SOUNDNESS IN NEGOTIATIONS

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ABSTRACT. Negotiations are a formalism for describing multiparty distributed cooperation. Alternatively, they can be seen as a model of concurrency with synchronized choice as communication primitive.

Well-designed negotiations must be sound, meaning that, whatever its current state, the negotiation can still be completed. In earlier work, Esparza and Desel have shown that deciding soundness of a negotiation is PSPACE-complete, and in PTIME if the negotiation is deterministic. They have also extended their polynomial soundness algorithm to an intermediate class of acyclic, non-deterministic negotiations. However, they did not analyze the runtime of the extended algorithm, and also left open the complexity of the soundness problem for the intermediate class.

In the first part of this paper we revisit the soundness problem for deterministic negotiations, and show that it is NLOGSPACE-complete, improving on the earlier algorithm, which requires linear space.

In the second part we answer the question left open by Esparza and Desel. We prove that the soundness problem can be solved in polynomial time for acyclic, weakly nondeterministic negotiations, a more general class than the one considered by them.

In the third and final part, we show that the techniques developed in the first two parts of the paper can be applied to analysis problems other than soundness, including the problem of detecting race conditions, and several classical static analysis problems. More specifically, we show that, while these problems are intractable for arbitrary acyclic deterministic negotiations, they become tractable in the sound case. So soundness is not only a desirable behavioral property in itself, but also helps to analyze other properties.

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1. Introduction

A multiparty atomic negotiation is an event in which several processes (agents) synchronize in order to select one out of a number of possible outcomes. In [3] Esparza and Desel introduced negotiation diagrams, or just negotiations, a model of concurrency with multiparty atomic negotiation as interaction primitive. A negotiation diagram describes a workflow of "atomic" negotiations. After an atomic negotiation concludes with the selection of an outcome, the workflow determines the set of atomic negotiations each agent is ready to participate in next.

Negotiation diagrams are closely related to workflow Petri nets, a very successful formalism for the description of business processes, and a back-end for graphical notations like BPMN (Business Process Modeling Notation), EPC (Event-driven Process Chain), or UML Activity Diagrams (see e.g. [14, 15]). In a nutshell, negotiation diagrams are workflow Petri nets that can be decomposed into communicating sequential Petri nets, a feature that makes them more amenable to theoretical study, while the translation into workflow nets (described in [1]) allows to transfer results and algorithms to business process applications.

The most prominent analysis problem for the negotiation model is *soundness*, a notion originally introduced for workflow Petri nets. Loosely speaking, a negotiation is sound if from every reachable configuration there is an execution leading to proper termination of the negotiation. In [3] it is shown that the soundness problem is PSPACE-complete for non-deterministic negotiations and CONP-complete for acyclic non-deterministic negotiations. For this reason, and in search of a tractable class, [3] introduces the class of *deterministic negotiations*. In deterministic negotiations all agents are deterministic, meaning that they are never ready to engage in more than one atomic negotiation per outcome (in the same way that in a deterministic automaton, for each action the automaton is only ready to move to one state). In [3] the soundness problem is investigated for acyclic negotiations. The main results are a polynomial time algorithm for checking soundness of deterministic negotiations, and an extension of the algorithm to the more expressive class of weakly deterministic negotiations. However, whether the extended algorithm is polynomial or not is left open. In [4] the polynomial result for acyclic deterministic negotiations is extended to the cyclic case.

In this paper we continue the line of research initiated in [3, 4], and present three contributions.

In the first contribution we revisit the soundness problem for deterministic negotiations. We identify *anti-patterns*, i.e., structures of the graph of a negotiation, and show that a deterministic negotiation is unsound iff it exhibits at least one of them. As an easy consequence of this theorem, we obtain a NLOGSPACE algorithm for checking soundness, contrary to the algorithm of [4], which requires linear space. Since soundness of deterministic negotiations is easily shown to be NLOGSPACE-hard, our result settles the complexity of the soundness problem.

In the second contribution we answer the question left open in [3]. We prove that the soundness problem can be solved in polynomial time for acyclic, weakly non-deterministic negotiations, a class even more general than the one considered in [3]². The result is based on a game-theoretic solution to the *omitting problem*, an analysis problem of independent

¹In [3] the notion of soundness has one more requirement, which makes the soundness problem for acyclic negotiations coNP-hard and in DP.

²The weakly deterministic negotiations of [3] are called *very weakly non-deterministic negotiations* in this paper. As the name indicates, every negotiation in this class is also weakly non-deterministic.

interest. Further, we show that if we leave out one of the two assumptions, acyclicity or weak non-determinism, then the problem becomes CoNP-complete³. These results set a limit to the class of negotiations with a polynomial soundness problems, but also admit a positive interpretation. Indeed, the soundness problem for arbitrary negotiations is PSPACE-complete [3], and so of higher complexity (under the usual assmption PTIME \subset NP \cup CONP \subset PSPACE).

In the third and final contribution, we show that the techniques developed in the first two parts of the paper, namely anti-patterns and our game-theoretic solution to the omitting problem, can be applied to analysis problems other than soundness. More specifically, we show that, while these problems are intractable for arbitrary deterministic negotiations, they become tractable in the sound case. So soundness is not only a desirable behavioral property in itself, but also helps to analyze other properties. The first problem we consider is the existence of races, i.e., executions in which two given atomic negotiations are concurrently enabled. We show that for acyclic deterministic negotiations the problem is in NLOGSPACE. Then we analyze several classical program analysis problems for negotiations that manipulate data, for example whether every value written into a variable is guaranteed to be read, or whether a variable can be allocated and deallocated by two atomic negotiations taking place in parallel. Such problems have been studied for workflow nets in [13, 11], and exponential algorithms have been proposed. We show that for acyclic sound deterministic negotiations the problems can be solved in polynomial time.

Related formalisms and related work. The connection between negotiations and Petri nets is studied in detail in [1]. Every negotiation can be transformed into an exponentially larger 1-safe workflow Petri net with an isomorphic reachability graph. Every deterministic negotiation is equivalent to a 1-safe workflow free-choice net with a linear blow-up. Conversely, every sound workflow free-choice net can be transformed into a sound deterministic negotiation with a linear blow-up. Recent papers on free-choice workflow Petri nets are [7, 5]. In [7] soundness is also characterized in terms of anti-patterns, which can be used to explain why a given workflow net is unsound. Our work provides an anti-pattern characterization for acyclic weakly non-deterministic negotiations, which goes beyond the free-choice case. In [5] a polynomial reduction algorithm for free-choice workflow Petri nets is presented. Our results show that soundness is also polynomial for workflow Petri nets coming from acyclic weakly deterministic negotiations.

As a process-based concurrent model, negotiations can be compared with another well-studied model for distributed computation, namely Zielonka automata [17, 2, 10]. Such an automaton is a parallel composition of finite transition systems with synchronization on common actions. The important point is that a synchronization involves exchange of information between states of agents: the result of the synchronization depends on the states of all the components taking part in it. Zielonka automata have the same expressive power as arbitrary, possibly nondeterministic negotiations. Deterministic negotiations correspond to a subclass that does not seem to have been studied yet, and for which verification becomes considerably easier. For example, the question whether some local state occurs in some execution is Pspace-complete for "sound" Zielonka automata, while it can be answered in polynomial time for sound deterministic negotiations.

³We show that CoNP-hardness holds even for a very mild relaxation of acyclicity.

A somewhat similar graphical formalism are message sequence charts/graphs, used to describe asynchronous communication. Questions like non-emptiness of intersection are in general undecidable for this model, even assuming that communication buffers are bounded. Subclasses of message sequence graphs with decidable model-checking problem were proposed, but the complexity is PSPACE-complete [8].

Overview. Section 2 introduces definitions and notations. Section 3 revisits the soundness problem for deterministic negotiations. Section 4 shows that soundness of acyclic weakly non-deterministic negotiations can be decided in polynomial time; the first part of the section solves the omitting problem, and the second part applies the solution to the soundness problem. Section 5 proves that dropping acyclicity or weak non-determinism makes the soundness problem intractable. Section 6 gives polynomial algorithms for the race problem and the static analysis problems of sound deterministic negotiations. Section 7 presents our conclusions.

2. Negotiations

A negotiation \mathbb{N} is a tuple $\langle Proc, N, dom, R, \delta \rangle$, where Proc is a finite set of processes (or agents) that can participate in negotiations, and N is a finite set of nodes (or atomic negotiations) where the processes can synchronize; The function $dom: N \to \mathcal{P}(Proc)$ associates to every atomic negotiation $n \in N$ the (non-empty) set dom(n) of processes participating in it. Negotiations come equipped with two distinguished initial and final atomic negotiations n_{init} and n_{fin} in which all processes in Proc participate. Nodes are denoted as m or n, and processes as p or q; possibly with indices.

The set of possible outcomes of atomic negotiations is denoted R, and we use a, b, ... to range over its elements. We assume that every atomic negotiation (except possibly for n_{fin}) has at least one outcome. The control flow in a negotiation is determined by a partial transition function $\delta: N \times R \times Proc \xrightarrow{\cdot} \mathcal{P}(N)$, telling that after the outcome a of an atomic negotiation n, process $p \in dom(n)$ is ready to participate in any of the negotiations from the set $\delta(n, a, p)$. So for every $n' \in \delta(n, a, p)$ we have $p \in dom(n') \cap dom(n)$. Every atomic negotiation $n \in N$ has its set of possible outcomes out(n), and for every $n, a \in out(n)$ and $p \in dom(n)$ the result $\delta(n, a, p)$ has to be defined. So all processes involved in an atomic negotiation should be ready for all its possible outcomes. Observe that atomic negotiations may have one single participant process, and/or have one single outcome.

Negotiations admit a graphical representation. A node n is represented as a black bar with a white circle, or port, for each process in dom(n). An entry $\delta(n,a,p) = \{n_1,\ldots,n_k\}$ is represented by a hyper-arc, labeled by a, that connects the port of process p in n with the ports of process p in n_1,\ldots,n_k . Figure 1 shows the graphical representations of three negotiations. The one on the left has $Proc = \{p_0, p_1\}$, $N = \{n_0, \ldots, n_5\}$, $R = \{a, b\}$, and, for example $dom(n_1) = \{p_0\}$, $dom(n_4) = \{p_0, p_1\}$, $\delta(n_4, b, p_0) = \{n_1\}$ and $\delta(n_4, b, p_1) = \{n_2\}$. This negotiation does not contain any proper hyper-arcs, but the second one does; there we have $\delta(n_0, a, p_1) = \{n_2, n_3\}$. More details can be found in [3].

