Pattern Matching and Membership for Hierarchical Message Sequence Charts

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Abstract. Several formalisms and tools for software development use hierarchy for system design, for instance statecharts and diagrams in UML. Message sequence charts (MSCs) are a standardized notation for asynchronously communicating processes. The norm Z.120 includes also hierarchical HMSCs in form of High-level MSCs (HMSC). Algorithms on MSCs rarely take into account all possibilities covered by the norm. In particular, hierarchy is not taken into account since the model usually considered are MSC-graphs that correspond to the unfolding of (hierarchical) HMSCs. However, complexity can increase exponentially by unfolding. The aim of this paper is to show that basic algorithms can be designed such that they avoid the costly unfolding of hierarchical MSCs and HMSCs. We consider the membership and the pattern matching problem to illustrate the way to proceed. We show that the membership problem for hierarchical HMSCs is PSPACE-complete. Second, we describe a polynomial-time algorithm for the pattern-matching problem on hierarchical MSCs.

1 Introduction

It is common to use macros to write a program or to specify the behavior of a system. Macros or hierarchical specifications allow a modular design of complex systems and have the advantage of being more succinct and user-friendly. Several formalisms and tools for software development use hierarchy for system design. One of the most prominent examples is the formalism of statecharts [11], which is a component of several object-oriented notations, such as the Unified Modeling Language (UML). Besides statecharts, UML widely uses several kinds of diagrams (activity, interaction diagrams etc), all based on the ITU standard Z.120 of message sequence charts (MSCs). While statecharts extend finite state machines by hierarchy and communication mechanisms, MSCs are a visual notation for asynchronously communicating processes. The usual application of MSCs in telecommunication is for capturing requirements of communication protocols in form of scenarios in early design stages. MSCs usually represent incomplete specifications, obtained from a preliminary view of the system that abstracts away several details such as variables or message contents. High-level MSCs (HMSCs) combine basic MSCs using choice and iteration, thus describing possibly infinite collections of scenarios. For abstract specifications as with HMSCs, hierarchy is of primary importance. Since a scenario corresponds to a specification level which can be very abstract, a designer should be able to merge different specification cases yielding the same abstract scenario and to use this scenario as a macro. By using macros designers may identify sub scenarios which have to be refined at a later stage.

Algorithms on MSCs rarely take into account the whole spectrum of the HMSC standard definition. In particular, hierarchy is not taken into account since the models usually considered are MSC-graphs, that correspond to the unfolding of (hierarchical) HMSCs. However, complexity can increase exponentially by unfolding. The aim of this paper is to show that this exponential blow-up is avoidable in many cases, by avoiding the expensive unfolding and using the hierarchy for computing the desired results in a modular way. We use techniques stemming from combinatorics on compressed texts, since hierarchical MSC definitions can be seen as a kind of compression by means of Straight-Line Programs (SLP).

In this paper we consider two fundamental problems for *hierarchical* HMSCs, that are called here *nested high-level MSCs* (nHMSCs for short): membership problem and pattern matching. However, we think the techniques described here can be used to solve other algorithmic problems on nHMSCs as well. The membership problem is a basic question, asking for instance whether a negative scenario occurs in a system specification, or asking whether a positive scenario is redundant, since already covered by the specification. Without hierarchy, the membership problem for HMSCs has been shown to be NP-complete, [1]. The reason for this complexity blow-up (compared to finite-state machines) is that MSCs are partial-order models. We show that hierarchy yields a small increase in complexity, precisely we show that the membership problem for nHMSCs is PSPACE-complete. Surprisingly, hierarchy alone is the source of the complexity. We show namely that the membership problem for hierarchical automata is already PSPACE-complete. This result shows a difference between membership and reachability, since reachability for communicating hierarchical automata is already EXPSPACE-complete [12].

The second problem considered in this paper is pattern matching for nMSCs. Given two nMSCs M, N, we want to know whether M occurs as a pattern of N. A polynomial time solution for this problem is not immediate. We apply some nice combinatorial techniques stemming from pattern matching on compressed texts and we obtain an algorithm of time $O(|\mathcal{C}_M|^2 \cdot |M|^2 \cdot |N|^2)$, where $|M|, |N|$ denote the sizes of the description of M and N, and $|\mathcal{C}_M|$ is the number of connected components in the communication graph of M. This question subsumes the test of equality of two nMSC, and shows that equality is decidable in PTIME as well.

Related work. Regarding the complexity of extended finite state machines, [12] considers the reachability and trace equivalence problems for communicating FSMs (Finite States Machines). Model-checking hierarchical FSMs against LTL and CTL properties is the topic of [4]. The paper [3] combines hierarchy and concurrency, analyzing the complexity of several problems (reachability, equivalence etc.) for communicating, hierarchical FSMs.

Several verification problems on MSCs and MSC-graphs have been considered recently, e.g. detecting races [2, 18], model-checking [5], pattern matching with gaps [19], inference [1] and realizability [17, 9, 14], model-checking against partial-order logics [16, 21]. Hierarchical MSCs have been also considered in [5] for the model-checking problem. We note however that our definition of nested HMSCs captures a larger class of MSC specifications than [5].

An extended abstract of this paper was presented at LATIN'02 [8]. As additional result we show here how to extend the polynomial time algorithm for pattern matching nMSCs to the case where the pattern is not connected.

2 Syntax and Semantics of Nested MSCs

We adopt the definition of (basic) message sequence charts (MSC for short), as described in the ITU-standard [13].

Definition 1. *(***Message Sequence Charts.***) A message sequence chart is a tuple $M = \langle P, E, \mathcal{C}, \ell, m, < \rangle$ where:*

- *P is a finite set of processes,*
- *E is a finite set of events,*
- *$\mathcal{C}$ is a finite set of names for messages and local actions,*
- *$\ell : E \to \mathcal{T} = \{i!j(a), i?j(a), i(a) \mid i \neq j \in P,\ a \in \mathcal{C}\}$ labels each event with its* type*: on process $i \in P$, the type is either a send $i!j(a)$ of message a to process j, or a receive $i?j(a)$ of message a from process j, or a local event $i(a)$. The labeling ℓ partitions the set of events by type (send, receive, or local), $E = S \dot\bigcup R \dot\bigcup L$, and by process, $E = \dot\bigcup_{i \in P} E_i$. We denote by $P(e)$ the process of event e (i.e., $P(e) = i$ iff $e \in E_i$).*
- *$m : S \to R$ is a bijection matching each send to the corresponding receive. If $m(s) = r$, then $\ell(s) = i!j(a)$ and $\ell(r) = j?i(a)$ for some processes $i, j \in P$ and some message name $a \in \mathcal{C}$. We denote the events s, r as matching events and the pair (s, r) as message.*
- *$< \subseteq E \times E$ is an acyclic relation between events consisting of:*
  - *a total order on E_i, for every process $i \in P$, and*
  - *$s < r$, whenever $m(s) = r$.*

The upper left part of Figure 1 depicts an MSC M on three processes with two messages and four events. Each vertical line corresponds to a process, with time increasing from top to bottom.

For the questions considered here, message names are irrelevant. Thus, send events will be of type $i!j$ and receive events of type $i?j$. Moreover, whenever we refer to an MSC in this paper, we mean actually its isomorphism class, where an isomorphism on the set of events E is a bijection that is compatible with the type function ℓ and the message function m.

For communication protocols it is natural to assume that each communication channel delivers messages first-in-first-out (*FIFO*). We assume the FIFO condition throughout the paper. That is, for all messages (e_k, f_k), $k = 1, 2$, such that $\ell(e_1) = \ell(e_2)$ and $\ell(f_1) = \ell(f_2)$ we require that $e_1 < e_2$ iff $f_1 < f_2$. The reflexive-transitive closure $\leq$ of the acyclic relation $<$ is a partial order called *visual order*. Every total order on E extending $\leq$ is then called *linearization* of M. A *configuration* (prefix) C of an MSC M is a downward closed subset of events, that is, if $e < f \in E$ with $f \in C$, then $e \in C$.

