beliefs are stable each time a new view is defined. Thus for $i \neq j$, $consensus(i)$ and $consensus(j)$ need not be solved by the same set of processes. Continuing the replicated data analogy, the majority requirement in the VSC model is similar to the dynamic voting scheme [9], which has been shown to lead to higher data availability than the static voting scheme.
4 The View Component

The view component operates whenever a link failure repairs, a process begins executing, recovers after a crash, and whenever the multicast component informs it that the current view has terminated (Section 5). \( \text{VC}_p \) defines p's current view by interaction with other \( \text{VC} \) components, and by using \( \text{FS}_p \) information.

\( \text{VC}_p \) defines a new view when it detects (or learns about through some other \( \text{VC} \) component) agreement on \( \text{CommSet}_p() \) among the members of \( \text{CommSet}_p() \). The new view will be the largest subset of processes (containing p) satisfying this agreement.

4.1 The View Component Algorithm

In this section we outline how \( \text{VC}_p \) detects or learns about \( \text{CommSet}_p() \) agreement.

When \( \text{VC}_p \) is activated, it knows a near approximation of \( \text{CommSet}_p() \) from \( \text{FS}_p \).\(^3\) Whenever \( \text{VC}_p \) receives an \text{comm}(r) message from \( \text{FS}_p \), it updates this approximation. Along cut \( c \), \( \text{VC}_p \) uses a deterministic function, \( \text{vc-Coo}(p) \),\(^4\) on the set \( \text{CommSet}_p(c) \) which returns a unique process identifier, and satisfies

\[
\left( \text{CommSet}_p(c) = \text{CommSet}_q(c) \right) \Rightarrow \left( \text{vc-Coo}(p) = \text{vc-Coo}(q) \right).
\]

For example, \( \text{vc-Coo}(p) \) might be "choose the 'smallest' identifier from \( \text{CommSet}_p(c) \)."

Each process also maintains a local counter, \( \text{seq}_p \), which is incremented every time \( \text{vc}_p \) considers \( \text{vc-Coo}(p) \) to have changed (this is not necessarily every time \( \text{CommSet}_p(c) \) changes. For liveness, however, \( \text{vc-Coo}(p) \) must change when \( \text{VC}_p \) receives \text{not-comm}(\text{vc-Coo}(p)) \) from \( \text{FS}_p \). The counter \( \text{seq}_p \) is initially zero and is essential in allowing us to define non-intersecting, concurrent views. The tuple \( (p, \text{seq}_p) \) fully describes \( p \) on any consistent cut.

Finally, the formula \( \text{COMMSETEQ}(S) \) holds on \( c \) if and only if all \( p \in S \) have identical \( \text{CommSet()} \) sets at \( c \). That is,

\[
\text{COMMSETEQ}(S) \overset{\text{def}}{=} \bigwedge_{p,q \in S} \left( \text{CommSet}_p() = \text{CommSet}_q() \right)
\]

\(^3\)There may be notifications from \( \text{FS}_p \) that have not yet reached \( \text{VC}_p \).

\(^4\)Technically, we should name some cut explicitly since the function's value depends \( p \)'s indexical communicate-with set. We omit the cut reference, but with the understanding that \( \text{vc-Coo}(p) \) has a temporal dependence. In fact \( p \) never knows which particular cut it is on, but at any point in its execution \( \text{VC}_p \) has some set of process identifiers that satisfy a certain condition. It determines a coordinator by applying some rule to this set. The presence of \( c \) would only clarify matters for the omniscient reasoner.
In our protocol, each $p$ sends its current $\text{CommSet}_p()$ and current $\text{seq}_p$ number to $\text{vc-Cooord}(p)$ every time $\text{CommSet}_p()$ changes.

