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## Distributed Systems

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# 1

## Introduction

### 1.1 Defining a distributed system

We define a **distributed system** as a collection of processes  $p_1, p_2, \dots, p_n$  that run on different computers and cooperate to solve a problem. The processes communicate using **channels** and we assume the system is fully connected, meaning that every pair of processes can exchange messages between them. The channels are reliable, in the sense that the message arrive, but may delivered messages out of order.

The simple model for a distributed system is called **asynchronous**, that have no upper bound on the speed of processes and no upper bound for the delay of a message. If these upper bounds exists the system is called **synchronous**. The second one is stronger, in the sense that we do more assumptions and this means that every program that run on a synchronous system can run on an asynchronous system (the opposite can be possible but not sure).

A distributed system can have different properties:

- **Consistency**: every part of the system has the same information at every time
- **Availability**: the information are available at every time
- **Partition tolerance**: if one part of the system goes offline the system can continue to run

Every type of distributed system can have up to 2 of this properties.

### 1.2 Computation

We describe the execution of a program on a distributed system as a collection of processes. Every process is defined as a sequence of **events**. The events can be internal or involve communication like the events `send(m)` and `receive(m)`.

We label an event with the notation:

$$e_i^k$$

in which  $i$  represent the index of the process  $p_i$  and  $k$  the order of the event for that specific process.

### Local and global history

The **local history** of a process  $p_i$  is a sequence of events  $h_i = e_i^1 e_i^2 \dots$  that represent the sequential execution of events in the process. We use  $h_i^k$  to represent the sequence of the first  $k$  events in the process  $p_i$ .

The **global history** of the computation is a set that contains all events:

$$H = h_1 \cup h_2 \cup \dots \cup h_n$$

The global history gives us no information about the time in which these events are executed, because in an asynchronous system there is no global clock.

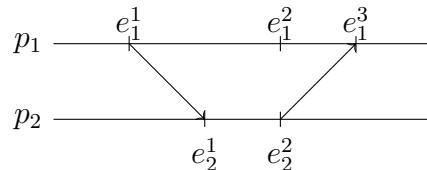
### Relation of cause-effect

Events can be labeled based of the notion of **cause-effect**, defining a relation (with symbol  $\rightarrow$ ) between two events such that:

- $\forall e_i^k, e_i^l \in h_i \wedge k < l \implies e_i^k \rightarrow e_i^l$
- $e_i = \text{send}(m) \wedge e_j = \text{receive}(m) \implies e_i \rightarrow e_j$
- $e \rightarrow e' \wedge e' \rightarrow e'' \implies e \rightarrow e''$  (Transitive)

This relation mean that  $e \rightarrow e'$  if  $e$  causally precedes  $e'$ , so the computation of  $e'$  is influenced by  $e$ . Two events can be unrelated, so neither  $e \rightarrow e'$  nor  $e' \rightarrow e$ . We call this pair of events as **concurrent** and write them as  $e || e'$ .

We can graphically represent a computation with a space-time diagram like this:



An arrow from  $p_1$  to  $p_2$  means that  $p_1$  sends a message to  $p_2$ .

### Run

A **run** is a total order of all events in the global history, consistent with each local history, so the events in history  $h_i$  appear in the same order in  $R$ :

$$R = e_1^1 e_2^1 \dots$$

A single program can have many different runs because some events are unrelated.

## 1.3 Monitoring computations

### Local and global state

We denote  $\sigma_i^k$  the **local state** of a process  $p_i$  after the event  $e_i^k$ .

The **global state** of the computation is an  $n$ -tuple of local states:

$$\Sigma = (\sigma_1, \sigma_2, \dots, \sigma_n)$$

### Cut

A **cut** is a collection of local histories:

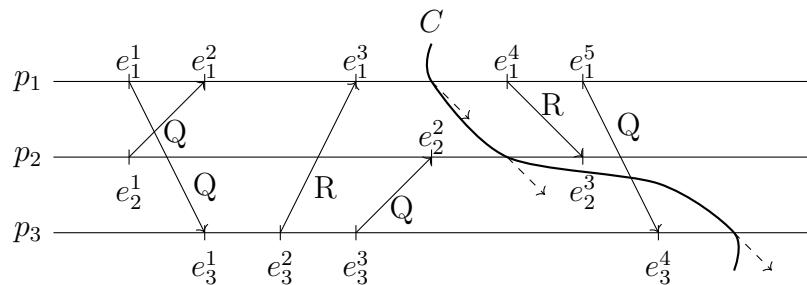
$$C = \langle h_1^{k_1}, h_2^{k_2}, \dots, h_n^{k_n} \rangle$$

To monitor the computation or to compute a global problem (for example knowing if the system is in deadlock) we add a process  $p_0$ .

The first idea is to make  $p_0$  send a message to every process to which a process  $p_i$  will respond with the current state  $\sigma_i$ . After all the response  $p_0$  can construct a global state, that defines a cut.

#### Example:

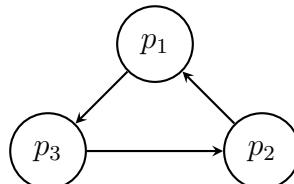
Considering the follows space-time diagram with the cut  $C$  (the dashed arrow represent when the process sends the response to  $p_0$ ):



If we use the responses to create a graph, based on the states, we can say that:

- $p_1$  is going to send a response to  $p_2$
- $p_2$  is going to send a response to  $p_3$
- $p_3$  is going to send a response to  $p_1$

So the graph generated by  $p_0$  will be:



It seems the system has a deadlock, but if we watch carefully we can see that this is not true. The problem is that the global state defined by the cut  $C$  could never happen during a computation.