Note that with the FIFO message order, any total order on a set of events E defines at most one MSC. We obtain this MSC from the event sequence by matching the n-th send from i to j with the n-th receive on j from i, for each pair of distinct processes i, j.

A special case of the pattern matching problem considered in the paper is the equality test of two (nested) MSCs. In order to check the equality of two MSCs M, N (i.e., up to isomorphism) one can choose any linearization of M and check whether it is a linearization of N, too. An alternative approach, that will be used in our algorithms, is to check equality on each process. Thus, for an MSC $M = \langle P, E, \mathcal{C}, \ell, m, < \rangle$ and a process $i \in P$ we let $M|_i$ denote the projection of M on the set E_i of events located on i. We have $M = N$ if and only if M and N have the same set of processes, that is $P(M) = P(N) = P$, and if their projection on any process is equal, that is $M|_i = N|_i$ for each $i \in P$ (up to isomorphism). Note that both tests rely on the FIFO order of messages. Without the FIFO order, a linearization (or the projections on processes) does not suffice for recovering the MSC. For example, the linearization $s_1 s_2 r_1 r_2$ where s_1, s_2 are sends and $r_1 r_2$ are receives from process 1 to process 2, can produce two MSCs, one where $m(s_1) = r_1, m(s_2) = r_2$ and one where $m(s_1) = r_2, m(s_2) = r_1$.

We follow the ITU norm and define *nested MSCs* (nMSC for short) by allowing the reuse of an already defined MSC in a definition. The definition we give below aims at preserving the visual character of MSCs (see also Figure 1).

Definition 2. *(***Nested MSC, nMSC.***) A nested MSC $M = (M_q)_{q=1}^n$ is a finite sequence of macros of the form $M_q = \langle P_q, E_q, B_q, \varphi_q, \mathcal{C}, \ell_q, m_q, <_q \rangle$.*

Each macro M_q consists of:

- *A finite set E_q of events.*
- *A finite set P_q of processes.*
- *A finite set B_q of references (boxes) used by M_q.*
- *A function φ_q that associates each reference $b \in B_q$ with an index $q < \varphi_q(b) \leq n$. Thus, reference b refers to the macro $M_{\varphi_q(b)}$. We require that $P_{\varphi_q(b)} \subseteq P_q$.*
- *The type function $\ell_q : E_q \longrightarrow \mathcal{T}$, that associates each event with a type $i!j, i?j$ or $i(a)$, with $i, j \in P_q$, $i \neq j$ and $a \in \mathcal{C}$. The labeling ℓ partitions the set of events by type (send, receive, or local), $E_q = S_q \dot\bigcup R_q \dot\bigcup L_q$, and by process, $E_q = \dot\bigcup_{i \in P} E_{q,i}$. We denote by $P(e)$ the process of event e (i.e., $P(e) = i$ iff $e \in E_{q,i}$).*
- *The message function $m_q : S_q \longrightarrow R_q$ that maps each (send) event of type $i!j$ with a (receive) event of type $j?i$, for all $i \neq j$.*
- *The acyclic relation $<_q$ over the set of events and references $E_q \cup B_q$, defined by:*

- *For each process $k \in P_q$, the relation $<_q$ is a total order over the set $E_{q,k}$ of events located on k and the set of references $b \in B_q$ with $k \in P_{\varphi_q(b)}$.*
- *$e <_q f$ whenever $m_q(e) = f$ in M_q.*

The nesting depth of M is the maximal d such that there exists some sequence $q_1 < \cdots < q_{d+1}$ with $\varphi_{q_j}(b) = q_{j+1}$ for some $b \in B_{q_j}$, for all $1 \leq j \leq d$.

We define the indices such that the lowest levels of hierarchy, which stands for levels which do not use references to other level, corresponds to small indices.

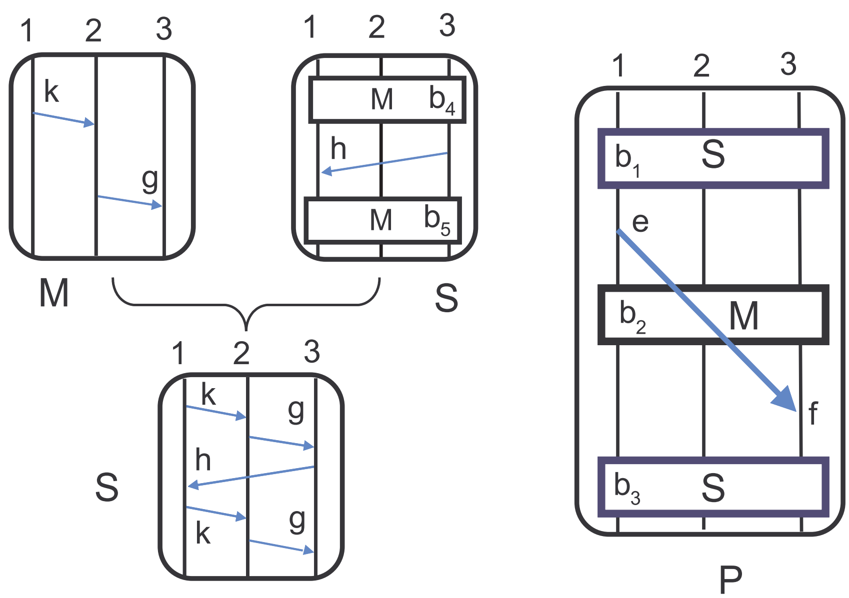


Fig. 1. An nMSC P using two references, S and M.

Example 1. Consider the nMSC P in Figure 1. It uses three references, $B_P = \{b_1, b_2, b_3\}$ that correspond to $\varphi_P(b_1) = \varphi_P(b_3) = S$ and $\varphi_P(b_2) = M$. The nesting depth of P is 2. The visual order $<_P$ of P requires on process 1 the order $b_1 <_P e <_P b_2 <_P b_3$. Notice that the definition of a nMSC forbids (f, e) to be a message, with f the send and e the receive, since this would contradict the acyclicity of $<_P$, even in the case where M would be empty.

The semantics of an nMSC is the MSC defined by replacing each reference of M by the corresponding MSC. Inductively it suffices to define the semantics of nMSCs of nesting depth one. Let $M = (M_q)_{q=1}^n$ be an nMSC of nesting depth one, with $M_q = \langle P_q, E_q, B_q, \varphi_q, \mathcal{C}, \ell_q, m_q, <_q \rangle$. For simplifying the notation below, we write instead of $\varphi_1(b)$ just b.

The MSC $\langle P, E, \mathcal{C}, \ell, m, < \rangle$ defined by $M = (M_q)_{q=1}^n$ is given by $P = P_1$, $E = \dot{\bigcup}_{b \in B_1} E_b \dot{\bigcup} E_1$, $\ell = \cup_{q=1}^n \ell_q$ and $m = \cup_{q=1}^n m_q$. The visual order $<$ is defined by $e < f$ if and only if either $m(e) = f$, or $P(e) = P(f)$ and one of the following conditions holds:

- $e, f \in E_1$ and $e <_1 f$,
- $e, f \in E_b$ and $e <_b f$,
- $e \in E_1, f \in E_b$ and $e <_1 b$,
- $e \in E_b, f \in E_1$ and $b <_1 f$,
- $e \in E_b, f \in E_{b'}$ and $b <_1 b'$,

where $b, b' \in B_1$. For simplicity, we denote the MSC defined by $M = (M_q)_{q=1}^n$ as M, too.