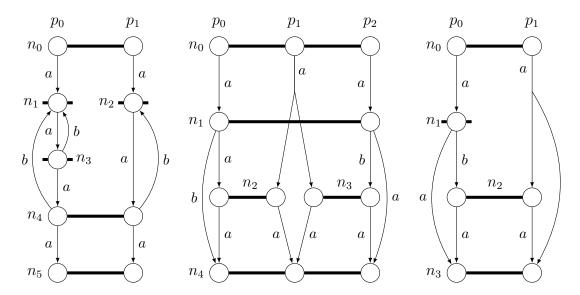


Figure 1: Three negotiations

Configurations. A configuration of a negotiation is a function $C \colon Proc \to \mathcal{P}(N)$ mapping each process p to the (non-empty) set of atomic negotiations in which p is ready to engage. The initial and final configurations C_{init} , C_{fin} are given by $C_{init}(p) = \{n_{init}\}$ and $C_{fin}(p) = \{n_{fin}\}$ for all $p \in Proc$. An atomic negotiation n is enabled in a configuration C if $n \in C(p)$ for every $p \in dom(n)$, that is, if all processes that participate in n are ready to proceed with it. A configuration is a deadlock if no atomic negotiation is enabled in it. If an atomic negotiation n is enabled in C, and n is an outcome of n, then we say that n can be executed, and its execution produces a new configuration n given by n for n for n for n and n derivative n for n for n and n derivative n for n

Runs. A run of a negotiation \mathbb{N} from a configuration C_1 is a finite or infinite sequence $w = (n_1, a_1)(n_2, a_2) \dots$ such that there are configurations C_2, C_3, \dots with

$$C_1 \xrightarrow{(n_1,a_1)} C_2 \xrightarrow{(n_2,a_2)} C_3 \cdots$$

We denote this by $C_1 \xrightarrow{w}$, or $C_1 \xrightarrow{w} C_k$ if the sequence is finite and finishes with C_k . In the latter case we say that C_k is reachable from C_1 on w. We simply call it reachable if w is irrelevant, and write $C_1 \xrightarrow{*} C_k$. Consider for example the third negotiation of Figure 1. If we represent a configuration C by the tuple $(C(p_0), C(p_1))$ then

$$(\{n_0\}, \{n_0\}) \xrightarrow{(n_0, a)} (\{n_1\}, \{n_2, n_3\}) \xrightarrow{(n_1, b)} (\{n_2\}, \{n_2, n_3\}) \xrightarrow{(n_2, a)} (\{n_3\}, \{n_3\})$$

is a run.

A run is called *initial* if it starts in C_{init} . An initial run is *successful* if it starts in C_{init} and ends in C_{fin} . In the three negotiations of Figure 1 we have $n_{init} = n_0$, and $n_{fin} = n_5, n_4, n_3$, respectively. The run shown above is both initial and a successful run.

Acyclicity. The graph of a negotiation has N, the set of atomic negotiations, as set of vertices; the edges are $n \xrightarrow{p,a} n'$ if $n' \in \delta(n,a,p)$. Observe that $p \in dom(n) \cap dom(n')$.

A negotiation is *acyclic* if its graph is so. For an acyclic negotiation \mathbb{N} we fix a linear order $\preceq_{\mathbb{N}}$ on its nodes that is a topological order on the graph of \mathbb{N} . This means that if there is an edge from m to n in the graph of \mathbb{N} then $m \preceq_{\mathbb{N}} n$. The last two negotiations of Figure 1 are acyclic, while the first one is not. For the negotiation in the middle of the figure there are two options for the topological order, corresponding to fixing $n_2 \preceq_{\mathbb{N}} n_3$ or $n_3 \preceq_{\mathbb{N}} n_2$.

Soundness. A negotiation \mathbb{N} is *sound* if every initial run can be completed to a successful run. If a negotiation has no infinite runs (for example, this is the case if the negotiation is acyclic), then it is sound iff it has no reachable deadlock configuration. The three negotiations of Figure 1 are sound. If in the negotiation on the right we change $\delta(n_0, a, p_1)$ from $\{n_2, n_3\}$ to $\{n_3\}$, then the negotiation is no longer sound. Indeed, after the change the negotiation has the run

$$(\{n_0\}, \{n_0\}) \xrightarrow{(n_0, a)} (\{n_1\}, \{n_3\}) \xrightarrow{(n_1, b)} (\{n_2\}, \{n_3\})$$

which leads to a deadlock.

Remark 2.1. Our definition of soundness is slightly different from the one used in [3]. The definition of [3], which follows the definition of soundness for workflow Petri nets introduced in [14], requires an additional property: for every atomic negotiation n there is an initial run that enables n. We use the weaker definition because it leads to cleaner theoretical results, and because, as we shall see, the two definitions are essentially equivalent for deterministic negotiations (see Remark 3.13).

Determinism. Process p is deterministic in a negotiation \mathbb{N} if for every $n \in N$ and $a \in out(n)$, the set of possible next negotiations, $\delta(n,a,p)$, is a singleton. A negotiation is deterministic if every process $p \in Proc$ is deterministic. Graphically, a negotiation is deterministic if it does not have any proper hyper-arc. The negotiation on the left of Figure 1 is deterministic.

A negotiation is weakly non-deterministic if for every $n \in N$ at least one of the processes in dom(n) is deterministic. A negotiation is very weakly non-deterministic⁴ if for every $n \in N$, $a \in out(n)$, and $p \in Proc$, there is a deterministic process q such that $q \in dom(n')$ for all $n' \in \delta(n, a, p)$.

The negotiation in the middle of Figure 1 is weakly non-deterministic. Indeed, the processes p_0 and p_2 are deterministic, and every node has p_0 or p_2 (or both) in its domain. However, it is not very weakly non-deterministic. To see this, observe that $\delta(n_0, a, p_1) = \{n_1, n_2\}$, but the intersection $dom(n_1) \cap dom(n_2) = \{p_1\}$ does not contain any deterministic process. On the contrary, the negotiation on the right of the figure is very weakly non-deterministic, because the deterministic process p_0 belongs to the domain of all nodes.

Weakly non-deterministic negotiations allow to model deterministic negotiations with global resources (see Section 6). The resource (say, a piece of data) can be modeled as an additional process, which participates in the atomic negotiations that use the resource. For example, the negotiation in the middle of Figure 1 models a situation in which processes p_0

⁴This class was called *weakly deterministic* in [3].

and p_2 negotiate in n_1 which of the two will have access to the resource modeled by p_1 . If the outcome of n_1 is a, then p_0 has access to the resource at node n_2 , and if it is b, then p_2 has access to it at node n_3 .

3. Soundness of deterministic negotiations

We revisit the soundness problem for deterministic negotiations. We give the first NLOGSPACE algorithm for the problem, in contrast with the polynomial algorithm of [3], which requires linear space. The algorithm is derived from a theorem that provides a novel characterization of soundness in terms of *anti-patterns*. The theorem allows not only to check soundness, but also to diagnose why a given negotiation is unsound.

Fix a negotiation \mathbb{N} . A local path of \mathbb{N} is a path $n_0 \xrightarrow{p_0, a_0} n_1 \xrightarrow{p_1, a_1} \dots \xrightarrow{p_{k-1}, a_{k-1}} n_k$ in the graph of \mathbb{N} . A local path is

- a circuit if $n_0 = n_k$ and $k \ge 1$;
- a *p-path* if $p_0 = \cdots = p_{k-1} = p$;
- realizable from a configuration C if there is a run

$$C \xrightarrow{(n_0, a_0)} C_0' \xrightarrow{w_0} C_1 \xrightarrow{(n_1, a_1)} C_1' \cdots C_{k-1} \xrightarrow{(n_{k-1}, a_{k-1})} C_{k-1}' \xrightarrow{w_k} C_k$$

such that $p_i \notin dom(w_{i+1})$ for all i = 0, ..., k-1. (Here dom(v) denotes the set of processes involved in v, that is, $dom(v) = \bigcup \{dom(n) : (n, a) \text{ appears in } v \text{ for some } a \in out(n)\}$.) We say that the run realizes the path from C.

Example 3.1. For example, $n_0 \xrightarrow{p_1,a} n_2 \xrightarrow{p_1,a} n_4 \xrightarrow{p_0,a} n_5$ is a local path of the first negotiation of Figure 1. The path is realized from the initial configuration by the run

$$(\{n_0\}, \{n_0\}) \xrightarrow{(n_0, a)} (\{n_1\}, \{n_2\}) \xrightarrow{(n_1, a)} (\{n_3\}, \{n_2\})$$
$$\xrightarrow{(n_3, a)} (\{n_4\}, \{n_2\}) \xrightarrow{(n_2, a)} (\{n_4\}, \{n_4\}) \xrightarrow{(n_4, a)} (\{n_5\}, \{n_5\})$$

Indeed, we can take $w = (n_0, a) w_1 (n_2, a) w_2 (n_4, a)$, with $w_1 = (n_1, a) (n_3, a)$ and $w_2 = \epsilon$.

The following lemma shows that every local path of a sound and deterministic negotiation is realizable from some reachable configuration.

Lemma 3.2. Let π be a local path of a sound deterministic negotiation \mathbb{N} , and let n_0 be the first node of π . Then π is realizable from every reachable configuration that enables n_0 .

Proof. Let $\pi = n_0 \xrightarrow{p_0, a_0} n_1 \xrightarrow{p_1, a_1} \cdots \xrightarrow{p_{k-1}, a_{k-1}} n_k$, and let C be a reachable configuration such that $C(p) = n_0$ for every $p \in dom(n_0)$. By induction on i we show that there is a run $C \xrightarrow{*} C_i$ realizing $n_0 \xrightarrow{p_0, a_0} n_1 \xrightarrow{p_1, a_1} \cdots \xrightarrow{p_{i-1}, a_{i-1}} n_i$ and such that n_i is enabled in C_i .

For i = 0, we simply take $C_i = C$. For the induction step we assume the existence of C_i in which n_i is enabled. Let C'_{i+1} be the result of executing (n_i, a_i) from C_i . Observe that $C'_{i+1}(p_i) = n_{i+1}$ (recall that \mathcal{N} is deterministic). Since \mathcal{N} is sound, and C'_{i+1} is reachable, there is a run from C'_{i+1} to C_{fin} . We set then C_{i+1} to be the first configuration on this run when n_{i+1} is enabled.

In particular, Lemma 3.2 states that there is an initial run containing the atomic negotiation m iff there is a local path from n_{init} to m. If $dom(m) \cap dom(n) \neq \emptyset$ then the lemma also provides an easy test for deciding the existence of a run containing both m, n: it suffices to check the existence of a local path $n_{init} \stackrel{*}{\to} m \stackrel{*}{\to} n$, or with m, n interchanged.

Our algorithm for checking soundness of deterministic negotiations checks for certain patterns in the graph of the negotiation. Since negotiations exhibiting the patterns are unsound, we call them *anti-patterns*. In order to define them we need to introduce *forks* and recall the notion of *dominating node* of a local path introduced in [4].

Definition 3.3. Let $\mathcal{N} = (Proc, N, dom, R, \delta)$ be a deterministic negotiation. A tuple $(p_1, p_2, n_1, n_2) \in Proc^2 \times N^2$ is a *fork* of \mathcal{N} if there exists a local path from n_{init} to a node $n \in N$ and an outcome $a \in out(n)$ such that

- $p_i \in dom(n) \cap dom(n_i)$ for i = 1, 2;
- for i = 1, 2 there exists a p_i -path π_i leading from $\delta(n, a, p_i)$ to n_i ; and
- π_1 and π_2 are disjoint, i.e., no node appears in both.