### Consistent cut

A cut is **consistent** if only if:

$$\forall e \rightarrow e' \wedge e' \in C \implies e \in C$$

So a cut is consistent if the global state generated by the cut could be possible during a computation.

If we use  $p_0$  only to receive messages and we make every process notify the events to  $p_0$  with a timestamp associated,  $p_0$  can reorder the events to create a run. This is true only if there is a global clock, which is not true. To overcome this problem we have to create a local clock for every process and update it in a consistent way.

### Clock condition

A run is consistent if only if follows the **clock condition**:

$$\forall e \rightarrow e' \implies TS(e) < TS(e')$$

If  $TS(e) < TS(e')$  is possible but not guaranteed that  $e \rightarrow e'$ .

### Local clock

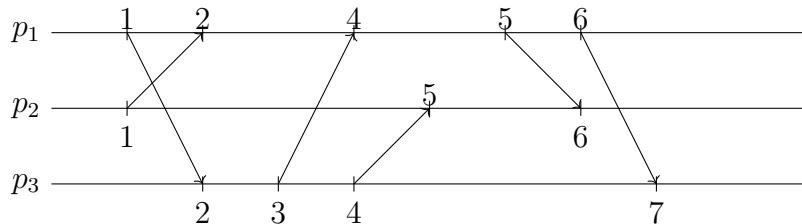
We define the **local clock** of a process as follows:

$$LC(e_i) = \begin{cases} LC + 1 & e_i \text{ is internal or } \texttt{send} \\ \max(LC, TS(m)) + 1 & e_i \text{ is a } \texttt{receive}(m) \end{cases}$$

The local clocks satisfy the clock condition.

### Example:

In the previous example the timestamps of the local clocks will be:



## 1.4 Vector clocks

We want to ensure that the process  $p_0$  sends the messages from the network level to the upper level in order, so the communication between two processes has to be FIFO (First-In First-Out). To do that in a synchronous system with a global clock and an upper bound for a message  $\Delta t$ , we can follow a simple rule for  $p_0$ :

- On a certain time  $t$ , deliver all the messages with timestamp lower than  $t - \Delta t$  in order based on the global clock.

In an asynchronous system we have to decide when to deliver a message, so we have to be sure that all the previous messages are already delivered. We could wait for every other process to send to  $p_0$  every notification with local timestamp lower than the one received before, but this could stuck the system if a process doesn't have enough events to notify.

### Strong clock condition

We expand the definition of clock condition changing the implication with a if only if:

$$e \rightarrow e' \iff TS(e) < TS(e')$$

### History of an event

We define the **history** of an event as a set:

$$H(e) = \{e' | e' \rightarrow e\} \cup \{e\}$$

This includes all the previous events that could have changed the result of event  $e$ .

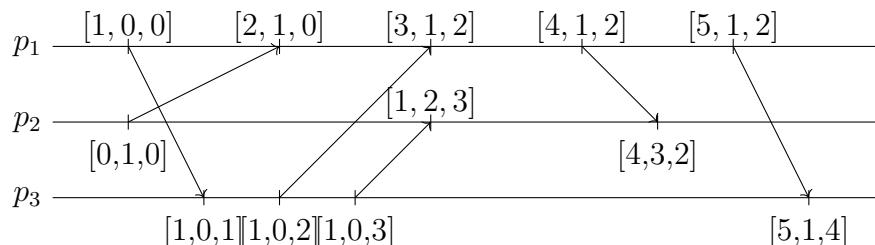
We will identify an history of a process  $e$  with a vector in which every index  $i$  represent the last event of process  $p_{i+1}$  in  $H(e)$ . For example the history  $H(e_1^4) = \{e_1^1, e_1^2, e_1^3, e_1^4, e_2^1, e_3^1, e_2^2, e_3^3\}$  is identified by the vector  $[4, 1, 3]$ . This vector clocks are updated for a process  $p_i$  following the rule:

$$VC = \begin{cases} VC[i] = VC[i] + 1 & \text{always} \\ VC[j] = \max(VC[j], TS(m)[j]) \quad \forall j \neq i & e_i \text{ is a } \texttt{receive}(m) \end{cases}$$

These clocks respect the strong clock condition.

#### Example:

Using this new vector clocks in the previous example the result is:



To know when to deliver a notification  $p_0$  has a counter  $D$  in which  $D[i]$  contains the number of messages delivered from  $p_i$ .  $p_0$  will deliver the notification of an event  $e_i$  if:

$$\begin{aligned} VC(e_i)[i] &= D[i] + 1 \\ VC(e_i)[j] &\leq D[j] \quad \forall j \neq i \end{aligned}$$

### Gap detection

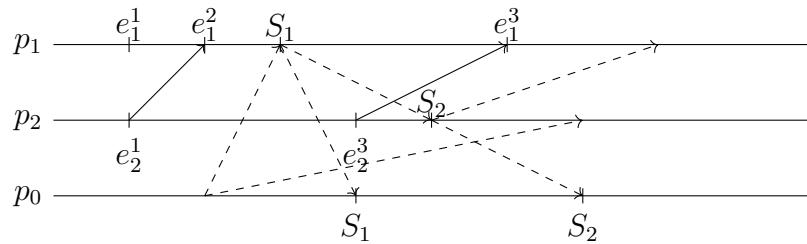
Taken two events  $e_i, e_j$  is possible to know if exists an event between the two:

$$\exists e_k \ e_k \rightarrow e_j \wedge e_k \not\rightarrow e_i \iff VC(e_j)[k] > VC(e_i)[k]$$

## 1.5 Distributed snapshots

If we reuse the idea proposed in Section 1.3 where  $p_0$  sends a message to the other processes and they respond with their state, called **snapshot**. The cut generated by this idea was not surely consistent, so we apply a change.