4.2 Defining the New View

Let $\kappa = \text{vc-Cooord}(p)$, and $S = \text{CommSet}_\kappa(c)$ for some cut $c$. Then $\text{vc}_\kappa$ receives $\text{CommSet}_p()$ for $p \in S$. Whenever it receives a different $\text{CommSet}_p()$ from some $p$, $\text{vc}_\kappa$ discards the previous one and checks whether $\text{COMMSETEQ}(\text{CommSet}_\kappa())$ holds. If it does, $\text{vc}_\kappa$ sets the new view, $\text{View}_\kappa()$, to

$$\text{View}_\kappa() = V = \{(p, \text{seq}_p) \mid p \in \text{CommSet}_\kappa()\} \quad (1)$$

The coordinator $\kappa$ then sends the new view to each $\text{vc}_p$ (for $p \in V$) which then delivers the view to $\text{MC}_p$. $\text{MC}_p$ regains execution control and begins multicasting again. Unfortunately, as $\text{COMMSETEQ}(\text{CommSet}_\kappa())$ is not a stable property (i.e. once true, forever true) we must take care in announcing the new view. We return to this issue in Section 6.

4.3 The Partial Order

Correctness of $\text{vc}_p$ means that the coordinator successfully sends the new view to the $\text{vc}$ components of all reachable members in the new view. We will henceforth use $V$ to denote the (local) view that is agreed-upon by all the members of $V$.

Since process histories are linear, it makes sense to talk about the $x^{th}$ version of a process’s (local) view – we denote this by $\text{View}_p^x$.

**Definition** Given two agreement views $V$ and $V'$, $V \prec_I V'$ if and only if there is some $p$ in $V \cap V'$ such that $V = \text{View}_p^x$, and $V' = \text{View}_p^{x+1}$. The transitive closure of $\prec_I$ is denoted $\prec$.

It is not hard to see that the views defined by the collection of $\text{vc}_p$ components are partially ordered by $\prec$. We say $V$ and $V'$ are concurrent if and only if they are not $\prec$-related.

Proposition 4.1 trivially follows from the definition of views (Equation 1) and the increment rule for $\text{seq}_p$.

**Proposition 4.1** Let $V$ and $V'$ be concurrent views. Then $V \cap V' = \emptyset$. 

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5 The Multicast Component

The Multicast Component of process \( p \), \( MC_p \), is responsible for implementing virtually-synchronous communication. \( MC_p \) operates in two modes. In one mode it multicasts messages to the members of its current view \( \text{View}_p() \). In the other mode, it flushes outstanding multicasts to ensure they satisfy virtually-synchronous communication semantics, then terminates the current view. The transition from multicast mode to termination mode is triggered by any \( FS_p \not\text{-comm()} \) or \( \text{comm()} \) message. In this section, we define VSC semantics and the protocols \( MC_p \) uses.

5.1 Definitions

Informally, virtually-synchronous communication is such that, for any view \( V \), the processes of view \( V \) that mutually believe each other alive deliver the same set of multicasts.\(^5\) To make the definition of VSC precise we need to define formally the set of messages considered to have been multicast in \( V \), as well as the subset of processes that deliver them.

**Definition** Given a view \( V \), message \( m \) is a \( V \)-multicast if it was sent by some \( p \) along a cut \( c \) such that \( \text{View}_p(c) = V \).

**Definition (VSC)** Let \( V \prec_i V' \). Then communication in a system is virtually-synchronous if and only if all processes in \( V \) and in \( V' \) delivered the same set of \( V \)-multicasts. Moreover no message is delivered in more than one view.

It is important to notice that process sequence numbers are not used in the definition. These are low-level pieces of information; the application layer should only be concerned with process identifiers. For an application-layer process, VSC ensures two processes that if they progress together from one view to another, then they delivered the same set of messages in the first view. As a result, if process state is determined by an initial state and the set of multicasts delivered to the process, VSC means that if processes begin executing in view \( V \) in the same state, then switch together to view \( V' \), they will begin executing in \( V' \) in the same state.

\(^5\)For simplicity, we omit other forms of communication. Non-multicast communications do not introduce new problems.
5.2 Two Modes of Operation

The component MC\textsubscript{p} operates in two modes:

1. in \textit{normal} mode MC\textsubscript{p} reliably multicasts messages issued by the application layer of \textit{p}, and delivers to the application layer multicasts it receives from other \textit{MCs};

2. in \textit{view-termination} mode MC\textsubscript{p} does not multicast new messages; instead it attempts to flush outstanding multicasts to ensure the VSC semantics.