Definition 3.4. A node n of a local path π dominates π if $dom(m) \subseteq dom(n)$ for every node m of π .

Example 3.5. The tuple (p_0, p_1, n_3, n_4) is a fork of the negotiation on the left of Figure 1. We can choose $n = n_0$. The p_0 -path is $n_0 \xrightarrow{p_0,a} n_1 \xrightarrow{p_0,a} n_3$, and the p_1 -path is $n_0 \xrightarrow{p_1,a} n_4$. The tuple (p_0, p_1, n_4, n_5) is not a fork.

Consider now the local circuit $n_2 \xrightarrow{p_1,a} n_4 \xrightarrow{p_1,b} n_2$. The node n_4 is dominating, since its domain includes all processes.

Lemma 3.6 ([4], Lemma 2). Every reachable local circuit of a sound deterministic negotiation (that is, every local circuit containing a node reachable from n_{init} by a local path) has a dominating node. ⁵

Example 3.7. Lemma 3.6 does not hold for arbitrary sound negotiations. Consider the non-deterministic negotiation of Figure 2. It is easy to see that the negotiation is sound. However, the local circuit $n_1 \xrightarrow{p_1,a} n_2 \xrightarrow{p_1,a} n_1$ has no dominating node, because $dom(n_1) = \{p_0, p_1\}$ and $dom(n_2) = \{p_1, p_2\}$.

Definition 3.8. (1) An anti-pattern of type \mathcal{B} is a p-path leading from n_{init} to a node n such that no p-path leads from n to n_{fin} .

- (2) An anti-pattern of type \mathcal{F} is a fork (p_1, p_2, n_1, n_2) such that $p_2 \in dom(n_1)$ and $p_1 \in dom(n_2)$.
- (3) An anti-pattern of type C is a local circuit without a dominating node.

Example 3.9. The last two anti-patterns are illustrated in Figure 3. The tuple (p_0, p_1, n_1, n_2) is a fork of the negotiation on the left satisfying $p_1 \in dom(n_1)$ and $p_0 \in dom(n_2)$. The local circuit $n_1 \xrightarrow{p_0, a} n_2 \xrightarrow{p_1, a} n_3 \xrightarrow{p_2, a} n_1$ of the negotiation on the right has no dominating node. Observe that the negotiation has no anti-pattern of type \mathcal{F} .

Lemma 3.10. A deterministic negotiation containing an anti-pattern is unsound.

Proof. Assume that a deterministic negotiation \mathbb{N} contains an anti-pattern. If the anti-pattern is of type \mathcal{B} , then since the p-path leading to n is realizable (Lemma 3.2), some

⁵In [4] dominating nodes of circuits are called synchronizers.

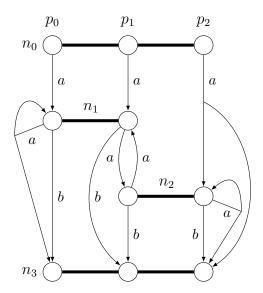


Figure 2: A local circuit without a dominating node

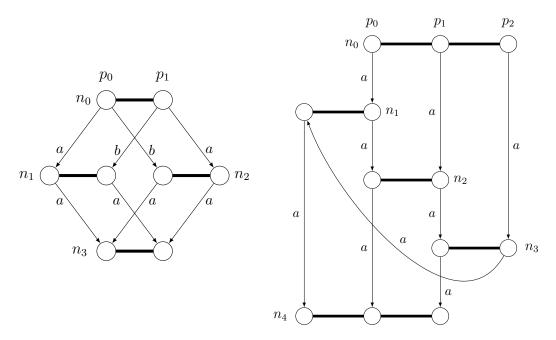


Figure 3: Anti-patterns

reachable configuration C satisfies $C(p) = \{n\}$. Since no p-path leads from n to n_{fin} , we have $C'(p) \neq \{n_{fin}\}$ for every configuration C' reachable from C, and so \mathcal{N} is not sound.

If the anti-pattern is of type C, then the result follows from Lemma 3.6.

If the anti-pattern is of type \mathcal{F} , then \mathcal{N} contains a fork (p_1, p_2, n_1, n_2) such that $p_2 \in dom(n_1)$, and $p_1 \in dom(n_2)$. Let n and a be the node and the outcome required by the definition of a fork. Since local paths of sound deterministic negotiations are realizable

(Lemma 3.2), some reachable configuration C enables n. Let $C \xrightarrow{(n,a)} C'$. By the definition of the anti-pattern, there are disjoint p_1 - and p_2 -paths π_1 and π_2 leading to n_1 and n_2 , respectively. Using again soundness and the fact that π_1 and π_2 are disjoint, we can show by induction on $|\pi_1| + |\pi_2|$ that there is a run $C' \xrightarrow{w} C''$ that realizes both π_1 and π_2 from C'. So we have $C''(p_1) = n_1$ and $C''(p_2) = n_2$, because neither n_1 can be executed before n_2 gets enabled, nor the other way round. Since $p_2 \in dom(n_1)$ and $p_1 \in dom(n_2)$, neither n_1 nor n_2 can ever occur from C''. By the same argument as above, this implies that the initial run leading to C'' cannot be extended to a successful run, and so \mathcal{N} is unsound. \square

Lemma 3.11. An unsound deterministic negotiation contains an anti-pattern.

Proof. Let \mathbb{N} be an unsound deterministic negotiation without anti-patterns of type \mathcal{B} . We prove that \mathbb{N} has an anti-pattern of type \mathcal{F} or \mathbb{C} .

Let $Proc = \{p_1, \ldots, p_n\}$ be the set of processes of \mathbb{N} . We first claim that \mathbb{N} has a deadlock. For every atomic negotiation n and $p \in dom(n)$, let d(n,p) be the length of the shortest p-path leading from n to n_{fin} (or ∞ if the path does not exist), and for every reachable configuration C, define $d(C) = (d(C(p_1), p_1), \ldots, d(C(p_n), p_n))$. Since \mathbb{N} is unsound, some initial run $C_{init} \xrightarrow{w} C$ cannot be extended to a successful run. Choose the run so that d(C) is minimal w.r.t. the lexicographic order. We claim that C is a deadlock. Assume the contrary. Then C enables an atomic negotiation n. Let p_i be the process of dom(n) with the smallest index. Since \mathbb{N} is deterministic, we have $\delta(n, a, p_i) = \{n_a\}$ for some atomic negotiation n_a . Since \mathbb{N} has no anti-patterns of type \mathbb{B} , there exists a p_i -path from n to n_{fin} , and so, by the definition of the distance function d, there exists an outcome $a \in out(n)$ such that $d(n, p_i) > d(n_a, p_i)$. Taking $C \xrightarrow{(n,a)} C'$, we get that d(C') is lexicographically smaller than C, contradicting the minimality of C, and so the claim is proved.

Let C be a deadlock configuration of \mathbb{N} . This means that there are processes p_0, \ldots, p_{k-1} , and nodes n_0, \ldots, n_{k-1} ($k \geq 2$) such that $C(p_i) = n_i$ and $p_i \in dom(n_{i+1})$ for every $0 \leq i < k$. (Intuitively, at the configuration C process p_0 waits for p_{k-1} , p_1 waits for p_0 , etc.) A run from C_{init} to C yields a fork $(p_i, p_{i+1}, n_i, n_{i+1})$ for every pair of nodes n_i, n_{i+1} : Indeed, we can choose the pair (n, a) required in the definition of a fork as the last pair in the run satisfying $\{p_i, p_{i+1}\} \subseteq dom(n)$, and choose the path π (resp. π') as a p_i -path from $\delta(n, a, p_i)$ to n_i (resp. a p_{i+1} -path from $\delta(n, a, p_{i+1})$ to n_{i+1}).

If k = 2 then (p_0, p_1, n_0, n_1) is a pattern of type \mathcal{F} . So assume that k > 2 and that \mathcal{N} has no pattern of type \mathcal{F} . We prove that \mathcal{N} has a local circuit without a dominating node.

We claim that for every $0 \leq i < k$ there is some p_i -path π_i from n_i to n_{i+1} (setting $n_k = n_0$). Assume the contrary, and let (n, a) be the pair in the definition of the fork $(p_i, p_{i+1}, n_i, n_{i+1})$. Then we can extend the p_i -path π from $\delta(n, a, p_i)$ to n_i with some p_i -path π' from n_i to some node n' with $p_{i+1} \in dom(n')$, and such that p_{i+1} does not occur in π' except for n'. Such a node n' exists because \mathbb{N} has no anti-pattern of type \mathbb{B} , and the final node n_{fin} contains all processes. The p_i -path $\pi\pi'$ is still disjoint with the p_{i+1} -path leading from $\delta(n, a, p_{i+1})$ to n_{i+1} . Therefore, the fork $(p_i, p_{i+1}, n', n_{i+1})$ is an anti-pattern of type \mathbb{F} , contradicting the hypothesis, and the claim is proven.

Note that $p_{i+1} \notin dom(n_i)$, since otherwise $(p_i, p_{i+1}, n_i, n_{i+1})$ would be a pattern of type \mathcal{F} . Also, p_{i+1} cannot occur in π_i except for n_{i+1} : otherwise we could take the shortest prefix π of π_i ending in a node $n'_i \neq n_{i+1}$ with $\{p_i, p_{i+1}\} \subseteq dom(n'_i)$ and find another \mathcal{F} pattern, using the paths $\pi_i \pi$ and π_{i+1} . Let γ be the concatenation of $\pi_0, \pi_1, \dots, \pi_{k-1}$. Then γ is a

local circuit of \mathbb{N} , and every p_i belongs to the domain of some node of γ . Since p_{i+1} does not belong to the domain of any node of π_i , except n_{i+1} , no node of γ has all of p_0, \ldots, p_{k-1} in its domain. So γ has no dominating node, and therefore \mathbb{N} contains a pattern of type \mathbb{C} .

Now we can easily prove:

Theorem 3.12. Soundness of deterministic negotiations is NLOGSPACE-complete.

Proof. The existence of an anti-pattern of type \mathcal{B} or \mathcal{F} is clearly verifiable in NLOGSPACE. To non-deterministically check the existence of a path of type \mathcal{C} using logarithmic space, we guess an integer M and a process p. Then, we guess on-the fly a circuit π , such that (1) each node of π has domain of size at most M, (2) p occurs in π , and (3) one node of π has domain of size M and does not contain p. Clearly, this implies that π contains at least M+1 processes, but no dominant node.

For the lower bound we reduce a reachability question $s \stackrel{*}{\to} t$ in a directed graph G (where s has no incoming and t no outgoing edges) to the soundness of a deterministic negotiation $\mathcal{N}(G)$ with one process p. The atomic negotiations of $\mathcal{N}(G)$ are the vertices of G, with domain $\{p\}$. The initial node is s and the final node is t. If $u \to v$ is an edge of G, then the atomic negotiation u has an outcome (u,v) with $\delta(u,u,v),p)=v$. Moreover, for every vertex u of G except t the atomic negotiation u has an outcome back with $\delta(u,back,p)=s$. Clearly, $\mathcal{N}(G)$ is sound iff there is some path $s\stackrel{*}{\to} t$ in G.