When a process receive the message "take snapshot" from  $p_0$ , it broadcast his state to the other processes. If another process receive the state of a process it will also broadcast his state. A process will send his state only one time during the protocol, so if he already sent his state to  $p_0$  he will not send it again even if he receives the state of another process. Every process knows the protocol ended when he receives the state of all the other processes. This protocol, called **Chandy-Lampart Protocol** generates a consistent cut if the channels are FIFO.



# 2

## Atomic transaction

### Atomic commit

An atomic commit is the problem that occurs when there is a transaction  $T$  that involves multiple sites in a distributed system. So all the sites do the transaction, or no one does.

We can explain an atomic commit as if every process involved in the transaction votes yes or no to committing the transaction. The system after the votes decide to commit or abort.

An atomic commit has different properties:

- If a process reaches a decision, it must be the same.
- If a process reaches a decision, it cannot change that.
- Decision is to commit only if every process votes yes.
- If there are not failures and all votes are yes the decision must be commit.
- If all failures are fixed then the protocol should terminate.

### 2.1 2 Phase commit

One simple protocol to execute an atomic commit is the **2 Phase commit**, in which there is a coordinator process and there other processes are participants.

The protocol follows 4 phases:

1. The coordinator sends a vote request to every participant.
2. Each participant votes yes or no. If the vote is no the participant can already abort.
3. The coordinator controls the votes and if they are all yes sends a commit message to every participant. Else it sends an abort message.
4. The participants execute the decision received.

The main problems that can occur during this protocol are message not delivered or process that fail during the execution.

To resolve the message failure a timestamp can be used. At every phase the timeout is used in a different way:

1. The participant is waiting for the vote request and reaches the timeout, so it votes no and abort.
2. The coordinator is waiting for the votes and reaches the timeout, so it sends an abort message.
3. The participant is waiting for the decision and reaches the timeout, so it asks other processes for their decision. If they have a decision, it executes the same, if every process is waiting they have to continue waiting.

To resolve the site's failure the processes log during the protocol. Also in this case every phase has a different method to log:

1. The coordinator logs the start of the transaction:

START2PC → DTlog

It is better to log and then send the messages because in the case of a crash in between the two:

- If the log is empty: nothing has been sent, so a transaction can start safely.
- If the log is full: it is safe to resend the request.

2. The participant logs its vote:

yes/no → DTlog

Also in this case it is better to log and then send the message because in the case of a crash in between the two:

- If the log is empty: nothing has been sent, so it is safe to abort.
- If the log is full: it is safe to resend the request.

3. The coordinator logs the votes received and after it sends the decision.
4. The participant logs the decision before executing it.

In some distributed systems with a lot of servers with the same data, this protocol is not the best, because if even one server fails the transaction is aborted.

## 2.2 Paxos

The **Paxos** protocol has the goal to make a system work even in presence of failures. The servers are divided into three categories:

- **Proposer**: starts the transaction and proposes the value to vote.
- **Acceptors**: vote the value.
- **Learners**: keep track of the decision made by the Acceptors.

The protocol has two additional rules:

- All the servers need to choose the same value at the end of the protocol.
- The value chosen must be suggested from outside.

The protocol works with minimum one Proposer and one Learner. For the Acceptors the majority is needed so the maximum number of failures tolerated is with  $n$  Acceptors:

$$f = \left\lfloor \frac{n - 1}{2} \right\rfloor$$

And the size of the quorum will be:

$$|Q| = n - f$$

The protocol works in **rounds** and every round is associated with only one transaction. A round works in this way:

1. Proposer sends a message to all the Acceptors:

$$P \rightarrow A \quad \text{prepare(curr-round)}$$

2. Acceptors respond to the Proposer:

$$A \rightarrow P \quad \text{promise(curr-round, last-round, last-value)}$$

In which they promise to participate in the current round and to not participate to any round lower than that. In addition it also sends the last round in which it has voted and the last value voted.

3. Proposer waits for a quorum of promises and after sends to every Acceptor a message with the value proposed  $x$ , that is chosen in this way:

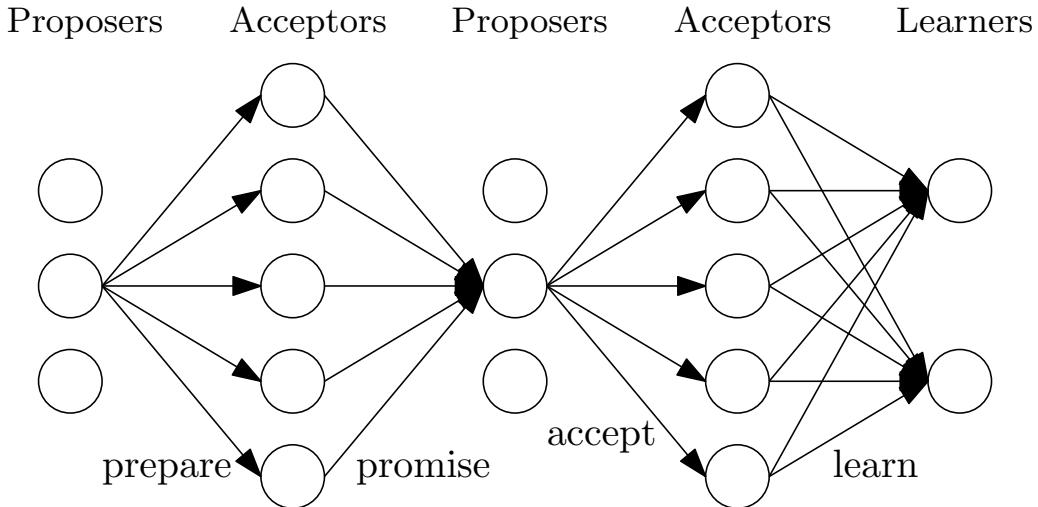
- If no Acceptor has ever voted the value  $x$  can be chosen at random (value that comes from the client).
- Else the value  $x$  is the last value voted associated with the largest value of last-round in the promises.