Example 2. For the nMSC P in Figure 1, the lower right part of the picture shows the MSC defined by S. Note that event $g \in E_M$ occurs twice in S – for simplicity, we denote both occurrences as g.

Note also that the semantics requires that $b_1 <_1 e$, but this does not mean that all events of $S = \varphi_P(b_1)$ must happen before $e \in E_P$. For instance, the first occurrence of g in S precedes event e of P, but the second occurrence is concurrent with e.

Obviously, a syntactically correct nMSC M might not yield an MSC because of the FIFO order. For example, the message (e, f) of P would violate the FIFO condition if M would contain a message from process 1 to process 3. Fortunately, it can be verified easily (polynomial time) whether an nMSC satisfies the FIFO condition. To test for the FIFO condition, it suffices to test that there is no $e < g < h < f$ and no $e < b < f$ with b containing a send from i to j, where $(e, f), (g, h)$ are two messages from i to j.

Size of nMSC. For complexity estimations we will denote by $\wp$ the overall number of processes. The size of an nMSC M is denoted as $|M|$. It represents the size of the syntactical description of M, where an event is of size one and the size of a reference is the number of its processes.

3 Nested High-Level MSC

An MSC can only describe a finite scenario. For specifying more complex behaviors, in particular infinite sets of scenarios, the ITU norm proposes to compose MSCs in form of MSC-graphs, by using choice and iteration.

Definition 3. *(***MSC-graph***) An MSC-graph is given as a tuple $G = \langle V, E, s, f, \varphi \rangle$, where:*

- *(V, E) is a directed graph with starting vertex $s \in V$ and final vertex $f \in V$.*
- *Each vertex v is labeled by the MSC $\varphi(v)$.*

In the same way as we defined nested MSCs from (flat) MSCs we can generalize MSC-graphs to *hierarchical HMSCs* (or *nested high-level MSCs*, nHMSC for short).

Definition 4. *(***Nested high-level MSC.***) An nHMSC is a finite sequence $G = (G_q)_{q=1}^n$, where each G_q is either a labeled graph or an nMSC. A labeled graph G_q is a tuple $\langle V_q, E_q, \varphi_q, s_q, f_q \rangle$ consisting of:*

- *A directed graph (V_q, E_q) with starting vertex s_q and final vertex f_q.*
- *A function φ_q that associates each vertex v with a reference $q < \varphi_q(v) \le n$, representing $G_{\varphi_q(v)}$.*

Thus, a node in an nHMSC can be mapped either to some graph or to an nMSC. This definition combines hierarchical automata as defined in [4] with our definition of nMSC. The special case where there is only one process (i.e., no concurrency) yields the hierarchical automata used in [4][1].

We first need to define the composition of two MSCs $N_1 N_2$ with $N_k = \langle P_k, E_k, \mathcal{C}_k, \ell_k, m_k, <_k \rangle$. Intuitively, we just glue together the two diagrams processwise. Let $N_1 N_2 = \langle P, E, \mathcal{C}, \ell, m, < \rangle$ with $E = E_1 \dot\bigcup E_2$, $P = P_1 \cup P_2$, $\mathcal{C} = \mathcal{C}_1 \cup \mathcal{C}_2$, $\ell = \ell_1 \cup \ell_2$, $m = m_1 \cup m_2$ and

$$< \; = \; <_1 \cup <_2 \cup \bigcup_{i \in P} E_{1,i} \times E_{2,i}.$$

The semantics of an nHMSC $G = (G_q)_{q=1}^n$ is a (possibly infinite) set of MSCs $L(G)$ defined recursively. If G_q is an nMSC, then $L(G_q)$ is a singleton consisting of the MSC defined by G_q. Let us consider a labeled graph G_q. Then $L(G_q)$ is the set of MSCs associated with the accepting paths of G_q, that is, paths starting in s_q and ending in f_q. With a path $v_1, \ldots, v_n$ in G_q we associate the set of all MSCs $M_1 \cdots M_n$, where $M_i \in L(G_{\varphi_q(v_i)})$ for all $1 \le i \le n$. The set of executions of G is defined as $L(G) = L(G_1)$.

As in [1] we also consider a weaker semantics for nHMSCs, that does not use the composition of MSCs (called *weak closure in [1]*). This semantics is based on taking the product of the sequential behaviors of single processes. Several algorithmic problems can be solved more efficiently for the weak closure of MSC-graphs. This makes it interesting to compare it with the usual semantics also in the setting of nHMSCs.

Weak closure of nHMSC. Let G be an nHMSC. Then $L^w(G)$ denotes the set of MSCs M such that for each process i there is some MSC $N \in L(G)$ such that $M|_i$ is equal to $N|_i$. Note that $L(G) \subseteq L^w(G)$ and that the inclusion is strict, in general (see [1]).

[1] Actually, [4] allows several final nodes in each automaton, which counts for the complexity of their algorithms.

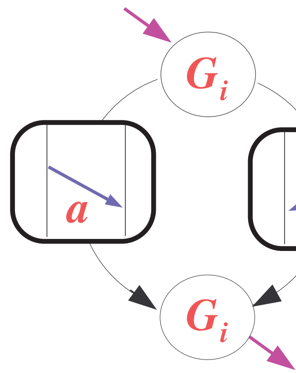


Fig. 2. An nHMSC G_{i+1} generating $(a+b)^{2^i-1}$ with $G_1 = \epsilon$.

4 Membership Problem

Checking the membership of an MSC M in an MSC-graph G is used typically for checking that no bad scenario can occur in a given specification. Another application is checking whether a good scenario is already covered by the specification. Checking membership is not an easy task already because of the concurrency implied by the MSC composition, all the more in the presence of hierarchy. The MSC membership problem $M \overset{?}{\in} L(G)$ with M an MSC and G an MSC-graph was considered in [1], together with the weak membership problem $M \overset{?}{\in} L^w(G)$. The results of [1] can be summarized as follows:

- The MSC membership problem is NP-complete. A deterministic algorithm of time $O(|G| \cdot |M|^{\wp})$ solves it[2], where $\wp$ is the number of processes.
- The *weak* MSC membership problem is solvable in time $O(|G| \cdot |M|)$.

So the MSC membership problem is solvable in polynomial time if we fix the number of processes.

4.1 Hierarchical Membership Problem

The membership problem seems *a priori* more difficult for an nMSC M against an nHMSC G, since the naive approach of guessing a path of G and checking equality with M is too expensive (both the path of G and the MSC defined by M can be of exponential size). However, it is easy to show that we can test membership in polynomial space:

Theorem 1 (Hierarchical MSC Membership Problem) *Given an nMSC M and an nHMSC G, we can decide whether $M \overset{?}{\in} L(G)$ in polynomial space.*

[2] This is a slightly improved runtime compared to the result stated in [1].

Proof. The idea of the algorithm is straightforward: we guess an MSC in $L(G)$ and we match it against the nMSC M, however expanding neither M nor G. Recall that for testing equality of two MSCs M, N, it suffices to choose one linearization of N and check whether it is a linearization of M. Hence, we can choose the linearization of the MSC in G (as long as we do not exclude any MSC in $L(G)$, that is as long as we do not exclude every linearization of one MSC). We consider only the linearizations in $Lin^0(G)$, where $Lin^0(G)$ is defined recursively. If G_q is an nMSC, then $Lin^0(G_q)$ is the set of linearization of G_q. With a path $v_1, \ldots, v_n$ in G_q we associate the set of all linearizations $u_1 \cdots u_n$, where $u_i \in Lin^0(G_{\varphi_q(v_i)})$ for all $1 \leq i \leq n$. Let us consider a labeled graph G_q. Then $Lin^0(G_q)$ is the set of linearizations associated with the accepting paths of G_q, that is, paths starting in s_q and ending in f_q. We define $Lin^0(G) = Lin^0(G_1)$. Intuitively, it means that we do not consider linearizations $uavbw$ of path $v_1 \cdots v_n$ where a belongs to a node v_i and b to v_j with $j < i$, that is every node needs to be fully executed before the next node can be considered.