After receiving a view from VC\textsubscript{p}, MC\textsubscript{p} is in normal mode. It enters view-termination mode as soon as it receives any (in)accessibility notification from FS\textsubscript{p}. When view-termination mode ends, MC\textsubscript{p} gives control back to VC\textsubscript{p}. MC\textsubscript{p} is inactive until it receives a new view from VC\textsubscript{p}, whereupon MC\textsubscript{p} begins normal mode again.

5.3 MC\textsubscript{p} Normal Mode

Suppose VC\textsubscript{p} defines a view \( V = \text{View}_p() \) and delivers this to MC\textsubscript{p}. Recall that views are sets of tuples, which we call process \textit{signatures}:

\[
\text{View}_p() = \{ \sigma_q = (q, \text{seq}_q) \}.
\]

Upon receiving \( \text{View}_p() \), MC\textsubscript{p} enters normal mode, in which it multicasts and delivers messages. Each message \( m \) issued by the application layer of process \( p \) is multicast by MC\textsubscript{p} to all \( q \in V \). Before issuing the message, MC\textsubscript{p} adds \( \sigma_p \) to \( m \). Let \( \text{sender}(m) \) be the signature of the process from which \( m \) originated.

When MC\textsubscript{p} receives a message the following sequence of events occurs:

1. MC\textsubscript{p} delivers \( m \) (to the application layer) if \( \text{sender}(m) \in V \), and discards \( m \) otherwise;

2. MC\textsubscript{p} also buffers any message it receives and delivers in \( V \) until it knows all other processes in \( V \) have received \( m \).\textsuperscript{6} When \( m \) is received by all processes in \( V \) we say it is \textit{stable}.

By delivering only \( V \)-multicasts, the normal mode ensures that no multicast can be delivered in more than one view (see the VSC definition).

\textsuperscript{6}There are many standard ways of achieving this – e.g. piggybacking information on messages.
5.4 $MC_p$ View-Termination Mode

Consider a view $V = \text{View}_p()$. Component $MC_p$ switches from normal mode to view-termination mode after receiving from $FS_p$ either 1) $\text{not-comm}(q)$ for $q \in \text{View}_p()$, or 2) $\text{comm}(r)$ for $r \not\in \text{View}_p()$. This is because whenever a change in the communication topology is detected a new view must be defined reflecting that change. However, before defining a new view, $MC$ in view-termination mode must ensure the VSC definition is satisfied.

Once $MC_p$ enters view-termination mode, it need only consider relevant $\text{not-comm}()$ events from $FS_v$ to terminate $V$. Thus, while executing in view-termination mode, $MC_p$ builds its own approximation of $\text{NotCommSet}_p()$. This means failure notifications have a permanent effect until view-termination mode ends: $\text{comm}(q)$ received by $MC_p$ in view-termination mode after $\text{not-comm}(q)$ (for example due to a partition) cannot undo the $\text{not-comm}(q)$ information.

Just as a new view for $p$ is defined according to agreement on $\text{CommSet}()$s, successfully terminating $V$ involves partitioning $V$ according to $\text{NotCommSet}()$ agreement.

**Definition** The indexical set $\text{Survives}_p(V)$ is $V$ minus the set of processes $MC_p$ believes failed in $V$:

$$\text{Survives}_p(V) = V - \{(q, seq_q) \mid \text{NOTCOMM}_p(q)\}$$

Before we can explain how to ensure VSC, we need the following data structures.

**Definition** Consider $V = \text{View}_p()$ and consistent cut $c$. The vector $msg_p(V, c)$ (of size $|V|$) is defined such that:

- its $p^{th}$ component, $msg_p(V, c)[p]$, is the number of $V$-multicasts that originated from $p$ (up to $c$);
- for $q \in V, q \neq p$, its $q^{th}$ component, $msg_p(V, c)[q]$, is the number of $V$-multicasts $MC_v$ delivered up to $c$ that originated from $q$.

**Definition (View Terminated)** Consider view $V$ and $S$ such that $\emptyset \neq S \subseteq \text{Ids}(V)$ (where $\text{Ids}(V)$ is the set of process identifiers appearing in $V$). Then $vT(V, S)$ holds along cut $c$ if and only if

$$\bigwedge_{p, q \in S} \left( (msg_p(V, c) = msg_q(V, c)) \land (\text{Survives}_p(V, c) = \text{Survives}_q(V, c)) \right)$$

It is not hard to see $S = \text{Ids}(\text{Survives}_p(V))$. ■
In other words $VT(V, S)$ is true exactly when the processes in $S$ agree on both the messages multicast in $V$ and on their respective $Survives(V)$ sets. For $MC_p$, detecting termination of $V = View_p()$ is thus reduced to detecting $VT(V, S)$ (for $p \in S \subseteq Ids(V)$).