$$P \rightarrow A \quad \text{accept(curr-round, } x)$$

4. Acceptors send their vote to the Learners:

$$A \rightarrow L \quad \text{learn(curr-round, } x)$$

5. Learners control the votes and if a quorum voted the same value in the same round the value is chosen. If there is no quorum after a timeout the Proposer will ask the Learners for a decision, if there is no decision there will be another round.



There are some problems with transaction that could start in the same time. If the prepare message of a transaction with a lower round arrives after one with an higher round the first transaction will never be done. On the other hand if the prepare message of a transaction with an higher round arrives before the vote of a transaction with a lower round this transaction will be stopped after the acceptors promise with the higher round. This is the reason Paxos is not **live**, in the sense that if this continue to occur the protocol doesn't do nothing.

This protocol is **safe**, means that doesn't do anything wrong even in presence of more then  $f$  failures.

### Safety condition

For Paxos the safety condition is:

Is an acceptor  $A_k$  votes for a value  $x$  in round  $i$ , no value different from  $x$  can be chosen by learners in a previous round.

### Proof:

The demonstration can be done by induction:

1. Base case: round 1 doesn't have previous rounds so the property is valid.
2. Induction ipothesis: for every round  $i$ , the value to vote is equal to the value  $x$  associated with the max last round  $j$  in the promises. If for  $j$  is valid the property, also is valid for  $i$ .
3. Induction step: the value voted in round  $i$  must be voted by some acceptors in round  $j$ , so in round  $j$  is impossible to have a quorum on a value different than  $x$ . All the acceptors that sent promises in round  $i$  don't vote in rounds between  $j$  and  $i$  so any of this rounds can have a quorum of votes.

### 2.2.1 Fast-Paxos

The **Fast-Paxos** protocol is a different version of Paxos, in which there is a Coordinator in the Proposers that prepare the first round before the value to vote arrives. A round works like this:

- Coordinator sends a message to all the Acceptors:

$C \rightarrow A \quad \text{prepare(curr-round)}$

- Acceptors respond to the Coordinator:

$A \rightarrow C \quad \text{promise(curr-round, last-round, last-value)}$

In which they promise to participate in the current round and to not participate to any round lower than that. In addition it also sends the last round in which it has voted and the last value voted.

- Coordinator waits for a quorum of promises and after sends to every Acceptor a message that depends on the promises:

- If no Acceptor has ever voted the Coordinator sends an acceptany message to tell the Acceptors to accept values from any Proposer.

$C \rightarrow A \quad \text{acceptany(curr-round)}$

- Else the value  $x$  is the most voted value in the max last round.

$P \rightarrow A \quad \text{accept(curr-round, } x\text{)}$

- When a value arrives to the Proposers they send an accept message to the Acceptors with the value  $x$ .

$P \rightarrow A \quad \text{accept(curr-round, } x\text{)}$

- Acceptors sends their vote to the Learners:

$A \rightarrow L \quad \text{learn(curr-round, } x\text{)}$

- Learners control the votes and if a quorum voted the same value in the same round the value is chosen. If there is no quorum after a timeout the Coordinator will start another round.

In a round there could be different values voted, to overcome this problem the quorum is modified to be equal to  $\frac{2}{3} + 1$ . So, the maximum number of failures tolerated decrease to:

$$f = \left\lfloor \frac{n - 1}{3} \right\rfloor$$

The protocol Fast-Paxos is safe. It also works with every simple protocol for leader election.

#### Proof:

The demonstration can be done by induction:

- Base case: round 1 doesn't have previous rounds so the property is valid.
- Induction hypothesis: for every round  $i$ , the value to vote is equal to the value  $x$  most common in the max last round  $j$  in the promises. If for  $j$  is valid the property, also is valid for  $i$ .
- Induction step: the value voted in round  $i$  is the most common value voted in round  $j$ , and is the only value that can reach a quorum in round  $j$ , because only 1 value every round has the possibility to have a quorum of votes. This is because with a quorum of  $\frac{2}{3}$  only 1 value can have more than  $\frac{1}{3}$  votes in the promises that could result in a quorum if the  $\frac{1}{3}$  outside the quorum all voted for that value. All the acceptors that sent promises in round  $i$  don't vote in rounds between  $j$  and  $i$  so any of this rounds can have a quorum of votes.

### 2.2.2 Multi-Paxos

Normally the problem is not only to agree on a single value, but on a series of values and on their order. This can be done with the protocol **Multi-Paxos** in which a lot of instances of Paxos (or Fast-Paxos) are run at the same time. Every instance agree on a value and the order in which the transactions are executed is the order of the instances.