Remark 3.13. As announced in Remark 2.1, for deterministic negotiations the notion of soundness used in this paper and the one of [3] essentially coincide. More precisely, we show that the two notions coincide under the very weak assumption that for every atomic negotiation n there is a local path from n_{init} to n. (Atomic negotiations that do not satisfy this condition can be identified and removed in NLOGSPACE, and their removal does not change the behavior of the negotiation.)

Recall that the definition of [3] requires that (a) every run can be extended to a successful run and (b) that for every atomic negotiation n some initial run enables n. Now, let \mathbb{N} be a deterministic negotiation that is sound according to the definition of this paper, i.e., satisfies (a), and such that for every every atomic negotiation n there is a local path from n_{init} to n. We show that (b) also holds. For $n = n_{fin}$, (b) holds because every run can be extended to a successful run. For $n \neq n_{fin}$, we observe that, by Lemma 3.2, the local path leading from n_{init} to n is realizable. By the definition of realizability, there is a reachable configuration C and a process p such that $C(p) = \{n\}$. Since \mathbb{N} is sound, it has a run leading from C to C_{fin} . Since \mathbb{N} is deterministic and $n \neq n_{fin}$, this run necessarily executes n, and we are done.

4. Beyond determinism: Tractable cases

In this section we investigate how much nondeterminism we can allow, while retaining polynomiality of the soundness problem, a question that was left open in [3, 4]. We prove that soundness of acyclic, weakly non-deterministic negotiations can be decided in polynomial time. In Section 5 we show that the problem becomes intractable for both arbitrary weakly non-deterministic negotiations, and for acyclic non-deterministic negotiations.

The polynomial time algorithm is based on a game-theoretic solution to the *omitting* problem, which is a problem of independent interest. Section 4.1 introduces it, and shows

that it is fixed-parameter tractable for sound, acyclic and deterministic negotiations. Section 4.2 uses this result to prove that the soundness problem for weakly non-deterministic negotiations is in PTIME.

4.1. **Omitting problem.** Let $B \subseteq N$ be a set of nodes of a negotiation \mathbb{N} . We say that a run $(n_1, a_1)(n_2, a_2) \cdots$ of \mathbb{N} omits B if $n_i \notin B$ for all i. Let $P \subseteq N \times R$ be a set of pairs consisting of a node and an outcome. We say that a run of \mathbb{N} includes P and omits B if it omits B and contains all the pairs from P.

Definition 4.1. The *omitting problem* consists of deciding, given \mathbb{N} , P, and B, whether there is a successful run of \mathbb{N} including P and omitting B.

Given a constant K, the K-omitting problem is the subproblem of the omitting problem in which P has size at most K.

We show that the omitting problem for sound, acyclic, and deterministic negotiations can be reduced to solving a safety game (see e.g. [12] for an introduction to games). As a first step we define a two-player game G(N, B) with Players Adam and Eve, where the goal of Eve is to produce a successful run that omits B:

- the game positions of Eve are $N \setminus B$,
- the game positions of Adam are $N \times R$,
- from n, Eve can go to any (n, a) with $a \in out(n)$,
- from (n,a), Adam can choose any process $p \in dom(n)$ and go to $n' = \delta(n,a,p)$,
- the initial position is n_{init} ,
- Adam wins if the play reaches a node in B, Eve wins if the play reaches n_{fin} .

Observe that since \mathbb{N} is acyclic, the winning condition for Eve is actually a safety condition: every maximal play avoiding B is winning for Eve. So if Eve can win, then she wins with a positional strategy. A deterministic positional strategy for Eve is a function $\sigma: N \to R$, it indicates that at position n Eve should go to position $(n, \sigma(n))$. Since $G(\mathbb{N}, B)$ is a safety game for Eve, there is a biggest non-deterministic winning strategy for Eve, i.e., a strategy of type $\sigma_{max}: N \to \mathcal{P}(R)$. The strategy σ_{max} is obtained by computing the set W_E of all winning positions for Eve in $G(\mathbb{N}, B)$, and then setting for every $n \in \mathbb{N}$:

$$\sigma_{max}(n) = \{ a \in out(n) : \text{ for all } p \in dom(n), \, \delta(n, a, p) \in W_E \}.$$

Lemma 4.2. If \mathbb{N} has a successful run omitting B then Eve has a winning strategy in $G(\mathbb{N}, B)$.

Proof. Define $\sigma(n) = a$ if (n, a) appears in the run. For other nodes define the strategy arbitrarily. To check that this strategy is winning, it is enough to verify that every play respecting the strategy stays in the nodes appearing in the run.

Lemma 4.3. Let \mathbb{N} be a sound, acyclic, and deterministic negotiation such that Eve has a winning strategy $\sigma: \mathbb{N} \to \mathbb{R}$ in $G(\mathbb{N}, B)$. Consider the set S of nodes that are reachable on a play from n_{init} respecting σ . Then some successful run of \mathbb{N} contains precisely the nodes of S.

Proof. Recall that for an acyclic \mathbb{N} we fixed a topological order $\preccurlyeq_{\mathbb{N}}$. Let $n_1, n_2, ..., n_k$ be an enumeration of the nodes in $S \subseteq (N \setminus B)$ according to $\preccurlyeq_{\mathbb{N}}$. Let $w_i = (n_1, \sigma(n_1)) \cdots (n_i, \sigma(n_i))$. By induction on $i \in \{1, ..., k\}$ we prove that there is a configuration C_i such that $C_{init} \xrightarrow{w_i}$

 C_i is a run of \mathbb{N} . This will show that w_k is a successful run containing precisely the nodes of S.

For i = 1, $n_1 = n_{init}$, in C_{init} all processes are ready to do n_1 , so C_1 is the result of performing $(n_1, \sigma(n_1))$.

For the inductive step, we assume that we have a run $C_{init} \xrightarrow{w_i} C_i$, and we want to extend it by $C_i \xrightarrow{(n_{i+1},\sigma(n_{i+1}))} C_{i+1}$. Consider a play respecting σ and reaching n_{i+1} . The last step in this play is $(n_j,\sigma(n_j)) \to n_{i+1}$, for some $j \leq i$ and n_j in S. Since $\mathbb N$ is deterministic, we have $\delta(n_j,\sigma(n_j),p) = n_{i+1}$ for some process p. Since $j \leq i$ and $(n_j,\sigma(n_j))$ occurred in w_i (but not n_{i+1}), we have $C_i(p) = n_{i+1}$. If we show that $C_i(q) = \{n_{i+1}\}$ for all $q \in dom(n_{i+1})$ then we obtain that n_{i+1} is enabled in C_i and we get the required C_{i+1} . Suppose by contradiction that $C_i(q) = \{n_l\}$ for some $l \neq i+1$. We must have l > i+1, since otherwise n_l already occurred in w_i . By definition of our indexing $n_{i+1} \prec_{\mathbb N} n_l$. But then no run from C_i can bring process q to a state where it is ready to participate in negotiation n_{i+1} , and so, since $\mathbb N$ is deterministic, we have $C(p) = \{n_{i+1}\}$ for every configuration C reachable from C_i . This contradicts the fact that $\mathbb N$ is sound.

Lemmas 4.2 and 4.3 show that Eve wins in G(N, B) iff N has a successful run omitting B. We apply this result to the omitting problem.

Theorem 4.4. For every constant K, the K-omitting problem for deterministic, acyclic, and sound negotiations is in PTIME.

Proof. If for some atomic negotiation m we have $(m, a) \in P$ and $(m, b) \in P$ for $a \neq b$ then the answer is negative as \mathbb{N} is acyclic. So let us suppose that it is not the case. By Lemmas 4.2 and 4.3 our problem is equivalent to determining the existence of a deterministic strategy σ for Eve in the game $G(\mathbb{N}, B)$ such that $\sigma(m) = a$ for all $(m, a) \in P$, and all these (m, a) are reachable on a play respecting σ .

To decide this we calculate σ_{max} , the biggest non-deterministic winning strategy for Eve in $G(\mathbb{N}, B)$. This can be done in PTIME as the size of $G(\mathbb{N}, B)$ is proportional to the size of the negotiation. Strategy σ_{max} defines a graph $G(\sigma_{max})$ whose nodes are atomic negotiations, and edges are (m, a, m') if $(m, a) \in \sigma_{max}$ and $m' = \delta(m, a, p)$ for some process p. The size of this graph is proportional to the size of the negotiation. In this graph we look for a subgraph H such that:

- for every node m in H there is at most one a such that (m, a, m') is an edge of H for some m';
- for every $(m, a) \in P$ there is an edge (m, a, m') in H for some m', and moreover m is reachable from n_{init} in H.

We show that such a graph H exists iff there is a strategy σ with the required properties.

Suppose there is a deterministic winning strategy σ such that $\sigma(m) = a$ for all $(m, a) \in P$, and all these (m, a) are reachable on a play respecting σ . We now define H by putting an edge (m, a, m') in H if $\sigma(m) = a$ and $m' = \delta(m, a, p)$ for some process p. As σ is deterministic and winning, this definition guarantees that H satisfies the first item above. The second item is guaranteed by the reachability property that σ satisfies.

For the other direction, given such a graph H we define a deterministic strategy σ_H . We put $\sigma_H(m) = a$ if (m, a, m') is an edge of H. If m is not a node in H, or has no outgoing edges in H then we put $\sigma_H(m) = b$ for some arbitrary $b \in \sigma_{max}(m)$. It should be clear that σ_H is winning since every play respecting σ_H stays in winning nodes for Eve. By definition $\sigma_H(m) = a$ for all $(m, a) \in P$, and all these (m, a) are reachable on a play respecting σ_H .

So we have reduced the problem stated in the theorem to finding a subgraph H of $G(\sigma_{max})$ as described above. If there is such a subgraph H then there is one in form of a tree, where the edges leading to leaves are of the form (m, a, m') with $(m, a) \in P$. Moreover, there is such a tree with at most |P| nodes with more than one child. So finding such a tree can be done by guessing the |P| branching nodes and solving |P| + 1 reachability problems in $G(\sigma_{max})$. This can be done in PTIME since the size of P is bounded by K.

4.2. Soundness of acyclic weakly non-deterministic negotiations. In this section we consider acyclic, weakly non-deterministic negotiations, c.f. page 6. That is, we allow some processes to be non-deterministic, but every atomic negotiation should involve at least one deterministic process.

Definition 4.5. The restriction of a negotiation $\mathcal{N} = \langle Proc, N, dom, R, \delta \rangle$ to a subset $Proc' \subseteq Proc$ of its processes is the negotiation $\langle Proc', N', dom', R, \delta' \rangle$ where $N' = \{n \in N: dom(n) \cap Proc' \neq \emptyset\}$, $dom'(n) = dom(n) \cap Proc'$, and $\delta'(n, r, p) = \delta(n, r, p) \cap N'$. The restriction of \mathcal{N} to its deterministic processes is denoted \mathcal{N}_D .