We need to store a configuration of M, corresponding to the events already matched with the events from G. Since a configuration is a downward closed set of events, it can be stored as a tuple of $\wp$ events (remind that $\wp$ is the number of processes), representing the last event of the configuration on each process. Such a tuple is of linear size w.r.t. the size of M. Each event e of $M = (M_q)_{q=1}^n$ will be represented by a sequence $b_1, \ldots, b_m$ of references corresponding to the unfolding of references yielding e. That is, we inductively remind b_m for $e \in E_{\varphi(b_m)}$ where b_m is a reference of $\varphi(b_{m-1})$, plus the position of e in $M_{\varphi(b_m)}$. Thus, each event can be stored using linear-size memory. In our figure 1, the first occurrence of g in P corresponds to (b_1, b_4, g), the second occurrence to (b_1, b_5, g), and so on.

Similarly, we can store the current configuration of the linearization in $Lin^0(G)$ in space polynomial in $|G|$ (an event of G is represented by a sequence of nodes then of references). Since a new node is started only when the previous node is fully executed, the last event for every process belongs to the same node. The non-deterministic algorithm consists in guessing a successor configuration of G, obtained by extending the current configuration by an event e such that the new configuration is still a prefix of some linearization in $Lin^0(G)$. Then we check that e can extend the current linearization of M as well. The algorithm stops when the configuration that corresponds to the path being guessed in G is equal to M and the path of G is accepting.

$\square$

Theorem 2 below shows that PSPACE is the lowest complexity we can obtain for the hierarchical membership problem. The lower bound holds even if there is only one process (Theorem 2), or if the graph G is not hierarchical (Theorem 3), but not both (Theorem 4). This shows also that fixing the number of processes does not lower the complexity of the problem, unlike in the non hierarchical case.

We show the PSPACE lower bound for the following problem: given a straight-line program W (see below) and a hierarchical automaton $\mathcal{A}$, test whether $W \in L(\mathcal{A})$. This question corresponds to the hierarchical membership problem with a single process. Notice also that the weak membership problem $M \overset{?}{\in} L^w(G)$ [1] can be reduced to this question.

Straight-line programs. A straight-line program (SLP for short) over the alphabet Σ is a context-free grammar with variables $V = \{X_1, \ldots, X_k\}$, initial variable X_1 and rules from $V \times (V \cup \Sigma)^+$. The rules are such that there is exactly one rule for each left-hand side variable and if $X_i \longrightarrow \alpha$, then each X_j in α satisfies $j > i$.

The constraints on the rules make that any variable X_i generates a unique word. For convenience, we denote the word generated by the variable X_i also as X_i. The length of a variable X_i represents the length of the word generated by X_i and is denoted as $||X_i||$. Clearly, $||X_i||$ can be at most exponential in the number of rules. The size $|X_i|$ of an SLP is the sum of the sizes of the rules. Without loss of generality, we can assume that rules are of size 2, that is of the form $X \longrightarrow YZ$ with $Y, Z \in V \cup \Sigma$.

Since any MSC M is determined by its projections $(M|_i)_{i \in P}$, an nMSC M can be identified with $\wp$ SLPs $L^1, \ldots, L^\wp$. The SLP L^i generates the projection $M|_i$ of M on the set of events of process $i \in P$. We denote the variables used by L^i as $X|_i$, where $X \in \{M_q \mid q = 1 \cdots n\}$. The initial variable of each L^i is $M_n|_i$. Actually, the SLPs are not in Chomsky normal form to preserve this representation of nMSCs.

Example 3. For the nMSC P is Figure 1 we have the following SLP generating the projection on process 1: $P|_1 \to S|_1 e M|_1 S|_1$, $S|_1 \to M|_1 h M|_1$ and $M|_1 \to k$.

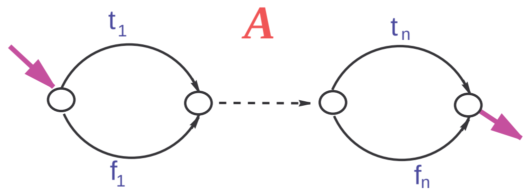
Theorem 2 *It is PSPACE-complete to check whether $W \in L(\mathcal{A})$ for some SLP W and hierarchical automaton $\mathcal{A}$. If the alphabet is unary, then the membership problem is NP-complete.*

Remark 1 The NP-hardness result in the unary case follows also from [23].

Proof. We first reduce (1-in-3) SAT to the unary membership problem, since we use this reduction in the general case, too. This problem is NP-complete, see [24, 6].

Let $\varphi = \wedge_{j=1}^m C(\alpha_j, \beta_j, \gamma_j)$ be an instance of (1-in-3) SAT over n variables $(x_i)_{i=1,n}$. Here, a clause $C(\alpha_j, \beta_j, \gamma_j)$ is true iff exactly one of the literals $\alpha_j, \beta_j, \gamma_j$ is true. We use the unary alphabet $\{a\}$. Note that any word $x \in a^*$ is uniquely defined by its length.

We associate with each clause $C_j = C(\alpha_j, \beta_j, \gamma_j)$ the word $w_j \in a^*$ of length 4^j. This word can be defined by an SLP of polynomial size. Let $W = w_1 \cdots w_m \in a^*$ be the word of length $\sum_{j=1}^m 4^j$. The automaton $\mathcal{A}$ consists of a sequence of choices with transitions labeled by t_i and f_i, for i varying from 1 to n, where $t_i = \sum_{j \in R_i} 4^j$ and $R_i = \{j \mid x_i \in \{\alpha_j, \beta_j, \gamma_j\}\}$. In the same way, $f_i = \sum_{j \in S_i} 4^j$ and $S_i = \{j \mid (\neg x_i) \in \{\alpha_j, \beta_j, \gamma_j\}\}$.



Any path ρ of $\mathcal{A}$ corresponds to a valuation σ where each variable x_i is true if the path chooses t_i, and false if it chooses f_i. Let n_j be the number of literals of C_j that are set true by σ. Recall that σ satisfies the formula φ iff $n_j = 1$ for all j. It is easy to see that ρ is labeled by the word $L \in a^*$ of length $\sum_{j=1}^{m} n_j 4^j$. Notice that since each clause has three literals, $n_j \in \{0, 1, 2, 3\}$ for all j. The length of L in base 4 is thus $(n_m n_{m-1} \ldots n_1 0)_4$. We have $W = L$ iff $(11 \ldots 10)_4 = (n_m n_{m-1} \ldots n_1 0)_4$, thus iff $n_j = 1$ for all j. That is, there is a path in $\mathcal{A}$ labeled by W iff there is a valuation satisfying φ, which implies that the membership problem for hierarchical automaton on a unary alphabet is NP-hard.