Having detected $VT(V, S)$, whether $S = Ids(V)$ or $S \subset Ids(V)$ is important in determining the new view. In the first case, whatever view, $V'$, $VC_p$ later defines, VSC is satisfied with respect to the pair $(V, V')$. In the second case $MC_p$ must pass $Survives_p(V)$ to $VC_p$; we will want the new view to be a subset of $Survives_p(V)$.

To guarantee that every non-crashed process in $V$ eventually detects $VT(V, S)$ for some $S$, $MC_p$ behaves as follows in view-termination mode:

- it stops multicasting new messages;\(^7\)
- it rejects any message $m$ such that $sender(m) \notin Survives_p(V)$.
- upon receiving $not-comm(q)$ from $MC_p$ (for $q \in V$), $MC_p$ signs and forwards any $V$-multicasts originating from $q$ that are still in $p$’s buffer (Section 5.3). $MC_p$ then removes these messages from its buffer. $MC_q$ rejects the re-issued message if $NOTCOMM_q(p)$ holds (i.e. if $MC_q$ has received $not-comm(p)$ from $FS_q$).\(^8\)

**Proposition 5.1** Consider view-termination mode as described above. Then for each $p \in V$, there exists a set, $S_p$ such that $p \in S_p$ and $VT(V, S_p)$ holds.

**Proof** (sketch) We introduce the following notation:

- $VT_1(V, S) \overset{\text{def}}{=} \bigwedge_{p,q \in S} msg_p(V) = msg_q(V)$
- $VT_2(V, S) \overset{\text{def}}{=} \bigwedge_{p,q \in S} Survives_p(V) = Survives_q(V))$

Consider $p \in V$. We build a sequence $S^0_p, \ldots, S^i_p, \ldots, S^n_p$, where $\forall i$, $p \in S^i_p$ and $S^i_p \subseteq Ids(V)$, such that finally $VT(V, S^n_p)$ holds. Initially take $S^0_p = Ids(V)$. The proof ends as soon as $VT(V, S^i_p)$ holds, for some $i$. If not, then $VT_1(V, S^i_p)$ or $VT_2(V, S^i_p)$ does not hold. We obtain $S^{i+1}_p$ from $S^i_p$ by removing a process (if necessary). Because (1) $S^0_p$ is finite, (2) the number of messages sent in a view is finite once view-terminaton mode is started (processes do not issue

\(^7\)If the network were a broadcast domain, $MC_p$ could continue multicasting using a new signature $(p, seq_p + 1)$. The problem for less general environments is that the new multicast view (destination set) is not yet known.

\(^8\)Duplicate messages are recognized and discarded as usual.
new multicasts in this mode), and (3) \( \text{VT}(V, \{p\}) \) is trivially true, the construction finally ends with \( S_p^i \) such that \( \text{VT}(V, S_p^i) \) holds. We briefly discuss the proof reasoning for the case when either \( \text{VT}_1(V, S_p^i) \) or \( \text{VT}_2(V, S_p^i) \) does not hold.

(a) If \( \text{VT}_1(V, S_p^i) \) does not hold, then the message set of some \( q \) in \( S_p^i \) differs from \( p \)'s, in some component: \( \exists q, r \in S_p^i : (msg_p(V)[r] \neq msg_q(V)[r]) \). If eventually these sets become equal, then take \( S_p^{i+1} = S_p^i \). If not (i.e., \( msg_p(V)[r] \) never equals \( msg_q(V)[r] \)), then either \( \text{DOWN}_r \), or \( \text{NOTCOMM}_r(p) \), or \( \text{NOTCOMM}_r(q) \) holds. So suppose \( \text{NOTCOMM}_r(p) \) holds (analogous arguments hold for \( \text{NOTCOMM}_r(q) \) and \( \text{DOWN}_r \)). Then eventually \( \text{NOTCOMM}_p(r) \) holds (from \( \text{FS}_p \) Reciprocity). The Reissuing rule in view-termination mode means that \( p \) will forward to \( q \) all messages it received from \( r \) that \( q \) did not. However, since the message sets never agree this transfer will not succeed completely before \( \text{NOTCOMM}_q(p) \) eventually holds. Reciprocity ensures that \( \text{NOTCOMM}_p(q) \) holds, and at this point we define \( S_p^{i+1} \) to be \( S_p^i - \{q\} \).