The prepare message now contains also the instance of Paxos and multiple instances can be prepared at the same time, for example for  $n$  instances:

$C \rightarrow A \quad \text{prepare}(1-n, \text{curr-round})$

Also Acceptors can promise to multiple instances:

$A \rightarrow C \quad \text{promise}(1-n, \text{curr-round}, \text{last-round}, \text{last-value})$

And the acceptany message can be sent for multiple instances:

$C \rightarrow A \quad \text{acceptany}(1-n, \text{curr-round})$

The accept messages have to be sent for every instance singularly and the instance to send a particular value is chosen by the Proposers. If the Prosopers are also Learners they know which instances have already chosen a value and can send the new values to other instances.

## 2.3 RAFT

This protocol uses leader election and is designed to make consensus on a sequence of values. Every server has a log with all the decisions taken over time.

The protocol works in **terms**, and a term is a period in which there is the same leader. In each term the leader election works as follows:

1. At the start everyone is a follower.
2. Someone becomes a candidate, so wants to become the leader, and send a message to the other servers asking for votes.
3. The other followers respond with a vote.
4. If a quorum of votes ( $50\%+1$ ) arrives the candidate become the leader for that term.

To choose how a server become a leader every server have a random timeout each term and after that it becomes a candidate. To resolve problems in case of more candidates another timeout is set (higher than the upper bound of the random one) and after this time if there is no leader the server goes to the successive term. This is the reason why this protocol is not live. In case a leader is chosen, the term is terminated when something bad happen to the leader, so in the ideal case there is only one term.

After choosing the leader, only it can write in the logs, so every transaction is sent to it. A transaction works in this way:

1. The leader write the transaction in its log with the term number.
2. The leader sends the transaction to every other server.

3. The followers write the transaction in their log and respond to the leader with an OK message.
4. If the leader receive a quorum of OK, it marks the transaction as committed and can be executed.

When a server that is not the leader receive a transaction, it marks all the previous as committed. If there is no next transaction the leader will send an empty one only to commit the last.

RAFT has some properties:

1. Election safety: there is at most one leader per term.
2. Committed entries are always in the log of the leader.
3. If two logs contain a value for the same index and term, it's the same value.

To ensure the property 2 there is a clause on the votes: a server can vote only for another server that has at least the same entries in the log.

## 2.4 Ben-Or

The **Ben-Or** protocol is a randomized protocol that permits to solve consensus and leader election. It's not used in practice but is important from a theoretical point of view.

The protocol has some properties:

- Agreement: all the processes choose the same value (0 or 1)
- Termination: the protocol terminates at some point
- Validity: the value chosen is one of the input

The number of fault tolerated by the protocol is:

$$f = \left\lfloor \frac{n - 1}{2} \right\rfloor$$

Each node follows this algorithm:

**Algorithm:** Ben-Or

```

def Ben-Or():
    preferences = input
    round = 1
    while true :
        send(1,round,preference)// broadcast of type 1
        wait for a quorum of messages of type 1
        if (messages with a specific value  $v$ ) $> \frac{n}{2}$  :
            | send(2,round,v,ratify)// broadcast of type 2
        else :
            | send(2,round,?)
        wait for a quorum of messages of type 2
        if received a message with ratify  $v$  :
            | preference =  $v$ 
            | if (messages with ratify  $v$ ) $> f$  :
                | | return  $v$ 
        else :
            | | preference=random(0,1)
        round+=1
    
```

We assume retransmission, so every messages will arrive at a certain point.

We are sure that:

- If exist two messages:

$$\left. \begin{array}{l} (2, r, v, \text{ratify}) \\ (2, r, v', \text{ratify}) \end{array} \right\} \implies v = v'$$

- If a process receive more than  $f$  messages with ratify  $v$ , then every process has received at least one of them. So, in the next round they are gonna agree and is safe to output the value  $v$ .
- In the worst case scenario (with exactly  $f$  failures) all process need to have the same value to have a quorum. The value is random between 0 and 1 so the probability to exit the protocol on a given round is:

$$P = \frac{1}{2^{n-f-1}}$$

So the expected number of rounds needed is  $2^{n-f-1}$ . Theoretically the protocol can't continue for infinite time and at some point it will terminate, making it live.

## 2.5 Failure detectors

Inside every process there is a failure detector  $D$  that the process can query to know if another process has failed.

Failures detectors are classified by 2 properties:

- **Completeness:** if a process crashes then  $D$  can see it
- **Accuracy:** if  $D$  says that a process is dead, is true

We define  $\text{crash}(\sigma)$  the set of process crushed during the run  $\sigma$  and  $\text{Up}(\sigma)$  the set of process not crushed during run  $\sigma$ . Also  $D_q$  represent the failure detector of process  $q$ .

The completeness can be of two types:

- Strong completeness:

$$\forall \sigma \forall p \in \text{crash}(\sigma) \forall q \in \text{Up}(\sigma) \exists t | \forall t' > t \ p \in D_q(t', \sigma)$$

- Weak completeness:

$$\forall \sigma \forall p \in \text{crash}(\sigma) \exists q \in \text{Up}(\sigma) \exists t | \forall t' > t \ p \in D_q(t', \sigma)$$

The accuracy can be of 4 types:

- Strong accuracy:

$$\forall \sigma \forall t \forall p, q \in \text{Up}(\sigma) \ p \notin D_q(t, \sigma)$$

- Weak accuracy:

$$\forall \sigma \exists p \in \text{Up}(\sigma) | \forall t \forall q \in \text{Up}(\sigma) \ p \notin D_q(t, \sigma)$$

- Eventual strong accuracy:

$$\forall \sigma \exists t | \forall t' > t \forall p, q \in \text{Up}(\sigma) \ p \notin D_q(t', \sigma)$$