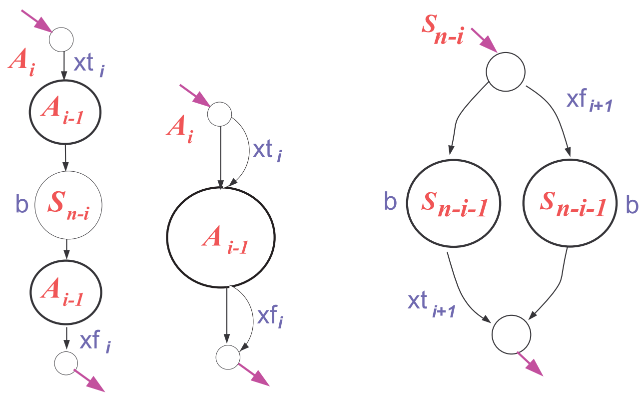
We now show the first statement of Theorem 2. We reduce the (1-in-3) QBF (one-in-three quantified boolean formula) to the hierarchical membership problem. Let φ be an instance of (1-in-3) QBF of the form $\varphi = Q_n x_n \cdots Q_1 x_1 \psi$, where $Q_i \in \{\exists, \forall\}$ and the formula ψ is of the form $\wedge_{j=1}^{m} C(\alpha_j, \beta_j, \gamma_j)$. As before, a clause $C(\alpha_j, \beta_j, \gamma_j)$ is true iff exactly one literal is true. The PSPACE-hardness of this problem is shown in [24, 6].

The idea is to make the valuations of the variables correspond to paths in the hierarchical automaton $(\mathcal{A}_i)_{i=0,n}$ and to validate the valuations using the SLPs $(W_i)_{i=0,n}$. We define the automata $\mathcal{A}_i$ and the SLPs W_i by induction on $i = 0, \ldots, n$. Here, we use the binary alphabet $\{a, b\}$. The letter a will have the same meaning as in the NP-case, and the letter b will be used as a delimiting symbol.

We define the words $w_j, t_i, f_i \in a^*$ with respect to ψ as before. That is, each w_j is associated with clause C_j and t_i, f_i are associated with variable x_i. Moreover, we associate with each variable x_i the word $w_{i+m} \in a^*$ of length 4^{i+m}. Let $W_0 = w_1 \cdots w_{n+m}$ be the word of a^* of length $\sum_{j=1}^{n+m} 4^j$, and let $\mathcal{A}_0$ be an automaton consisting of one ϵ-transition from its initial state to its final state. Let also S_0 be an automaton consisting of one transition labeled by b. The SLP-compressed words $(W_i)_{i=1,n}$, are defined by:

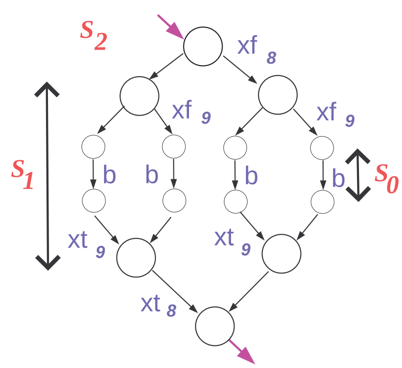
- $W_i \longrightarrow W_{i-1}$, if $Q_i = \exists$,
- $W_i \longrightarrow W_{i-1}\, b\, W_{i-1}$, if $Q_i = \forall$.

The recursive definition of the automata $(\mathcal{A}_i)_{i=1,n}$ and $(S_i)_{i=0,n-1}$ is illustrated in the figure below. Transitions are either labeled by ϵ, or by $\mathrm{xt}_i = t_i\, w_{i+m}$ or $\mathrm{xf}_i = f_i\, w_{i+m}$. The automaton on the left defines $\mathcal{A}_i$ when $Q_i = \forall$, the automaton in the middle defines $\mathcal{A}_i$ when $Q_i = \exists$, and the automaton on the right defines S_i. Note that the symbol b is only generated by S_0. In the figure we recall its position by marking a b aside each S_i.



The overall idea is as follows. The values of $x_{i+1}, \ldots, x_n$ are already chosen when an automaton calls $\mathcal{A}_i$ (from a higher hierarchy level). The automaton $\mathcal{A}_i$ on the left sets x_i true, then uses S_{n-i} to recover the fixed values of $x_{i+1}, \ldots x_n$, and finally it sets x_i false. The automaton $\mathcal{A}_i$ in the middle guesses whether x_i is true (by taking the transition labeled by xt_i) or false (by choosing the transition labeled by xf_i). If it chooses both transitions labeled by $\mathrm{xt}_i, \mathrm{xf}_i$ or none of them, then the word labeling this path will not be equal to W_n because W_n contains exactly one occurrence of w_{i+m} between any two consecutive b's. We illustrate how $\mathcal{A}_i$ works on figure 3, that shows the unfolding of the automaton $\mathcal{A}_2$ for $\varphi = \forall x_2 \forall x_1 \psi$ on the left and for $\varphi = \exists x_2 \forall x_1 \psi$ on the right.

To illustrate how S_{n-i} recovers the values of $x_{i+1}, \ldots, x_n$, we show S_{n-i} for $n = 9, i = 7$ in the figure below.



$\mathcal{A}_i$ and S_i are designed so that any path of $\mathcal{A}_i$ is labeled by at most one xt_i and at most one xf_i between any two consecutive b's, for each i (for convenience, we suppose that each automaton starts and ends with a fictive b transition). That is, a path can be labeled by xt_i and xf_i, but not by two xf_i or two xt_i. By contradiction, assume that there are two consecutive b's in $\mathcal{A}_i$ such that there is a path from one to another labeled by two xt_j (the case xf_j is symmetric). We take the minimal $\mathcal{A}_i$ which ensures this. By the minimality of $\mathcal{A}_i$, this can only happen either because of the first xt_j transition of $\mathcal{A}_i$, or between S_{n-i} and one of the two $\mathcal{A}_{i-1}$. Since in S_{n-i} all xt_k occur after the (unique) b, there is no xt_j

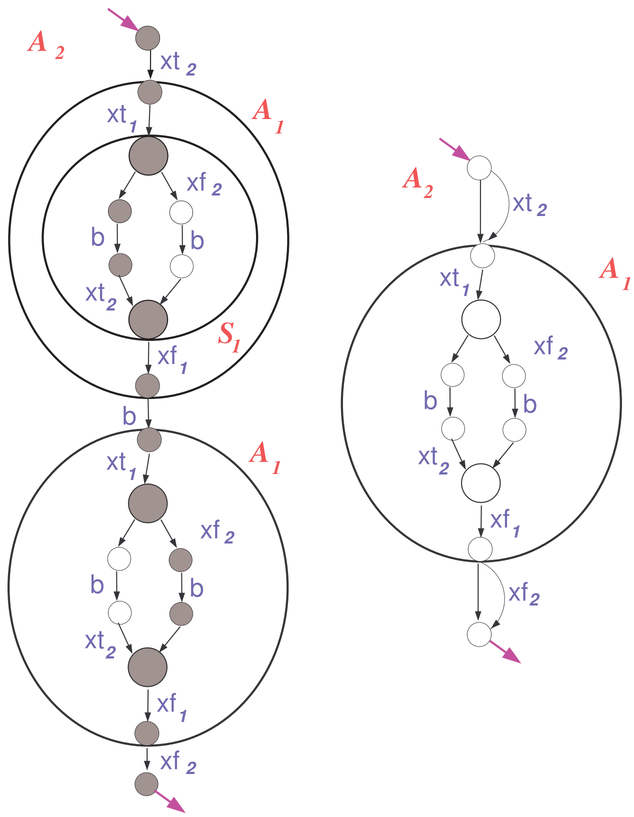


Fig. 3. Unfolding of A_2 for $Q_2x_2Q_1x_1 = \forall x_2 \forall x_1$ on the left, and on the right, unfolding of A_2 for $Q_2x_2Q_1x_1 = \exists x_2 \forall x_1$

in $\mathcal{A}_{i-1}$ before its first b (if any). It already shows a contradiction in the case where $Q_i = \exists$. Consider now the case $Q_i = \forall$. For the same reason as before, there can be at most one xt_j between the last b of $\mathcal{A}_{i-1}$ and the b in S_{n-i}, for all $k < i$. Finally, between the b of S_{n-i} and the first b of the second $\mathcal{A}_{i-1}$ there can be at most one xt_k with $k > i$ (from S_{n-i}) and at most one xt_k with $k < i$ (from $\mathcal{A}_{i-1}$). Thus, in all cases we contradict the assumption on $\mathcal{A}_i$.