(b) If \( \text{VT}_2(V, S_p^i) \) does not hold, then there is some \( q \) in \( S_p^i \) such that \( \text{Survives}_p(V) \neq \text{Survives}_q(V) \). Without loss of generality let \( r \in \text{Survives}_q(V) - \text{Survives}_p(V) \). Then Inaccessibility Propagation and Reciprocity mean that eventually either \( \text{NOTCOMM}_q(r) \), or \( \text{NOTCOMM}_p(q) \) holds. In the first case \( S_p^{i+1} \) to be \( S_p^i - \{q\} \); in the second case, take \( S_p^{i+1} = S_p^i \).

5.5 An Algorithm to Detect \( \text{VT}(V, S) \)

Like the \( \text{VC}_p \) algorithm detecting \( \text{COMMSETEQ}(\)\), the \( \text{MC}_p \) algorithm detecting \( \text{VT}(V, S_p) \) relies on a coordinator process. \( \text{MC}_p \) determines its view-termination coordinator with a deterministic function, \( mc-\text{Coord}(p) \), on the set \( \text{Survives}_p(V, c) \). We require that for \( p \) and \( q \) in \( V \), with identical \( \text{Survives}(V) \) sets, \( mc-\text{Coord}(p) = mc-\text{Coord}(q) \).

Let \( \chi = mc-\text{Coord}(p) \). Then \( \chi \) attempts to detect \( \text{VT}(V, \text{Survives}_\chi(V)) \). \( \text{MC}_p \) also increments the sequence number counter, \( \text{seq}_p \), whenever \( \text{MC}_p \) considers \( mc-\text{Coord}(p) \) to have changed (for liveness, the function \( mc-\text{Coord}(p) \) must change whenever \( \text{MC}_p \) receives \( \text{not-comm}(mc-\text{Coord}(p)) \) from \( \text{FS}_p \)).

Process \( p \) sends \( msg_p(V) \), \( \text{Survives}_p(V) \), and \( \text{seq}_p \) to \( mc-\text{Coord}(p) \) when \( \text{MC}_p \) first considers \( mc-\text{Coord}(p) \) to be its coordinator, and whenever \( msg_p(V) \) and \( \text{Survives}_p(V) \) are modified. If \( \chi = mc-\text{Coord}(p) \), then:

\[
\text{VT}(V, \text{Survives}(V)) \leftrightarrow \bigwedge_{p \in \text{Survives}_\chi(V)} \left( msg_\chi(V) = msg_p(V) \land \text{Survives}_\chi(V) = \text{Survives}_p(V) \right)
\]
Proposition 5.2 Consider a view $V$, with $p \in V$ and the view-termination protocol described above. Then eventually, either $p$ crashes or it detects $\text{vt}(V, \text{Survives}_x(V))$.

Proof (sketch) The proof is similar to that of Proposition 5.1. Here, we consider the perspective of $x = \text{mc-Coo}rd(p)$. The problem is that, due to transmission delays, $x$ may not detect $\text{vt}(V, \text{Survives}_x(V))$ as soon as it holds (transmission of $msg_p(V)$ and $\text{Survives}_p(V)$ from $p$ to $x$).

There are two cases: eventually $x$ receives the messages enabling it to detect $\text{vt}(x, \text{Survives}_x(V))$, or failures prevent $x$ from detecting it. In the second case, if both $\text{COMM}_p(x)$ and $\text{COMM}_x(p)$ hold, we can use the iterative construction, from the perspective of $x$, in the proof of Proposition 5.1. Otherwise we must consider the iterative construction with respect to $x'$, the coordinator replacing $x$ once it is no longer a member of $\text{Survives}_p(V)$. $\blacksquare$

Finally, the fact that $\text{vt}(V, S)$ is not stable poses the same problems as those posed by $\text{COMMS}ETEQ()$’s instability. We consider both in the next section.