Since \mathbb{N} is weakly non-deterministic, every atomic negotiation involves a deterministic process, so $N_D = N$. Recall also that for an acyclic negotiation \mathbb{N} we fixed some linear order $\leq_{\mathbb{N}}$ on N, that is a topological order of the graph of \mathbb{N} .

We show that deciding soundness for acyclic, weakly non-deterministic negotiations is in Ptime. The proof is divided into three parts. In Section 4.2.1 we prove some preliminary lemmas. In Section 4.2.2 we consider the special case in which the negotiation \mathcal{N} has one single non-deterministic process, and show that \mathcal{N} is sound iff \mathcal{N}_D is sound and Eve wins a certain instance of the omitting problem for \mathcal{N}_D . Finally, Section 4.2.3 shows how to reduce the general case to the case with only one non-deterministic process.

4.2.1. Preliminaries. We first show two auxiliary lemmas on the structure of runs in negotiations. Two runs w, w' of \mathbb{N} are called equivalent (and we write $w \equiv w'$) if one can be obtained from the other by repeatedly permuting adjacent pairs (m, a), (n, b), with $dom(m) \cap dom(n) = \emptyset$.

The next lemma shows that for acyclic negotiations we can restrict our considerations to runs respecting the order $\leq_{\mathcal{N}}$.

Lemma 4.6. Every run of an acyclic negotiation \mathbb{N} has an equivalent run that respects the topological order $\preceq_{\mathbb{N}}$.

Proof. Let $C_{init} \xrightarrow{(n_0,a_0)\cdots(n_p,a_p)} C$ be some run of $\mathbb N$. The proof is by induction on the number of pairs $0 \le i < j \le p$ such that $n_j \preccurlyeq_{\mathbb N} n_i$. If there are no such pairs, then w' = w. Otherwise, assume that a pair i,j as above exists. Note that in particular, $dom(n_i) \cap dom(n_j) = \emptyset$ holds. It is not hard to see that there exists some $i \le k < j$ such that $n_{k+1} \preccurlyeq_{\mathbb N} n_k$. As before, we note that $dom(n_k) \cap dom(n_{k+1}) = \emptyset$. Clearly, $w' = (n_0, a_0) \cdots (n_{k-1}, a_{k-1})(n_{k+1}, a_{k+1})(n_k, a_k)(n_{k+2}, a_{k+2}) \cdots (n_p, a_p)$ is an equivalent run of $\mathbb N$, with one less pair that violates $\preccurlyeq_{\mathbb N}$.

The next lemma states some properties of runs respecting the order $\leq_{\mathcal{N}}$.

Lemma 4.7. Let \mathbb{N} be an acyclic, weakly non-deterministic negotiation, and consider some reachable configuration C. Let n be the $\leq_{\mathbb{N}}$ -smallest atomic negotiation with C(d) = n for some deterministic process d. The following properties hold:

- (1) If \mathbb{N} is sound then n is enabled in \mathbb{C} .
- (2) For all runs $C \xrightarrow{w} C_{fin}$, all atomic negotiations n' in w are $\leq_{\mathcal{N}}$ -bigger than n.

Proof. For the first item we use the soundness of \mathbb{N} : if n were not enabled in C, then there would exist some n' that is enabled in C and such that the execution of n' makes the execution of n eventually possible. So in particular, we would have $n' \prec_{\mathbb{N}} n$, which contradicts the choice of n.

The second item is shown similarly: every atomic negotiation n' occurring in w satisfies $n'' \preccurlyeq_{\mathbb{N}} n'$ for some n'' that is enabled in C. By the choice of n we have $n \preccurlyeq_{\mathbb{N}} n''$, which shows the claim.

It is easy to see that whenever C is a reachable configuration of a weakly non-deterministic negotiation \mathbb{N} , the restriction C^D of C to deterministic processes is a reachable configuration of \mathbb{N}_D . The next lemma states the converse, lifting runs of \mathbb{N}_D to runs of \mathbb{N} . For this, we need to assume that the runs respect the order $\leq_{\mathbb{N}}$.

Lemma 4.8. Suppose \mathbb{N} is a sound, acyclic and weakly non-deterministic negotiation. Let $C_{init}^D \xrightarrow{w} C^D$ be a run of \mathbb{N}_D respecting the order $\leq_{\mathbb{N}}$. Let also C_v^D denote the configuration reached by some prefix v of w: $C_{init}^D \xrightarrow{v} C_v^D$. Then \mathbb{N} has a run $C_{init} \xrightarrow{v} C_v$ with $C_v(d) = C_v^D(d)$, for every prefix v of w and every deterministic process d.

Proof. The proof is by induction on the length of v. Consider a prefix v(m,a) of w. Recall that w respects the order $\leq_{\mathcal{N}}$, thus m is the $\leq_{\mathcal{N}}$ -smallest atomic negotiation enabled in C_v^D . In particular, Lemma 4.7 (2) applied to \mathcal{N}_D implies that m is the $\leq_{\mathcal{N}}$ -smallest atomic negotiation with $C_v(d) = m$ for some deterministic process. Applying Lemma 4.7 (1) to \mathcal{N} shows finally that m is also enabled in C_v .

The next lemma gives a necessary condition for the soundness of \mathbb{N} that is easy to check. It is proved by showing that \mathbb{N}_D cannot have much more behaviors than \mathbb{N} .

Lemma 4.9. If \mathbb{N} is a sound, acyclic, weakly non-deterministic negotiation then \mathbb{N}_D is sound.

Proof. Suppose to the contrary that \mathcal{N}_D has a run reaching a deadlock configuration $C_{init} \xrightarrow{w} C$. By Lemma 4.6 we can assume that w respects the $\leq_{\mathcal{N}}$ ordering. By Lemma 4.8 we get a run of \mathcal{N} to a deadlock configuration, but this is impossible.

4.2.2. Negotiations with one non-deterministic process. We first consider the case of a negotiation with only one non-deterministic process. The next lemma establishes the connection to the omitting problem: \mathcal{N} is unsound if the deterministic part \mathcal{N}_D is unsound, or \mathcal{N}_D is sound but has a certain successful omitting run.

Lemma 4.10. Let \mathbb{N} be an acyclic, weakly non-deterministic negotiation with a single non-deterministic process p. Then \mathbb{N} is not sound if and only if:

• \mathcal{N}_D is not sound, or

- \mathbb{N}_D is sound, and it has two nodes $m \preceq_{\mathbb{N}} n$ with outcomes $a \in out(m)$, $b \in out(n)$ such that:
 - $-p \in dom(m) \cap dom(n), n \notin \delta(m, a, p), and$
 - there is a successful run of \mathcal{N}_D containing $P = \{(m, a), (n, b)\}$ and omitting $B = \{n' \in \delta(m, a, p) : m \prec_{\mathcal{N}} n' \prec_{\mathcal{N}} n\}$.

Proof. Consider the right-to-left direction. We abbreviate $S_p := \delta(m, a, p)$. If \mathcal{N}_D is not sound then by Lemma 4.9, \mathcal{N} is not sound.

Suppose then that \mathcal{N}_D satisfies the second item from the statement of the lemma, and take a run w of \mathcal{N}_D as it is assumed there. By Lemma 4.6 we can assume that this run respects $\leq_{\mathcal{N}}$. Towards a contradiction suppose also that \mathcal{N} is sound. Lemma 4.8 says that w is also a run of \mathcal{N} . Let C_1 be the configuration of this run just after (m, a) was executed, so we have $C_1(p) = S_p$. Let C_2 be the first configuration after C_1 such that $C_2(d) = n$ for some process d (it may be that $C_2 = C_1$). We have $C_2(p) = S_p$ since the run w omits $\{n' \in S_p : m \prec_{\mathcal{N}_D} n' \prec_{\mathcal{N}_D} n\}$, so p cannot move between C_1 and C_2 . When we continue following w from C_2 we see that d cannot move since n will never be enabled for p. So this run leads to a deadlock, contradiction with the soundness of \mathcal{N} .

For the left-to-right direction, assume that \mathcal{N}_D is sound. We need to show the second item of the lemma. Observe that since \mathcal{N} is acyclic and not sound, there is a run $C_{init} \stackrel{w}{\to} C$ where C is a deadlock. By Lemma 4.6 we can assume that w respects $\leq_{\mathcal{N}}$. Let n be the $\leq_{\mathcal{N}}$ -smallest atomic negotiation such that C(d) = n for some deterministic process d. Applying Lemma 4.7 (1) to \mathcal{N}_D yields that n must be deterministically enabled in C: that is C(d') = n for all deterministic processes $d' \in dom(n)$. This implies $p \in dom(n)$, and $n \notin C(p)$.

Let us split w as u(m,a)v where $p \in dom(m)$ and all atomic negotiations in v involving p are $\preceq_{\mathbb{N}}$ -bigger than n (it may be that v is empty). That is to say, we define m as the last atomic negotiation in w involving p that is $\preceq_{\mathbb{N}}$ -smaller than n. Let C_m be the configuration reached after doing (m,a): $C_{init} \xrightarrow{u(m,a)} C_m$. Take $S_p = C_m(p) = \delta(m,a,p)$. By the choice of m and v, the run from C_m to C does not use any atomic negotiation from the set $\{n' \in S_p : m \prec_{\mathbb{N}} n' \prec_{\mathbb{N}} n\}$.

By soundness of \mathbb{N}_D , from C^D there is a run to the final configuration, and by the choice of n and Lemma 4.7 (2) this run cannot use any atomic negotiation that is $\leq_{\mathbb{N}^-}$ smaller than n. Let b be the outcome such that (n,b) appears in this run. Putting these pieces together we have a successful run of \mathbb{N}_D containing $\{(m,a),(n,b)\}$ and omitting $\{n' \in S_p : m \prec_{\mathbb{N}} n' \prec_{\mathbb{N}} n\}$. We have already observed that $p \in dom(m) \cap dom(n)$ and $n \notin S_p = \delta(m,a,p)$. So all the requirements of the lemma are met.

Lemma 4.11. Soundness of acyclic, weakly non-deterministic negotiation with only one non-deterministic process can be checked in Ptime.

Proof. For every $m \leq n$, a and b we check the conditions described in Lemma 4.10. The existence of a run of \mathcal{N}_D can be checked in Ptime thanks to Theorem 4.4 and the fact that the size of P is always 2.

4.2.3. General weakly non-deterministic negotiations. The next lemma deals with the case where there is more than one non-deterministic process. Loosely speaking, in this case \mathcal{N} is unsound iff there is a non-deterministic process such that the restriction of \mathcal{N} to the deterministic processes and this process is unsound.

Lemma 4.12. An acyclic, weakly non-deterministic negotiation \mathbb{N} is unsound iff:

- (1) its restriction \mathcal{N}_D to deterministic processes is unsound, or
- (2) for some non-deterministic process p, its restriction \mathbb{N}^p to p and the deterministic processes is unsound.