Let us prove that $W_n \in L(\mathcal{A}_n)$ iff there exists a satisfying valuation tree VT for φ. A valuation tree VT is a binary tree of height $n+1$ such that its root (level n) is labeled by x_n and all nodes on level l are labeled by x_l. The leaves are on level 0, and are unlabeled. A node v labeled by x_i corresponds to a valuation $\sigma(v)$ of the variables $x_{i+1}, \ldots, x_n$. For instance, if the valuation for a node is x_n is true, then its children must evaluate x_n to true, and evaluate x_{n-1} either to true or false. Moreover, a node on level k have two children if x_k is universally quantified (one child evaluate x_k to true and the other one to false), and one child if x_k is existentially quantified. We say that a valuation tree satisfies a QBF formula $\varphi = Q_n x_n \cdots Q_1 x_1 \psi$ if for every valuation of every leaf, ψ is true.

Using the property we just showed, we can note that between any two consecutive b's of any path of $\mathcal{A}_n$, there are at most three w_j and two w_{i+m} for any $1 \leq j \leq m, 1 \leq i \leq n$. Thus our coding in base four for determining whether a clause is true, is still applicable. Hence, a path ρ of $\mathcal{A}_n$ is labeled by W_n iff for all $1 \leq k \leq n+m$ there is exactly one w_k between any two consecutive b's.

Assume that VT is a valuation tree showing that φ is true. A valuation $\sigma(v)$ defines two words $T(v), F(v)$ as follows: the word $T(v)$ is the concatenation of all xt_j where $j > i$ and x_j is true in $\sigma(v)$. The word $F(v)$ is the concatenation of all xf_j where $j > i$ and x_j is false in $\sigma(v)$. Let v be a node of VT labeled by x_i. We define the word $\rho(v) = T^{-1}(v) W_i F^{-1}(v)$. We recall that $T(v), F(v)$ are words over a^*, hence $T^{-1}(v) W_i F^{-1}(v)$ is the word that results from W_i by deleting $|T(v)|$ many a's in the prefix and by deleting $|F(v)|$ many a's in the suffix.

Let us show by induction on level i that $\rho(v)$ is in $L(\mathcal{A}_i)$ for any node v of VT on level i.

If v is a leaf of VT, then it defines an accepting valuation for ψ, hence $T(v)F(v) = W_0$ using the same argument as in the NP-hardness case. Hence $\rho(v) = W_0 W_0^{-1} = \epsilon \in L(\mathcal{A}_0)$.

Consider an internal node v labeled by x_i with $Q_i = \forall$. Let v_1, v_2 be the children of v, with v_1 corresponding to x_i true, and v_2 to x_i false. By induction let us suppose that $\rho(v_1)$, $\rho(v_2)$ are in $L(\mathcal{A}_{i-1})$. Then,

$$\begin{aligned}
\rho(v) = T^{-1}(v) W_i F^{-1}(v) &= T^{-1}(v) W_{i-1} b W_{i-1} F^{-1}(v) \\
&= T^{-1}(v) T(v_1) \rho(v_1) F(v_1) b T(v_2) \rho(v_2) F(v_2) F^{-1}(v) \\
&= \mathrm{xt}_i \rho(v_1) F(v_1) b T(v_2) \rho(v_2) \mathrm{xf}_i
\end{aligned}$$

We used in the equations above $T^{-1}(v) T(v_1) = \mathrm{xt}_i$ for the positive child v_1 of v and $F^{-1}(v) F(v_2) = \mathrm{xf}_i$ for the negative child v_2 of v. Moreover, $F(v_1) b T(v_2) =$

$F(v)bT(v) \in L(S_{n-i})$ since the indices of false variables in $\sigma(v_1)$ and of true variables in $\sigma(v_2)$ form a partition of $\{i+1, \ldots, n\}$. This shows that $\rho(v) \in L(\mathcal{A}_i)$.

Consider an internal node v that is labeled by x_i with $Q_i = \exists$. Assume by symmetry that v_1 is the child of v in VT (thus, x_i is true). By induction we assume that $\rho(v_1)$ is in $L(\mathcal{A}_{i-1})$. It is easy to show now that $\rho(v) \in L(\mathcal{A}_i)$ using:

$$\begin{aligned}\rho(v) = T^{-1}(v)W_iF^{-1}(v) &= T^{-1}(v)W_{i-1}F^{-1}(v)\\ &= T^{-1}(v)T(v_1)\rho(v_1)F(v_1)F^{-1}(v)\\ &= \mathrm{xt}_i\rho(v_1)\end{aligned}$$

For the reverse direction the arguments are similar. From a word $W = W_n$ of $\mathcal{A} = \mathcal{A}_n$, we obtain subwords $\rho(v)$ in $L(\mathcal{A}_i)$ as above, labeled by $T^{-1}(v)W_iF^{-1}(v)$. For each leaf v this means that $\sigma(v)$ satisfies exactly one literal per clause.

$\square$

Theorem 2 shows immediately that the hierarchical membership problem is PSPACE-hard even with one process, by encoding the alphabet $\{a, b\}$ by local actions on a single process. Similar arguments can be used for the case where G is an MSC-graph with no hierarchy, as shown in the following theorem.

Theorem 3 *The hierarchical MSC membership problem $M \overset{?}{\in} L(G)$ is PSPACE-complete. The lower bound holds even if G is an MSC-graph, or if there is only one process.*

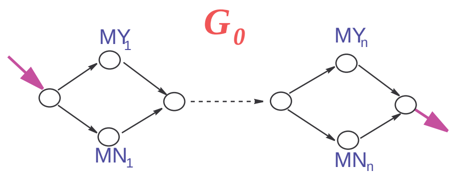
Proof. The problem we reduce from is (1-in-3)QBF. Let F be an instance of (1-in-3)QBF of the form $F = (Q_nx_n)\ldots(Q_1x_1)\varphi$, where $Q_i \in \{\exists, \forall\}$ and the formula φ is of the form $\wedge_{j=1\ldots m}R(\alpha_{j,1}, \alpha_{j,2}, \alpha_{j,3})$, with $\alpha_{j,k}$ literals.

The idea is to let valuations of the variables to correspond to paths of G and to validate the valuations using the nMSC M. We define the graph G and the nMSC M by induction on $F = F_n$. Let $F_i = (Q_ix_i)\,F_{i-1}$, with $F_0 = \varphi$. Each F_i will determine G_i, M_i.

The processes used in the construction are $\mathrm{SC}_1, \ldots, \mathrm{SC}_m$ and $\mathrm{RC}_1, \ldots, \mathrm{RC}_m$, plus $\mathrm{VY}_1, \ldots, \mathrm{VN}_n$ and $\mathrm{VN}_1, \ldots, \mathrm{VN}_n$. Here V means a variable and C a clause, S stands for "send", R for "receive", Y for "yes" and N for "no".

For all i, let MY_i be the MSC consisting of a message from VY_i to VN_i, then back from VN_i to VY_i, and a message from SC_j to RC_j for all j such that $x_i \in \{\alpha_{j,1}, \alpha_{j,2}, \alpha_{j,3}\}$. Symmetrically, let MN_i be the MSC consisting of a message from VN_i to VY_i, then back from VY_i to VN_i, and a message from SC_j to RC_j for all j such that $\neg x_i \in \{\alpha_{j,1}, \alpha_{j,2}, \alpha_{j,3}\}$.