6 Instability of COMMSETEQ() and vt(V, S)

As described in the previous sections, once $VC_p$ learns $\text{COMMS}ETEQ(\text{CommSet}_p())$ it switches control to $MC_p$; switching control from $MC_p$ to $VC_p$ is based on detecting $\text{vt}(\text{View}_p(), S)$. In both cases, the relevant property is not stable – it may become false after holding along some cut. Let $\text{switch}(VC, V')$ be the message announcing the new view, $V'$, and $\text{switch}(MC, \text{Survives}())$ be the message announcing termination of view $V$.

Since neither $\text{COMMS}ETEQ(S)$ nor $\text{vt}(V, S)$ are stable properties, we can arrive at the following situation:

- Take $p, q \in V$ such that $p$ and $q$ believe each other accessible, and let $\kappa$ be their mutual $VC$ coordinator ($\kappa = \text{vc-Coo}rd(p) = \text{vc-Coo}rd(q)$). Suppose $VC_\kappa$ determines the new view, $V'$ ($\kappa, p, q \in V'$), sends $\text{switch}(VC, V')$ to $p$ only, and then crashes. $VC_p$, upon receiving $\text{switch}(VC, V')$, adopts $\text{View}_p() = V'$ and switches control to $MC_p$ in normal mode.

- Now suppose that in addition to $VC_q$ not getting $\text{switch}(VC, V')$, $FS_q$ notifies $VC_q$ that $\kappa$ is inaccessible; $q$ continues executing in $VC_q$ waiting for some new coordinator $\kappa'$ to

\footnote{While we illustrate instability with $\text{COMMS}ETEQ()$ and the switch from $VC_p$ to $MC_p$, a similar situation arises for $\text{vt}(V, S)$ as well.}
inform it of the new view. In particular, suppose $\kappa' = p$.

- Since $p$ and $q$ continue to believe each other accessible, $FS_q$ gossips $not\text{-}comm(\kappa)$ to $FS_p$. At this point, $MC_p$ enters view-termination mode for view $View_p() = V'$, and $q$ is still executing in $VC_q$ waiting to receive the successor view to $V$. Observe that unless one of the processes crashes or a network partition splits them, $p$ and $q$ need never believe each other inaccessible.

- For $VC_q$ to make progress, its coordinator $VC_p$ must tell it some new view. Unfortunately, $VC_p$ cannot begin executing until $MC_p$ leaves view-termination mode. $MC_p$ cannot leave view-termination mode until it receives $Survives_q()$ from $MC_q$ (after all, $q \in V'$ and $q \in CommSet_p()$). In other words, $p$ and $q$ are deadlocked because their execution controls are out of phase. The control discrepancy prevents either one ($VC_q$ or $MC_p$) from making progress until one of them believes the other inaccessible – $q$ is stuck in $VC_q$, and $p$ is stuck in $MC_p$.

While processes being out of phase is not always destructive, and in fact is quite natural whenever partitions occur, it is destructive in this case since it induces deadlock. The following precludes deadlock.

### 6.1 Component-Switch Protocol

Let $\kappa$ be shorthand for $vc\text{-}Coord(p)$ when $VC_p$ is executing. We describe the protocol only for the switch from $VC_p$ to $MC_p$; the situation is analogous for the reverse switch. Let $V = View_p()$. We define the following concepts as depicted in Figure 2:

- From Section 4, each accessibility notification from $FS_p$ forces $VC_p$ to inform its coordinator $VC_\kappa$ of the change to $CommSet_p()$. Let $VC\text{-}alert_\kappa()$ denote the message $VC_p$ sends to $VC_\kappa$ to inform $VC_\kappa$ of the change to $CommSet_\kappa()$.

- Let $FS\text{-}VC\text{-}Notify_p(V')$ be the set of $not\text{-}comm(q)$ and $comm(r)$ accessibility notifications $VC_p$ received from $FS_p$ after sending its first $CommSet_p()$ to any coordinator and before receiving $switch(VC, V')$ from $VC_\kappa$;

So given $V'$ and $FS\text{-}VC\text{-}Notify_p(V')$, $VC_p$ can infer which $VC\text{-}alert_\kappa()$ messages reached $VC_\kappa$ before it detected $CommSetEQ(CommSet_\kappa())$ and which did not. Let $FS\text{-}VC\text{-}Late_\kappa$ be the subset of $FS\text{-}VC\text{-}Notify_p(V')$ for which the corresponding $VC\text{-}alert_\kappa()$ message did not reach $VC_\kappa$. 