- Eventual weak accuracy:

$$\forall \sigma \exists t | \forall t' > t \exists p \in \text{Up}(\sigma) | \forall q \in \text{Up}(\sigma) \ p \notin D_q(t', \sigma)$$

We can define a taxonomy of failure detectors based on this properties:

Accuracy → Completeness ↓	Strong (S)	Weak (W)	Eventual strong ( $\diamond S$ )	Eventual weak ( $\diamond W$ )
Strong (S)	Perfect (P)	Strong (S)	Eventual perfect ( $\diamond P$ )	Eventual strong ( $\diamond S$ )
Weak (W)	Theta ( $\theta$ )	Weak (W)	Eventual theta ( $\diamond \theta$ )	Eventual weak ( $\diamond W$ )

With a P or  $\diamond P$  detector eventually there will only be a leader, so we could make Paxos live. This type of detectors don't exist in distributed systems.

Given a detector with weak completeness, we can create a strong complete one making the process  $q$  that knows that  $p$  is failed broadcast the information so that everyone will also know. So, detector with weak completeness don't really exist, making the only possible detectors being the Strong and Eventual Strong.

# E

## Exercises

### E.1 Exercises on cuts and global states

#### E.1.1 Exercise 1

Let  $C_1$  and  $C_2$  be two consistent cuts. Show that the intersection of  $C_1$  and  $C_2$  is a consistent cut.

$$\begin{aligned} & \left\{ \begin{array}{l} C_1 \text{ consistent} \implies \forall e \in C_1 \wedge e' \rightarrow e \implies e' \in C_1 \\ C_2 \text{ consistent} \implies \forall e \in C_2 \wedge e' \rightarrow e \implies e' \in C_2 \end{array} \right. \\ & \forall e \in (C_1 \cap C_2) \wedge e' \rightarrow e \implies e \in C_1 \wedge e \in C_2 \xrightarrow{\text{consistency}} e' \in C_1 \wedge e' \in C_2 \\ & \implies e' \in (C_1 \cap C_2) \implies (C_1 \cap C_2) \text{ consistent} \end{aligned}$$

#### E.1.2 Exercise 2

Let  $C_1$  and  $C_2$  be two consistent cuts. Show that the union of  $C_1$  and  $C_2$  is a consistent cut.

$$\begin{aligned} & \left\{ \begin{array}{l} C_1 \text{ consistent} \implies \forall e \in C_1 \wedge e' \rightarrow e \implies e' \in C_1 \\ C_2 \text{ consistent} \implies \forall e \in C_2 \wedge e' \rightarrow e \implies e' \in C_2 \end{array} \right. \\ & \forall e \in (C_1 \cup C_2) \wedge e' \rightarrow e \implies e \in C_1 \vee e \in C_2 \xrightarrow{\text{consistency}} e' \in C_1 \vee e' \in C_2 \\ & \implies e' \in (C_1 \cup C_2) \implies (C_1 \cup C_2) \text{ consistent} \end{aligned}$$

#### E.1.3 Exercise 3

Show that every consistent global state can be reached by some consistent run.

Taken the run  $R = e_1, \dots, e_n$  where all the event  $e \in \Sigma$  are at the start in the same order as they appear in  $\Sigma$ ,  $R$  is consistent and reaches the global state  $\Sigma$ .

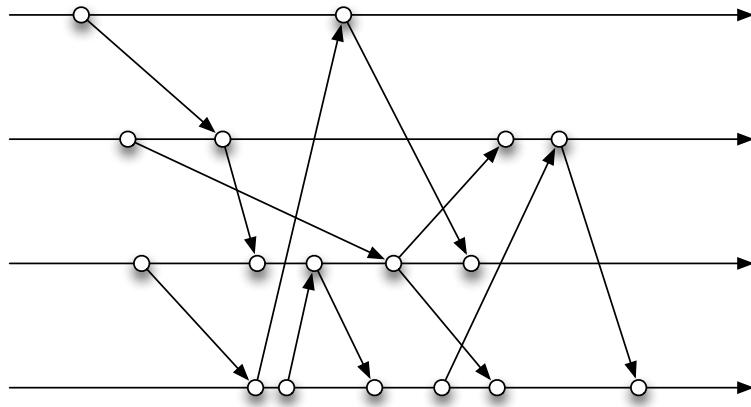
#### E.1.4 Exercise 4

Let  $C_1$  and  $C_2$  be two consistent cuts. If  $C_1$  is a subset of  $C_2$ , then  $C_2$  is reachable from  $C_1$ . (There exists a consistent run that reaches  $C_1$  and then reaches  $C_2$ ).

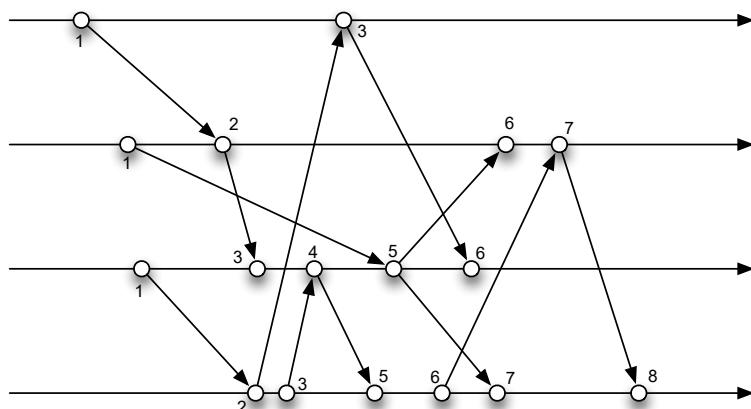
Taken the run  $R = e_1, \dots, e_n$  where all the event  $e \in C_1$  are at the start in the same order as they appear in  $C_1$  and after them there are all the event  $e \in C_2$  in the same order as they appear in  $C_2$ ,  $R$  is consistent and reaches the cut  $C_2$  passing through  $C_1$ .