Proof. For the right-to-left direction the case where \mathcal{N}_D is unsound follows directly from Lemma 4.9. It remains to check the case where \mathcal{N}^p is not sound for some non-deterministic process p. Consider a run $C^p_{init} \xrightarrow{w} C^p$ where C^p is a deadlock. By Lemma 4.6 we can assume that w respects $\preceq_{\mathcal{N}}$. Now let us try to make \mathcal{N} execute the sequence w.

If $C_{init} \xrightarrow{w} C$ is a run of \mathbb{N} then C is a deadlock. Indeed if in \mathbb{N} it would be possible to do $C \xrightarrow{(n,a)}$ for some (n,a) then $C^p \xrightarrow{(n,a)}$ would be possible in \mathbb{N}^p .

The other case is when in \mathbb{N} it is not possible to execute all the sequence w. Then we have w = v(n,a)v', a run $C_{init} \stackrel{v}{\to} C_1$ and from C_1 action (n,a) is not possible. Since $C_{init}^p \stackrel{v}{\to} C_1^p \stackrel{(n,a)}{\longrightarrow} C_2^p$ is a run of \mathbb{N}^p , we know from Lemma 4.7 (2) that there is a deterministic process d with $C_1(d) = n$, and $n \leq_{\mathbb{N}} C_1(d')$ for all other deterministic processes d'. Thus C_1 is a deadlock because by Lemma 4.7 (1) if there is an action possible from C_1 then this must be n.

For the left-to-right direction, suppose that \mathbb{N} is not sound and take a run $C_{init} \xrightarrow{w} C$ with C a deadlock configuration. Take the $\preceq_{\mathbb{N}}$ -smallest atomic negotiation n such that n = C(d) for some deterministic process d. Consider an arbitrary non-deterministic process p and a run $C_{init}^p \xrightarrow{w} C^p$ of \mathbb{N}^p ; it is indeed a run since \mathbb{N}^p is a restriction of \mathbb{N} . As \mathbb{N}^p is sound, it is possible to extend this run. By Lemma 4.7 (1), it should be possible to execute n from C^p . Hence for every deterministic process $d \in dom(n)$, we have C(d) = n. Moreover $n \in C(p)$ if $p \in dom(n)$ is non-deterministic. Since the choice of p was arbitrary, we have $n \in C(p)$ for all $p \in dom(n)$. Thus it is possible to execute n from C, a contradiction. \square

Theorem 4.13. Soundness can be decided in PTIME for acyclic, weakly non-deterministic negotiations.

Proof. By Lemma 4.12 we can restrict to negotiations \mathcal{N} with one non-deterministic process. For every $m \preccurlyeq_{\mathcal{N}} n$, a and b we check the conditions described in Lemma 4.10. The existence of a run of \mathcal{N}_D can be checked in PTIME thanks to Theorem 4.4 and the fact that the size of P is constant.

5. Beyond determinism: Intractable cases

We show that if we remove any of the two assumptions of Theorem 4.13 (acyclicity and weak non-determinism) then the soundness problem becomes CONP-complete. In fact, even a very mild relaxation of acyclicity suffices.

It is not very surprising that deciding soundness for acyclic, non-deterministic negotiations is CoNP-complete. In the acyclic case the negotiation is unsound iff it can reach a deadlock. Clearly, by acyclicity this would happen after at mots |N| steps. So it suffices to guess a run of linear length step by step and check if it leads to a deadlock. CoNP-hardness is proved by a simple reduction of SAT to the complement of the soundness problem. The reduction, which strongly relies on non-determinism, is presented in [3], and we get:

Proposition 5.1 ([3]). Soundness of acyclic non-deterministic negotiations is CONP-complete.

Now we consider a very mild relaxation of acyclicity: deterministic processes still need to be acyclic, but non-deterministic processes may have cycles.

Recall that \mathcal{N}_D is the restriction of \mathcal{N} to deterministic processes.

Definition 5.2. A negotiation \mathcal{N} is *det-acyclic* if \mathcal{N}_D is acyclic.

It follows easily from this definition that all runs of a weakly non-deterministic, detacyclic negotiation have length at most |N|. However, we show that even in the *very* weakly non-deterministic case (c.f. page 6) the soundness problem is CONP-complete.

Theorem 5.3. Non-soundness of det-acyclic, very weakly non-deterministic negotiations is NP-complete.

Proof. We describe a reduction from 3-SAT and fix a 3-CNF formula $\varphi = c_1 \wedge \cdots \wedge c_m$, with clauses c_1, \ldots, c_m , each of length 3, and k variables x_1, \ldots, x_k . Let $c_j = \ell_{j,1} \vee \ell_{j,2} \vee \ell_{j,3}$. We construct a det-acyclic, very weakly non-deterministic negotiation \mathcal{N} such that φ is satisfiable iff \mathcal{N} is not sound.

The atomic negotiations of \mathcal{N} are (apart from n_{init} and n_{fin}):

- An atomic negotiation m_0 , and for each variable x_i three atomic negotiations n_i^+, n_i^-, m_i .
- For every pair clause/literal (j, d), j = 1, ..., m and d = 1, 2, 3, three atomic negotiations $m_{j,d}$, $n_{j,d}$, and $r_{j,d}$.
- For every clause c_i , two auxiliary atomic negotiations t_i and t'_i .

The processes of \mathbb{N} are:

- A deterministic process E.
- For each clause c_i , a deterministic process V_i .
- For every pair clause/literal (j, d), where $j = 1, \ldots, m$ and d = 1, 2, 3: two deterministic processes $T_{j,d}, T'_{j,d}$, and a non-deterministic process $P_{j,d}$.

Now we describe the behavior of each process P by means of a graph. The nodes of the graph for P are the atomic negotiations in which P participates. The graph has an edge $n \to n'$ if there is an outcome a of n such that P moves with a from n to n'. If P is nondeterministic, and after a is ready to engage in a set of atomic negotiations $\{n_1, \ldots, n_k\}$, then the graph contains a hyperarc leading from n to $\{n_1, \ldots, n_k\}$.

The graphs of all processes are shown in Figure 4. Intuitively, process E is in charge of producing a valuation of x_1, \ldots, x_k : it chooses between n_1^+ and n_1^- , then between n_2^+ and n_2^- , etc. Choosing n_i^+ stands for setting x_i to true, and choosing n_i^- for setting x_i to false.

Observe that in the graph for the non-deterministic process $P_{j,d}$ we assume $\ell_{j,d} \in \{x_i, \overline{x_i}\}$. After n_{init} , the process goes to m_i , and then to n_i^+ or n_i^- in a deterministic way, depending on the outcome chosen at m_i . The rest of its behavior depends on whether $\ell_{j,d} = x_i$ or $\ell_{j,d} = \overline{x_i}$. If $\ell_{j,d} = x_i$, then after n_i^+ the process goes to to $r_{j,d}$, and then to one of $\{n_{fin}, n_{j,d}\}$ (nondeterminism!). For the other case, see Figure 4.

Process $P_{j,d}$ is designed with the following purpose. If process E sets literal $\ell_{j,d}$ to true, then $P_{j,d}$ (together with $T_{j,d}$) guarantees that $m_{j,d}$ is executed before $n_{j,d}$; if E sets literal $\ell_{j,d}$ to false, then $m_{j,d}$ and $n_{j,d}$ can occur in any order. In other words, in every successful run containing n_i^+ : if $\ell_{j,d} = x_i$ then node $m_{j,d}$ appears before $n_{j,d}$; if $\ell_{j,d} = \overline{x_i}$ then the nodes $m_{j,d}$ and $n_{j,d}$ can appear in any order. Similarly for runs containing n_i^- , interchanging $x_i, \overline{x_i}$.

It is easy to see that \mathcal{N} is very weakly non-deterministic. The only nondeterministic processes are the $P_{j,d}$ processes. Moreover, the sets $\{n_{fin}, r_{j,d}\}$ and $\{n_{fin}, n_{j,d}\}$ are the only

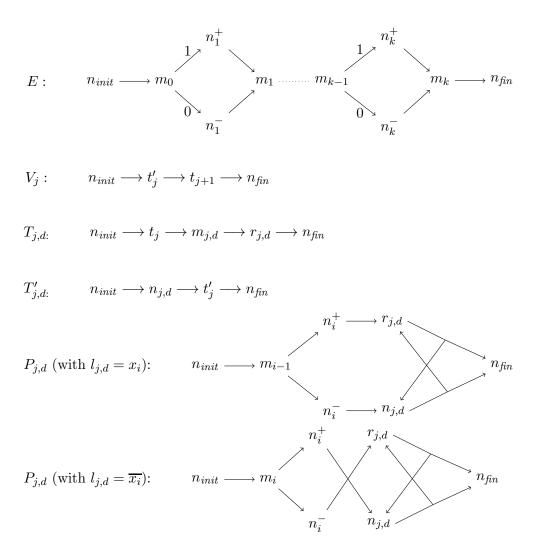


Figure 4: Graphs of the processes of \mathbb{N} .

two sets of atomic negotiations such that (a) there are configurations such that $P_{j,d}$ is ready to engage in them, and (b) contain more than one element. Since n_{fin} contains all processes, the condition for very weak non-determinism is clearly verified.

If the valuation chosen by E makes all c_j true, we claim that the partial run corresponding to this valuation cannot be completed to a successful run. Indeed, in this case for each c_j there is a true literal $\ell_{j,d}$, and so $P_{j,d}$ enforces that $m_{j,d}$ is executed before $n_{j,d}$. We denote this by $m_{j,d} < n_{j,d}$. But then, since process $T_{j,d}$ enforces $t_j < m_{j,d}$, process $T'_{j,d}$ enforces $n_{j,d} < t'_j$, and process V_j enforces $t'_j < t_{j+1 \mod (m+1)}$, we get the cycle

$$t_1 < t_1' < t_2 < \cdots t_m < t_m' < t_1$$

Since, say, t_2 cannot occur before and after t_1 , because the deterministic process are acyclic, the partial run cannot be completed.

Otherwise, if the valuation makes at least one c_j false, then no cycle is created and the partial run can be extended to a successful run.

Here is a more formal version of the proof. By construction, $dom(t_j) = \{V_{j-1}, T_{j,d} \mid d = 1, 2, 3\}, dom(t'_j) = \{V_j, T'_{j,d} \mid d = 1, 2, 3\}, dom(m_{j,d}) = \{T_{j,d}\}, dom(n_{j,d}) = \{T'_{j,d}, P_{j,d}\}, dom(r_{j,d}) = \{T_{j,d}, P_{j,d}\}.$

Let $\nu: \{1, \ldots, k\} \to \{0, 1\}$ be a valuation of the variables x_1, \ldots, x_k . By C_{ν} we denote the following configuration (note that only the position of processes $P_{j,d}$ depends on the valuation):

- $C_{\nu}(E) = n_{fin}$
- If the literal $\ell_{i,d}$ is true under ν then $C_{\nu}(P_{i,d}) = r_{i,d}$, and otherwise $C_{\nu}(P_{i,d}) = n_{i,d}$.
- $C_{\nu}(T_{j,d}) = t_j, C_{\nu}(T'_{i,d}) = n_{j,d}$
- $C_{\nu}(V_j) = t'_i$

The following property is easy to check:

Fact 1. Every configuration C_{ν} is reachable. Moreover, every maximal run has some trace-equivalent prefix that reaches one of the configurations C_{ν} .