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Figure 2: $FS-VC-Notify_p(v')$ (lightly-shaded rectangle), $VC-alert_p()$, $FS-VC-Late_p$ (darkly-shaded rectangle)
The Component Switch protocol for \( vcp \) is:

1. The coordinator \( \kappa \) sends the switch(\( vc, V' \)) message using a best effort reliable multicast [14] (a process receiving the message reissues it to all the destination processes).

2. Upon receiving switch(\( vc, V' \)), \( vc_p \):
   a. logically reorders it to be before \( vc_p \) sent any of the messages in \( FS-VC-Late_p \) (this will be clearer after 3);
   b. installs \( V' \) as View\(_p()\) and switches control to MC\(_p\), in normal mode;

3. MC\(_p\) handles messages in \( FS-VC-Late_p \) as if the corresponding notifications from FS\(_p\) had just arrived (i.e. while MC\(_p\) is executing, and not while VC\(_p\) was executing). Specifically, MC\(_p\) simulates receiving these accessibility notifications in View\(_p() = V'\).

Proposition 6.1 The Component-Switch Protocol prevents deadlock.

Proof (sketch) We restrict this discussion to a process \( p \), in view \( V \), switching from its \( vc_p \) to MC\(_p\) component, and suppose \( q \in V \). Suppose \( q \) never switches from \( vc_q \) to MC\(_p\) in view \( V \). We show this does not prevent \( p \) from later switching from MC\(_p\) back to VC\(_p\).

Because \( p \) switches to MC\(_p\) in view \( V \), \( p \) has received the switch(\( vc, V \)) message. By the Component-Switch Protocol, \( p \) has reissued switch(\( vc, V \)) to \( q \). Then either:

1. \( q \) never receives switch(\( vc, V \)), or
2. \( q \) receives switch(\( vc, V \)) after having already switched to MC\(_q\) in view \( V' \), with \( V' \neq V \).

In the first case, NOTCOMM\(_q(p)\) holds eventually. In the second, \( p \in V' \) contradicts \( p \in V \). Thus, NOTCOMM\(_q(p)\) holds, and FS\(_p\) Reciprocity means eventually either \( p \) crashes or NOTCOMM\(_p(q)\) holds. Once NOTCOMM\(_p(q)\) holds, \( p \)'s progress (i.e. switching back to VC\(_p\)) is decoupled from \( q \)'s progress; \( q \) cannot be responsible for blocking \( p \).

7 Concluding Remarks

This paper has shown how to implement virtually-synchronous communication using a three-component architecture for systems that experiences process crash failures and network partitions. The three-component architecture lead us to define a clear semantics for a Failure
Suspector (a necessary part of any live, asynchronous system) that guarantees liveness of the VC and MC components. Clearly defining these semantics allows one to implement the Failure Suspector as a modular tool—distinct from all other components—whose implementation can take advantage of the characteristics of the underlying network.

Considering a membership service in relation to virtually-synchronous communication also lead us to better understand the need for a strong-partial compared to a weak-partial membership service. Specifically, a strong-partial membership service (non-intersecting concurrent views) is naturally related to virtually-synchronous communication. We can understand this in the following way. The MC component must identify the sender of a message by its signature $\sigma_q$ to ensure that no multicast is delivered in more than one view. This led us to define a view as a set of process signatures. Considering the increment conditions of $\text{seq}_p$, two different views $V$ and $V'$ trivially have a non-empty intersection. In other words, by requiring that no multicast be delivered in more than one view, we were led to the partial-strong membership service. However if we remove the MC component, (i.e. if the membership service is only defined by FS and VC, without any reference to communication), then the sequence number $\text{seq}_p$ has no clear justification. In that case, a view is just a set of process identifiers (or a set of identifiers and an incarnation number). With this definition, the same VC protocol we described would define concurrent views that overlap, providing only a weak-partial membership service.

References