### E.1.5 Exercise 5

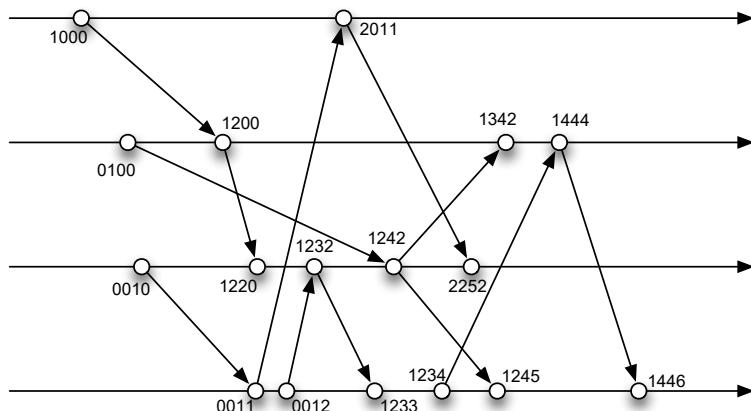
Label with proper logical clock all the events of the distributed computation in the image below. (You can consider events that receive a message and immediately send it as single events).



Logical clocks:



Vector clocks:



## E.2 Exercises on distributed snapshots

### E.2.1 Exercise 6

Show that the Chandy-Lamport Snapshot Protocol builds a consistent global state.

For every events  $e'_i \rightarrow e_j$ , if  $e \in S$  it means that the process  $p_i$  must have taken the snapshot after  $e'$ , because the channels are FIFO. So if the snapshot message from  $p_i$  to  $p_j$  is arrived after  $e$  it means it was sent after  $e'$  and  $e' \in S$ .

### E.2.2 Exercise 7

Show that the Chandy-Lamport Snapshot Protocol can build a global state that never happened.

An event  $e_i$  could happen after the snapshot taken by  $p_i$  but before an event  $e_j$  taken in the snapshot of  $p_j$  (event  $e_i$  and  $e_j$  must be concurrent).

### E.2.3 Exercise 8

What good is a distributed snapshot when the system was never in the state represented by the distributed snapshot? Give an application of distributed snapshots.

Even if the system was never in the state  $S$  represented by the sanpshot, the only difference between  $S$  and the real state happened are concurrent events. So, every check on the system, like deadlock detection works the same in the two cases.

### E.2.4 Exercise 9

Consider a distributed system where every node has its physical clock and all physical clocks are perfectly synchronized. Give an algorithm to record global state assuming the communication network is reliable. (Note that your algorithm should be simpler than the Chandy-Lamport algorithm).

If there is a global clock the monitor can send at time  $t$  a "take snapshot" message with a timestamp  $t + \Delta t$  where  $\Delta t$  is the maximum time required by a message to arrive. When a process recieves the message takes a snapshot at timestamp  $t + \Delta t$  and sends it to the monitor.

### E.2.5 Exercise 10

What modifications should be done to the Chandy-Lamport snapshot protocol so that it records a strongly consistent snapshot (i.e., all channel states are recorded empty).

When broadcasting its snapshot every process sends a flag to signal if its channels are empty. When a process recieves all the snapshot from other processes, if all signaled that they have empty channels te process sends its snapshot to the monitor.

### E.2.6 Exercise 11

Show that, if channels are not FIFO, then Chandy-Lamport snapshot algorithm does not work. A possible problem is the following:

There are two events  $e'_i \rightarrow e_j$ . The process  $p_i$  takes a snapshot in which  $e' \notin S_i$  and after the events  $e'$  occur sending a message. If the channels are not FIFO the message of event  $e'$  could arrive before the snapshot message, causing  $p_j$  to take a snapshot in which  $e \in S_j$ , so the snapshot is inconsistent.

### E.2.7 Exercise 12

Let  $S_0$  be the global state when the Chandy-Lamport snapshot protocol starts,  $S$  be the global state built by the protocol, and  $S_1$  be the global state when the protocol ends. Show that  $S$  is reachable from  $S_0$  and that  $S_1$  is reachable from  $S$ . Remember that  $S$  might not have happened.

From the run  $R_0$  that reaches  $S_0$  can be added all the events  $e \in (S - S_0)$  in the same order in which they appear in  $S_0$ , creating a consistent run  $R$  that reaches  $S$  from  $S_0$ . The same can be done with the events  $e \in (S_1 - S)$ , creating a consistent run that reaches  $S_1$  from  $S$ .

## E.3 Exercises on atomic commit

### E.3.1 Exercise 13

Give an ACP that also satisfies the converse of condition AC3. That is, if all processes vote Yes, then the decision must be Commit. Why is it not a good idea to enforce this condition?

An ACP that satisfies the converse of condition AC3, is the 2 Phase Commit protocol. Is not good to enforce the condition because to enforce that the fail tolerance must be 0. So, with every failure the transition is aborted.

### E.3.2 Exercise 14

Consider 2PC with the cooperative termination protocol. Describe a scenario (a particular execution) involving site failures only, which causes operational sites to become blocked.

We have an execution in which the coordinator fails after receiving the votes, but before sending the decision to the other processes. If the votes were all Yes the cooperative termination protocol can't unlock the system because no process knows the decision yet. So, the system is blocked.