Fact 2. If ν does not satisfy the formula then there is a successful run from C_{ν} .

Assume that ν does not satisfy clause c_j . By Fact 1, nodes $n_{j,d}$, d=1,2,3, are enabled. By executing these nodes we can reach a configuration with $P_{j,d}$ in $\{r_{j,d}, n_{fin}\}$ and $T'_{j,d}$ in t'_j . Now t'_j is executable, and V_j moves to node t_{j+1} . Notice that t_{j+1} is now executable, so $T_{j+1,d}$ can move first to $m_{j+1,d}$, then to $r_{j+1,d}$, d=1,2,3 (and V_j goes to n_{fin}). Now either $P_{j+1,d}$ is already in $r_{j+1,d}$ or it can come there after $n_{j+1,d}$ is executed. In the first case nodes $r_{j+1,d}, n_{j+1,d}$ can be executed (in this order), in the second case $n_{j+1,d}, r_{j+1,d}$ can be executed (in this order). In either case we can get to a configuration where processes $T_{j+1,d}$ are in n_{fin} and processes $T'_{j+1,d}$ are in t'_{j+1} . Therefore, t'_{j+1} is also executable. By iterating this argument we obtain a successful run executing $t'_j, t_{j+1}, t'_{j+1}, \ldots, t_1, t'_1, \ldots, t_j$ in this order.

Fact 3. If ν does satisfy the formula then there is no successful run from C_{ν} .

Suppose by contradiction that there is a successful run σ from C_{ν} . By assumption, for each j there is some d=1,2,3 such that $C_{\nu}(P_{j,d})=r_{j,d}$. Therefore $m_{j,d}$ is necessarily executed before $n_{j,d}$ in σ . By construction this implies that t_j is executed before t'_j in σ . Also by construction, t'_j must be executed before $t_{(j+1)mod(m+1)}$, for every j. This means that t_1 should be executed before t'_m , and t'_m before t_1 , a contradiction.

6. Beyond soundness

We show that the techniques we have developed for the soundness problem, like forks and the omitting game, can be used to check other functional properties of acyclic deterministic negotiations. Moreover, we show that soundness reduces the complexity of verifying these properties: while checking the property for arbitrary deterministic negotiations—sound or not—is an intractable problem, it is polynomial in the sound case.

In Section 6.1 we study the race problem: Given two atomic negotiations m and n, is there a run in which they occur concurrently? In Section 6.2 we address the static analysis of negotiations with data. We assume that the outcomes of a negotiation correspond to operations acting on a set of variables, and study some standard questions of static analysis, for instance whether a variable can be allocated and then never deallocated before the execution ends. These questions can be answered in exponential time by constructing the reachability graph of the negotiation and applying standard model checking algorithms.

(This is the approach followed in [14], which studies the questions for workflow Petri nets.) We exhibit polynomial algorithms for the acyclic deterministic case.

6.1. **Races.** For a given pair of atomic negotiations m, n of a deterministic negotiation $\mathcal{N} = \langle Proc, N, dom, R, \delta \rangle$, we want to determine if there is a race between m and n, i.e., if there is a reachable configuration that concurrently enables m and n.

Definition 6.1. Let $\mathbb{N} = \langle Proc, N, dom, R, \delta \rangle$ be a negotiation. Two atomic negotiations $m, n \in N$ can be concurrently enabled, denoted $m \parallel n$, if $dom(m) \cap dom(n) = \emptyset$ and there is a reachable configuration C of \mathbb{N} where both m and n are enabled.

This question was answered in [9] for live and safe free-choice nets, where a polynomial fixed point algorithm was given. The algorithm can also be applied to sound deterministic negotiations (cyclic or acyclic), but has cubic complexity in the number of atomic negotiations. We show that in the acyclic case there is a simple anti-pattern characterization of the race pairs m, n, which leads to a algorithm that runs in linear time and logarithmic space.

In the rest of the section we give a syntactic characterization of the pairs m, n such that $m \parallel n$. We proceed in several steps. Lemma 6.2 massages the semantic definition into an equivalent, more suitable condition, but still semantic. Proposition 6.3 transforms this condition into the conjunction of a syntactic and a semantic condition. Proposition 6.4 replaces the latter by the existence of a certain fork. The final result, given in Theorem 6.5, just puts the two propositions together.

Recall that two runs $w, w' \in (N \times R)^*$ are equivalent if w' can be obtained from w by repeatedly exchanging adjacent pairs (m, a)(n, b) into (n, b)(m, a) whenever $dom(m) \cap dom(n) = \emptyset$.

Lemma 6.2. Let \mathbb{N} be an acyclic, deterministic, sound negotiation, and let m, n be two atomic negotiations in \mathbb{N} . Then $m \parallel n$ iff every run w from n_{init} containing both m and n has an equivalent run $w' = w_1 w_2$ such that $w' = C_{init} \xrightarrow{w_1} C \xrightarrow{w_2} C'$ for some configuration C where both m and n are enabled.

Proof. It suffices to show the implication from left to right. So assume that there exists some reachable configuration C where both m and n are enabled. In particular, $dom(m) \cap dom(n) = \emptyset$. By way of contradiction, let us suppose that there exists some run containing both m and n, but this run cannot be reordered as claimed. We claim that there must be some local path from m to n in \mathbb{N} . To see this, assume the contrary and consider a run of the form $w = w_1(m, a)w_2(n, b)w_3$. The run w defines a partial order (actually a Mazurkiewicz trace) tr(w) with nodes corresponding to positions in w, and edges from (m', c) to (n', d) if $dom(m') \cap dom(n') \neq \emptyset$ and (m', c) precedes (n', d) in w. Since there is no path from m to n in \mathbb{N} , nodes (m, a) and (n, b) are unordered in tr(w). So we can choose a topological order w' of tr(w) of the form $w' = w'_1(m, a)(n, b)w'_2$. This shows the claim.

So let π be a path in \mathbb{N} from m, n_1, \ldots, n_k, n . Let p be some process such that $n_k \xrightarrow{p,a'} n$ for some outcome a'.

Let us go back to C. Since both m and n are enabled in C, we have a transition $C \xrightarrow{n,b} C_1$, for some $b \in out(n)$. Note that m is still enabled in C_1 , since $dom(m) \cap dom(n) = \emptyset$. So we can apply Lemma 3.2 to C_1 and π (because \mathbb{N} is sound), obtaining a configuration C_2 where $C_2(p) = n$. But since n was executed before C_1 , this violates the acyclicity of \mathbb{N} . \square

The next step is to convert the condition from Lemma 6.2 to a condition on the graph of a negotiation.

Proposition 6.3. Let \mathbb{N} be an acyclic, deterministic, sound negotiation, and let m, n be two atomic negotiations in \mathbb{N} with $dom(m) \cap dom(n) = \emptyset$. Then $m \parallel n$ iff there exists an initial run containing both m, n, and there is neither a local path from m to n nor a local path from n to m.

Proof. For the left-to-right implication, assume by contradiction that there is some local path π from m to n. Consider some reachable configuration C such that m is enabled in C. By Lemma 3.2 we also find a run $C \stackrel{*}{\to} C'$ such that n is enabled in C'. But note that the run $C_{init} \stackrel{*}{\to} C \stackrel{*}{\to} C'$ cannot be reordered as stated in Lemma 6.2, a contradiction.

For the converse, consider some run w containing both m, n. Since there are no local paths in \mathbb{N} between m, n, we can reorder, as in the proof of Lemma 6.2, the run w into some w' such that we find a configuration C of w' where both m and n are enabled.

Proposition 6.4. Let \mathbb{N} be a sound deterministic negotiation, and let m, n be two atomic negotiations of \mathbb{N} . Then \mathbb{N} has an initial run containing both m and n iff it has a fork (p, q, m, n) for some p, q.

Proof. Right-to-left implication: the proof is similar to the one of Lemma 3.2, but we need to consider three paths instead of a single one. First we realize the path from n_{init} to the branching node n' (with outcome a) in the definition of fork, using Lemma 3.2. Suppose that the run from C_{init} to the configuration C that enables n', contains neither m nor n (otherwise another application of Lemma 3.2 suffices). Let $C \xrightarrow{(n',a)} C_1$. We show by induction on the sum of the lengths of the two local paths of the fork how to construct a run containing both m and n. Let $C_1(p) = m_0$, $C_1(q) = n_0$ and let

$$m_0 \xrightarrow{p,a_0} \cdots \xrightarrow{p,a_{k-1}} m_k = m$$
 and $n_0 \xrightarrow{q,b_0} \cdots \xrightarrow{q,b_{l-1}} n_l = n$.

be the local paths of the fork. Since $\mathbb N$ is sound, some run leads from C_1 to the final configuration. Since $\mathbb N$ is deterministic, the run must execute both m_0 and n_0 at some point. Suppose without loss of generality that m_0 is executed first. We have $C_1 \stackrel{*}{\to} C_2 \stackrel{(m_0, a_0)}{\to} C_3$ for some C_2, C_3 such that $C_3(p) = m_1$ (since $\mathbb N$ is deterministic) and $C_3(q) = n_0$. We can now apply the induction hypothesis to the local paths $m_1 \stackrel{*}{\to} m$, $n_0 \stackrel{*}{\to} n$, and we are done.

Left-to-right implication: By assumption \mathbb{N} has a run from n_{init} of the form $w = w_1(m,b)w_2(n,c)$. Choose some $p \in dom(m), q \in dom(n)$. Let (n',a) be the rightmost letter of w_1 such that $\{p,q\} \subseteq dom(n_0)$ (which exists because $\{p,q\} \subseteq dom(n_{init})$ by definition), and let $m_0 = \delta(n',a,p), n_0 = \delta(n',a,q)$. Then we can extract from w_1 a p-path leading from m_0 to m, and from w_1w_2 a q-path leading from n_0 to n. By the choice of n' these two paths are disjoint.

From Proposition 6.3 and 6.4 we immediately obtain:

Theorem 6.5. For any acyclic, deterministic, sound negotiation \mathbb{N} we can decide in linear time (resp., in logarithmic space) whether two atomic negotiations m, n of \mathbb{N} satisfy $m \parallel n$. The above problem is NLOGSPACE-complete.

It is not difficult to show that soundness is essential for this result. Indeed, the race problem is NP-hard for acyclic and deterministic, but not necessarily sound negotiations. We sketch a reduction from SAT.

Theorem 1 of [3] shows how to reduce CNF-SAT to the non-soundness problem for acyclic negotiations. Loosely speaking, given a formula ϕ , the reduction constructs a negotiation \mathcal{N}_{ϕ} having a run for each truth assignment. The run for a given assignment can reach the final configuration iff it makes at least one clause of ϕ false, otherwise it gets stuck in a deadlock. Therefore, \mathcal{N}_{ϕ} is sound iff ϕ is unsatisfiable.