M_0 is an MSC consisting of one message from SC_j to RC_j, for all j. The MSC-graph G_0 consists of $4n$ vertices, labeled by MY_i, MN_i, or $\emptyset$. The graph chooses between MY_i and MN_i for all i, as depicted below:

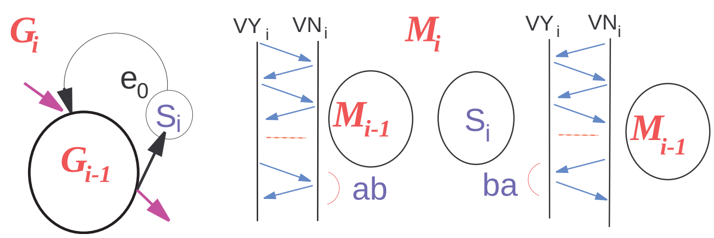


Note that all messages defined above commute, except for the ones between VY_i and VN_i. Let a_i be the message from VY_i to VN_i, and b_i the message from VN_i to VY_i. We will use the order between a_i, b_i as follows: The sequence $a_i b_i$ means that x_i is true, while $b_i a_i$ means that x_i is false.

Assume now that G_{i-1}, M_{i-1} are already defined, and that there are f universal quantifiers in F_{i-1}. For simplicity, we denote $a = a_i$ and $b = b_i$. Note that in a valuation tree for F showing that F is true, each value 0 or 1 assigned to the variable x_i is used by 2^f leaves. A valuation tree is defined as usual, by assigning each universally quantified variable two children labeled 0 and 1, respectively each existentially quantified variable one child labeled 0 or 1.

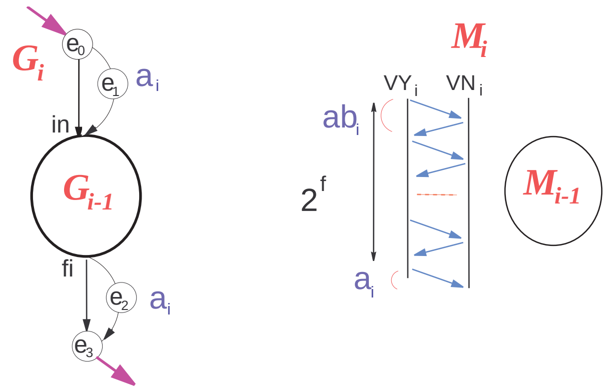
If $F_i = \forall x_i F_{i-1}$, then let $M_i = (ab)^{2^f} M_{i-1} S_i (ba)^{2^f} M_{i-1}$ (see Figure 4.1). The MSC S_i is used for synchronizing processes occurring in M_i. It contains a message between each (ordered) pair of processes of M_i (in some arbitrary order). Note that using the hierarchy we can describe $(ab)^{2^f}$, and thus M_i, by an expression of polynomial size.

Let $G_i = (V_i, E_i)$, where $V_i = V_{i-1} \cup \{e_0\}$ and $E_i = E_{i-1} \cup \{(\mathrm{Fin}, e_0), (e_0, \mathrm{In})\}$. The initial node In (the final node Fin, respectively) of G_i is the same as for G_{i-1}. The vertex e_0 is labeled by the synchronization MSC S_i.



The definition of M_i, G_i can be explained intuitively as follows. Let ρ be a path of G_i labeled by M_i. Note that the MSC S_i occurring in M_i has to match the MSC S_i of e_0. Thus $\rho = \rho_1 e_0 \rho_2$, with ρ_1 an accepting path of G_{i-1} labeled by $(ab)^{2^f} M_{i-1}$ and ρ_2 an accepting path of G_{i-1} labeled by $(ba)^{2^f} M_{i-1}$. Each time ρ_j goes through G_0 (which happens 2^f times), ρ_j consumes either ab of MY_i or ba of MN_i, so ρ_j consumes all occurrences of a, b in $(ab)^{2^f}$. In particular, all occurrences consumed by ρ_1 are of the form ab, which ensures that the valuation of x_i associated with ρ_1 is consistent (x_i is true). The same holds for the path ρ_2, where the value of x_i is forced to be false.

Suppose now that $F_i = \exists x_i F_{i-1}$. Let $M_i = (ab)^{2^f}(a)M_{i-1}$, and $G_i = (V_i, E_i)$, where $V_i = V_{i-1} \cup \{e_0, e_1, e_2, e_3\}$. Let $E_i = E_{i-1} \cup \{(e_0, \mathrm{In}), (\mathrm{Fin}, e_3), (e_0, e_1), (e_1, \mathrm{In}), (\mathrm{Fin}, e_2), (e_2, e_3)\}$, where as above In is the initial vertex and Fin is the final vertex of G_{i-1}. The initial and final vertices of G_i are e_0 et e_3. We label e_1 and e_2 with a, and e_0 et e_3 with $\emptyset$.



The underlying idea in this case is that the additional occurrence of a in M_i must be matched by e_1 or e_2 (nowhere else there is an a). If it is e_1, every time the path ρ goes through G_0, it must choose ba, hence it goes through VN_i. The corresponding value for x_i is then forced to be false. If it is e_2, then ρ must choose ab, hence it goes through VY_i. The rest of the proof is easy, see the proof of theorem 2.

□

However, if there is only one process and hierarchy is not allowed for the graph G (or the MSC/word M), then our lower bound proof does not work anymore. Indeed, we show below that in the case where the word W or the automaton $\mathcal{A}$ are flat, the membership problem is solvable in polynomial time.

Theorem 4 *1. Let W be a word defined by an SLP and let $\mathcal{A}$ be an NFA. Deciding whether $W \in L(\mathcal{A})$ can be achieved in time $O(|W| \cdot |\mathcal{A}|^3)$.*

2. Let W be a word and let $\mathcal{A}$ be a hierarchical automaton (hNFA for short). Deciding whether $W \in L(\mathcal{A})$ can be achieved in time $O(|W|^3 \cdot |\mathcal{A}|^3)$.

For the first statement in the theorem above a similar result (for Lempel-Ziv compressed words and regular expressions) has been shown in [23].

The polynomial time algorithms for Theorem 4 are stated below. The first algorithm computes in a dynamic programming way the set T_X of pairs (a, b) of states of a NFA $\mathcal{A}$ between which a path labeled by X exists, for each variable X of the SLP. A variable X is on the lowest level, if the rule associated with X is terminal.

Membership $((X_i)_{i=1,n}$ SLP-compressed word, A=(V,E,a_0,a_f) NFA)
For each variable X_i in the lowest level:
    T_{X_i} = { (a,b) $\in$ V $\times$ V | a $\xrightarrow{X_i}$ b };

```
For i = 1 ··· n:
    Let T_{X_i} = ∅;
    Let Y,Z s.t. X_i → Y Z;
    For all vertices a,b,c ∈ V:
        If (a,b) ∈ T_Y and (b,c) ∈ T_Z:
            T_{X_i} = T_{X_i} ∪ {(a,c)};
Return (a_0,a_f) ∈ T_{X_1};
```

The second algorithm computes for each sub-automaton B of a hNFA $\mathcal{A}$ the set T_B of factors $W_i \cdots W_j$ of a word W that it accepts. This algorithm uses another dynamic programming algorithm calculating for each factor $W_i \cdots W_j$ the set $T_{i,j}$ of pairs (a, b) of states of B between which a path labeled by $W_i \cdots W_j$ exists. For convenience, we assume without loss of generality that every transition is a subautomaton, but those on the lowest level of the hierarchy (else, we just replace every transition labeled by a by a subautomaton in the lowest level of the hierarchy with a unique transition from the initial to the final state labeled by a.). We use the fact that $(a, b) \in T_{i,j}$ if either there is a transition from a to b labeled by a subautomaton C accepting $W_i \cdots W_j$, or else the path labeled by $W_i \cdots W_j$ can be decomposed as a, c and c, b, and then there exists $0 < e < j-i$ such that $(a, c) \in T_{i,i+e}$ and $(c, b) \in T_{i+e,j}$. We thus compute first the lower levels of hierarchy, and we compute first the sets $T_{k,k+d}$ for small d, which allows us not to use a costly fix point algorithm.