### E.3.3 Exercise 15

Show that Paxos is not live.

In Paxos a round can be blocked by a successive round if the successive sends a prepare before acceptors voted in the previous round. The acceptors will respond with a promise to the successive round and will not vote in the previous round. This can happen infinite time in a row, so the protocol is not live.

### E.3.4 Exercise 16

Assume that acceptors do not change their vote. In other words, if they vote for value  $v$  in round  $i$ , they will not send learn messages with value different from  $v$  in larger rounds. Show

that Paxos, with this modification, is safe. Unfortunately, the modification introduces a severe liveness problem (the protocol can reach a livelock).

If a value  $v$  is accepted in round  $i$ , this means that a quorum of acceptors will always vote  $v$ , so is the only value that can be proposed and accepted in successive rounds. The problem is that if the acceptors voted three different values and no value has a quorum of votes, the system is in livelock.

### E.3.5 Exercise 17

How many messages are used in Paxos if no message is lost and in the best case? Is it possible to reduce the number of messages without losing tolerance to failures and without changing the number of proposers, acceptors, and learners?

In the best case the number of messages sent is  $(3 + l)a$ , with  $a$  acceptors and  $l$  learners. The messages can be reduced sending the votes only to the learner that associated to the round (if learners and proposers are the same).

### E.3.6 Exercise 18

Assume that you remove the property that every round is associate to a unique proposer. After collecting a quorum of  $n - f$  promises (where  $n$  is the number of acceptors and  $f$  is such that  $n = 2f + 1$ ), the proposer chooses one of the values voted in max round in the promises (of course it is not unique, the proposer chooses just one in an arbitrary way). Show that Paxos is not safe any more.

If the value is picked at random, a less voted value could be picked and this is unsafe because the other value could have been accepted in previous rounds.

### E.3.7 Exercise 19

Assume that all proposers are learners as well. Let even rounds be assigned to proposers with the rules that we know. Moreover, If round  $2i$  is assigned to proposer  $p$ , then also round  $2i + 1$  is assigned to proposer  $p$ . Odd rounds are “recovery” rounds. If round  $2i$  is a fast round and if the proposer of round  $2i$  sees a conflict (it is also a learner), then the proposer immediately sends an accept for round  $2i + 1$  with the value that has been most voted in round  $2i$ , without any prepare and any promise. Is safety violated? If yes, show an example. If not, demonstrate safety.

If two value are found in promises, choosing the most voted for the fast round is safe because only the most voted could be accepted in previous round, works the same as Fast-Paxos.

### E.3.8 Exercise 20

You are an optimization freak. You realize that in Fast Paxos, in some cases, it is not necessary that the proposer collects  $n - f$  (the Fast Paxos quorum) promises to take a decision. Which is the minimum quorum and under what hypothesis this minimum quorum is enough to take a decision?

To ensure safety only the value chosen by the proposer after receiving the promises could be accepted in previous round. So, the quorum could be set as 1 votes more than the maximum votes that the second most voted value could have received. If  $|v|$  is the number of votes for

the second most voted value and  $d$  the number of promises not received yet, the quorum can be set as:

$$Q = |v| + d + 1$$

### E.3.9 Exercise 21

Show that Raft is not live.

A term is skipped if a timeout is reached. This timeout can be reached infinite times in a row, so the protocol is not live.

### E.3.10 Exercise 22

In Raft, it is sometimes possible that the elected leader for term  $t$  has not all the log entries that are stored in the followers. Show that, in that case, the log entries missing at the leader are actually not committed and so they can be overwritten by the new leader.

Suppose the leader  $p$  doesn't have a log  $l$  hat is committed. To be committed  $l$  must be in at least the majority of logs, so this majority will not vote for  $p$  and  $p$  can't become leader, so is impossible that  $l$  is committed.

### E.3.11 Exercise 23

Show that the Ben-Or randomised consensus algorithm terminates with high probability (so that the probability that it does not terminate goes to zero as the number of rounds goes to infinity).

For every round, the probability to terminate the protocol is  $p < 1$ . The probability to not end before a round  $n$  is:

$$P(n) = (1 - p)^n$$

So, as  $n$  grows the probability becomes 0:

$$\lim_{n \rightarrow +\infty} (1 - p)^n = 0$$

### E.3.12 Exercise 24

Build a run of the Ben-Or randomised consensus algorithm that never terminates.

A run that never terminates is a run in which the majority doesn't choose the same value. An example is a system with 3 processes and one failed. So, to terminate the protocol each process must choose the same value. If we are unlucky the 2 processes could continue to choose different values.

### E.3.13 Exercise 25

Consider an asynchronous system of 5 processes that run the Ben-Or randomised consensus algorithm. The number of failures that the system allows is 2. Show that, if at most 2 failures occurs, then the probability that the protocol terminates after  $x$  rounds (or more) is smaller than  $\alpha^x$ , for some  $\alpha$ .

The lower probability to terminate the protocol is when there are 2 fails and the probability to terminate every round is:

$$p = \frac{1}{2^{n-f-1}} = \frac{1}{2^{5-2-1}} = \frac{1}{4}$$

So, the probability to terminate after round  $x$  with 2 fails is:

$$P(x) = (1 - p)^x = \left(\frac{3}{4}\right)^x$$

For any number of fails, the probability to terminate every round is  $p' \geq \frac{1}{4}$  and so, to terminate after round  $x$ :

$$P(x) = (1 - p')^x \leq \left(\frac{3}{4}\right)^x$$