It is easy to add two atomic negotiations m, n to \mathcal{N}_{ϕ} in such a way that they can only be executed if the deadlock is reached, and in this case they are executed concurrently. Therefore, there is a race between m and n iff ϕ is satisfiable. This reduces SAT to the race problem for acyclic negotiations, but not necessarily deterministic.

Let \mathcal{N}'_{ϕ} be the negotiation obtained by adding n and m. We apply a construction that transforms it into a deterministic negotiation \mathcal{N}''_{ϕ} . Loosely speaking, the transformation preserves the reachable configurations, and so the races, but it may introduce many deadlocks, and so it does not preserve soundness. It is defined as follows: for every entry $\delta'(n,a,p)=\{n_1,\ldots,n_k\}$ of \mathcal{N}'_{ϕ} such that $k\geq 2$, add a fresh atomic negotiation [n,a,p] with domain $\{p\}$ and outcomes a_1,\ldots,a_n , and replace the entry by $\delta''(n,a,p)=\{[n,a,p]\}$ and $\delta''([n,a,p],a_i,p)=\{n_i\}$. Loosely speaking, while in \mathcal{N}'_{ϕ} process p offers to participate in any of $\{n_1,\ldots,n_k\}$, and leaves the choice of one of these nodes to the environment, in \mathcal{N}''_{ϕ} process p unilaterally commits to one of $\{n_1,\ldots,n_k\}$. Observe that \mathcal{N}''_{ϕ} is deterministic. Further, a configuration of \mathcal{N}'_{ϕ} is reachable in \mathcal{N}'_{ϕ} iff it is reachable in \mathcal{N}''_{ϕ} . So there is a race betwen n and m in \mathcal{N}''_{ϕ} iff ϕ is satisfiable.

6.2. **Negotiations with data.** A negotiation with data is a negotiation over a given set X of variables (over finite domains), where each outcome $(n, a) \in N \times R$ comes with a set Σ of operations on the shared variables. In our examples this set Σ is composed of alloc(x), read(x), write(x), and dealloc(x).

Formally, a negotiation with data is a negotiation with one additional component: $\mathcal{N} = \langle Proc, N, dom, R, \delta, \ell \rangle$ where $\ell \colon (N \times R) \to \mathcal{P}(\Sigma \times X)$ maps every outcome to a (possibly empty) set of data operations on variables from X. We assume that for each $(n, a) \in N \times R$ and for each variable $x \in X$ the label $\ell(n, a)$ contains at most one operation on x, that is, at most one element of $\Sigma \times \{x\}$.

As an example, we enrich the deterministic negotiation on the left Figure 1 with data, as shown in Table 1. (This example is adapted from [13]) The variables are x_1, x_2 . The table gives for each outcome and for each operation the (indices of the) variables to which the outcome applies this operation.

atom. neg.	n_0	n_1	n_2	n_3		n_4		n_5
outcome	a	a	a	a	b	a	b	a
alloc	1, 2							
read			2	1			2	
write		1			2	2		
dealloc							1	2

Table 1: Data for the negotiation on the left of Figure 1.

In [13] some examples of data specifications for workflows are considered.

- (1) Inconsistent data: an atomic negotiation reads or writes a variable x while another atomic negotiation is writing, allocating, or deallocating it in parallel. In our example there is a run in which (n_2, a) and (n_3, b) read and write to x_2 in parallel.
- (2) Weakly redundant data: there is a run in which a variable is written and never read before it is deallocated or the run ends. In the example, there is a run in which x_2 is written by (n_3, b) , and never read again.
- (3) Never destroyed: there is an execution in which a variable is allocated and then never deallocated before the execution ends. The example has a run which never takes (n_4, b) , in which x_1 is never deallocated.

It is easy to give algorithms for these properties that are polynomial in the size of the reachability graph. We give the first algorithms that check these properties in polynomial time in the size of the negotiation, which can be exponentially smaller than its reachability graph. For the first property we can directly use the algorithm for the race problem: It suffices to check if the negotiation has two outcomes (m,a),(n,b) such that $m \parallel n$ and there is a variable x such that $\ell(a) \cap \{read(x), write(x)\} \neq \emptyset$ and $\ell(b) \cap \{write(x), alloc(x), dealloc(x)\} \neq \emptyset$. In the rest of the section we present a polynomial algorithm for the following abstract problem, which has the problems (2) and (3) above as special instances.

Given sets $\mathcal{O}_1, \mathcal{O}_2, \mathcal{O} \subseteq N \times R$ we say that the negotiation \mathcal{N} violates the specification $(\mathcal{O}_1, \mathcal{O}_2, \mathcal{O})$ if there is a successful run $w = (n_1, a_1) \cdots (n_k, a_k)$ of \mathcal{N} , and indices $0 \le i < j \le k$ such that $(n_i, a_i) \in \mathcal{O}_1$, $(n_j, a_j) \in \mathcal{O}_2$, and $(n_l, a_l) \notin \mathcal{O}$ for all i < l < j. In this case we also say that $(\mathcal{O}_1, \mathcal{O}_2, \mathcal{O})$ is violated at $(n_i, a_i), (n_j, a_j)$. Otherwise \mathcal{N} complies with $(\mathcal{O}_1, \mathcal{O}_2, \mathcal{O})$.

Example 6.6. In order to simplify notations below, we assume that n_{fin} has some outcome - if not, we can add a self-loop to n_{fin} .

Observe that a variable x is weakly redundant (specification of type (2)) iff \mathbb{N} violates $(\mathcal{O}_1, \mathcal{O}_2, \mathcal{O})$, where $\mathcal{O}_1 = \{(n, a) \in \mathbb{N} \times \mathbb{R} : write(x) \in \ell(n, a)\}, \mathcal{O}_2 = \{(n, b) \in \mathbb{N} \times \mathbb{R} : n = n_{fin} \lor dealloc(x) \in \ell(n, b)\}$ and $\mathcal{O} = \{(n, c) : \ell(n, c) \cap (\Sigma \times \{x\}) \neq \emptyset\}.$

A variable x is never destroyed (specification of type (3)) iff \mathbb{N} violates $(\mathcal{O}_1, \mathcal{O}_2, \mathcal{O})$, where $\mathcal{O}_1 = \{(n, a) \in N \times R : alloc(x) \in \ell(n, a)\}$, $\mathcal{O}_2 = \{n_{fin}\}$, $\mathcal{O} = \{(n, c) : n = n_{fin} \lor \ell(n, c) \cap \{alloc(x), dealloc(x)\} \neq \emptyset\}$.

For the next proposition it is convenient to use the notation $(m, a) \xrightarrow{+} (n, b)$, for $m, n \in N$, $a \in out(m)$, $b \in out(n)$ whenever there is a (non-empty) local path in \mathbb{N} from $\delta(m, a, p)$ to n, for some $p \in dom(m)$.

Proposition 6.7. Let \mathbb{N} be an acyclic, deterministic, sound negotiation with data, and $(\mathfrak{O}_1, \mathfrak{O}_2, \mathfrak{O})$ a specification. Let $(m, a) \in \mathfrak{O}_1$, $(n, b) \in \mathfrak{O}_2$. Then \mathbb{N} violates $(\mathfrak{O}_1, \mathfrak{O}_2, \mathfrak{O})$ at (m, a), (n, b) iff $(n, b) \not\xrightarrow{+} (m, a)$ and \mathbb{N} has a successful run containing $P = \{(m, a), (n, b)\}$, and omitting the set $B = \{(m', c) \in \mathfrak{O} : (m, a) \xrightarrow{+} (m', c) \xrightarrow{+} (n, b)\}$.

Proof. For the right-to-left direction: assume that \mathbb{N} has a run w as claimed. Since $(n,b) \not\to (m,a)$, w can be assumed to be of the form $w = w_1(m,a)w_2(n,b)w_3$. By reordering w we may suppose that for every (m',c) in w_2 , we have $(m,a) \xrightarrow{+} (m',c) \xrightarrow{+} (n,b)$. Thus, since w omits B this means that $(m',c) \notin \mathbb{O}$, so the claim follows.

For the left-to-right direction: if \mathbb{N} violates $(\mathcal{O}_1, \mathcal{O}_2, \mathcal{O})$ at (m, a), (n, b) then there is a run $w = (n_1, a_1) \cdots (n_k, a_k)$ with $(n_i, a_i) = (m, a), (n_j, a_j) = (n, b)$ and such that $(n_l, a_l) \notin \mathcal{O}$

for all i < l < j. Since $(\{(m',c) : (m,a) \xrightarrow{+} (m',c) \xrightarrow{+} (n,b)\} \cap \{n_i : 1 \le i \le k\}) \subseteq \{n_l : i < l < j\}$, the run w contains (m,a),(n,b) and omits B.

Remark 6.8. Note that Proposition 6.7 refers to omitting pairs of atomic negotiation/outcome, whereas the original omitting problem refers to omitting atomic negotiations. However, it is straightforward to adapt the omitting problem and the corresponding results as to handle pairs.

Putting together Proposition 6.7 and Theorem 4.4 we obtain:

Corollary 6.9. Given an acyclic, deterministic, sound negotiation with data \mathbb{N} , and a specification $(\mathcal{O}_1, \mathcal{O}_2, \mathcal{O})$, it can be checked in polynomial time whether \mathbb{N} complies with $(\mathcal{O}_1, \mathcal{O}_2, \mathcal{O})$.

7. Conclusions

Verification questions for finite-state of concurrent systems are very often PSPACE-hard because of the state explosion problem. One approach to this challenge is to search for restrictions in the communication primitives that permit non-trivial interactions, and yet are algorithmically easier to analyze. We have shown that negotiations are a suitable model for this line of attack. On the one hand, non-deterministic processes can simulate any other communication primitive (up to reasonable equivalences); on the other hand, the definition of non-determinism immediately suggests to investigate the deterministic and weakly non-deterministic classes. Even the deterministic negotiations have enough expressive power for interesting applications in the workflow modeling domain. Indeed, they are equivalent to workflow free-choice nets [1], and acyclicity and free-choiceness are quite common: about 70% of the industrial workflow nets of [16, 6, 5] are free-choice, and about 60% are both acyclic and free-choice (see e.g. the table at the end of [5]). We have shown that a number of verification problems for sound deterministic negotiations can be solved in PTIME or even in NLOGSPACE.

There are several open problems we are currently working on, or intend to address. We do not know if the soundness problem for acyclic weakly non-deterministic negotiations is PTIME-complete. Also, we conjecture that the polynomial algorithms for acyclic deterministic negotiations of Section 6 can be extended to the cyclic case. More generally, we would like to have a better understanding of which verification problems for sound deterministic negotiations can be solved in PTIME. Since these negotiations are not closed under products with automata, we should not expect to be able to polynomially decide arbitrary safety properties. Finally, the analogous question for sound weakly non-deterministic negotiations and the class NPis also increasingly interesting, due to the important advances in SMT tools.

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