Membership (W word, A=(V,E,a_0,a_f) hNFA)

```
For each sub-automaton B of A in the lowest level of hierarchy:
   T_B = {(i,j) | W_i ··· W_j is accepted by B};
For each sub-automaton B of A, by increasing hierarchical level:
   For d = 0, ..., |W|, for i = 1, ..., |W| - d,
       D_{i,i+d} = { (a,b) | ∃ subaut. C s.t. a --C--> b
                      and (i,i+d) ∈ T_C};
       For each e < d and every a,b,c vertices of B,
          If (a,b) ∈ D_{i,i+e} and (b,c) ∈ D_{i+e,i+d}:
             D_{i,i+d} = D_{i,i+d} ∪ {(a,c)};
   T_B= {(i,j) | (a_0,a_f) ∈ D_{i,j}};
Return (1,|W|) ∈ T_A
```

The figure below summarizes the complexities of the different variants for the hierarchical MSC membership problem as considered in this section. The last two columns correspond to the case of a single process (word case) and to the general MSC case, respectively. The fact that the membership problem is NP-complete for an MSC M and an nHMSC G is easy to show since it is already NP-hard for H an HMSC [1], and it suffices to guess a path of G of the size of M, which is polynomial, and check whether it is labeled by M.

M	G	words	MSC
Flat	Nested	P	NP-complete
Nested	Flat	P	PSPACE-complete
Nested	Nested	PSPACE-complete	PSPACE-complete

Fig. 4. Complexity of membership problems.

5 Pattern Matching of nMSCs

The aim of this section is to show that pattern matching on nMSCs can be achieved in polynomial time, i.e., without unfolding the nMSCs. We first consider a special case of pattern matching, namely testing equality of nMSCs. Then we describe first a pattern matching algorithm when the pattern nMSC is connected, and second the additional work to do when the pattern is not connected.

5.1 Equality of nMSCs

Recall first that the FIFO message order allows testing the equality of two MSCs M and N process-wise, which amounts to testing the equality of $\wp$ pairs of words (over the type alphabet $\mathcal{T}$). In the hierarchical case we already used in Section 4.1 the representation of an nMSC M by $\wp$ straight-line programs L^i, where the SLP L^i generates the projection $M|_i$ of M on process i.

Thus, for testing the equality of two nMSCs in polynomial time, we can use directly the following result:

Theorem 5 ([22]) *Let P be an SLP, and A, B be two variables of P. We can determine whether A and B generate the same word in time $O(|P|^5 \log(|P|))$.*

The theorem above provides an algorithm for testing $M = N$ of time $O((|M|+|N|)^5 \log(|M| + |N|))$. We can improve the running time by using the pattern matching algorithm described in the next section.

5.2 Pattern Matching nMSCs

Definition 5. *The pattern matching problem for two MSCs M and $N = \langle P, E, \mathcal{C}, \ell, m, < \rangle$ consists in knowing whether there exists some subset $F \subseteq E$ of events of N such that the restriction of the mappings ℓ, m to F equals M. Moreover, we require that F is convex, that is if $e, f \in F$ and $e < g < f$, then $g \in F$. In particular, the message mapping m must be one-to-one between the send and receive events in F. We call such an event set F an occurrence of M in N. If M, N are nMSCs, then M occurs as a pattern in N if the MSC defined by M is a pattern in the MSC defined by N, and we write $M \subseteq N$ in this case.*

It is easy to see that for an MSC M to be a pattern of an MSC N it does not suffice to have each $M|_i$ a pattern of $N|_i$. Of course, this condition is necessary. Before to consider the hierarchical case, we show a simple algorithm for the non-hierarchical case:

Theorem 6 *Let M, N be two MSCs. We can check whether M is a pattern of N in linear time.*

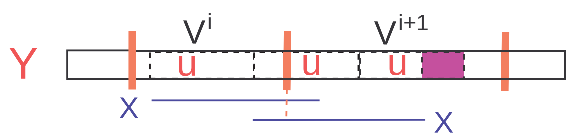
Proof. The main idea comes from pattern matching in trace monoids, [15].We need the linear time algorithm of Knuth-Morris-Pratt for determining occurrence of $M|_i$ in $N|_i$, for all $i \in P$. We search for tuples of occurrences of $(M|_i)_{i \in P}$ that form a factor of N. Thus, we look for a configuration of N such that on each process i, we have $M|_i$ as a suffix. This is done by recording the set J of processes i satisfying this condition and progressing one event at a time on processes $j \notin J$. If this is not possible, the next event on every $i \notin J$ is a receive from some $j \in J$, while the corresponding send from j to i in N has not been seen yet. We then progress on j, and update J by using Knuth-Morris-Pratt algorithm to know whether $j \in J$ or $j \notin J$. The overall complexity of the algorithm is linear, by taking care that each event in N is considered at most once. □

Definition 6. *Let $N = (N_i)_{i=1}^n$ be an nMSC (or an SLP), and $i, j \leq n$.*

1. *We write $N_i < N_j$ whenever N_i is used in the definition of N_j or in the definition of Z with $Z < N_j$. We write $N_i \leq N_j$ when $i = j$ or $N_i < N_j$.*
2. *We say that N_i occurs* literally *in N_j when N_i is used as a reference (variable resp.) in the definition of N_j, and we write $N_i \in N_j$ if it is the case.*

The strategy we will use for nMSC pattern matching is to compute an implicit representation of all positions where $M|_i$ occurs as a pattern in $N|_i$. In a second step we compute all positions where the projections $M|_i$ form a factor M. The basis of our algorithm is a pattern matching algorithm for SLP-compressed words, that was proposed in [20] (see also [22]):

Theorem 7 ([20]) *Let P be an SLP and let A, B be two variables of P. One can determine all occurrences of the word defined by A in the word defined by B in time $O(|A|^2|B|^2)$.*



arithmetic progression $Occ(X, Y, V^i)$

The idea of the algorithm in [20] is based on word combinatorics. Let X be a variable of A and suppose that X occurs in B, i.e. (the word defined by) X is a factor of (the word defined by) B. Suppose that X does not appear as a factor inside any variable Y of B with rule $Y \to \alpha \in \Sigma^*$. Then X occurs in a variable Y with $Y \longrightarrow V^1 \cdots V^k$. Let i be such that V^i is the first symbol (variable or letter) that this occurrence of X overlaps, and the occurrence ends beyond V^i (see also figure above). In particular, Y is the lowest variable that contains this occurrence of X. We let $\mathrm{Occ}(X, Y, V^i)$ denote the set of positions of Y at which an occurrence of X starts within V^i and ends beyond V^i, or starts and ends within V^i if V^i is in the lowest level of the hierarchy. Let $\mathrm{Occ}(X, Y) = \bigcup_i \mathrm{Occ}(X, Y, V^i)$.

Using a combinatorial argument (lemma of Fine and Wilf, [7]), it is shown in [20] that $\mathrm{Occ}(X, Y, V^i)$ is an arithmetic progression that can be computed by dynamic programming in polynomial time. Therefore, $\mathrm{Occ}(X, Y)$ consists of at most $|Y|$ arithmetic progressions, precisely at most one for each i (and $|V^i|$ for V^i on the lowest level of the hierarchy). That is, we can represent $\mathrm{Occ}(X, Y, V^i)$ by a triple of numbers (n, s, k) where n and $n + s$ are the positions in Y of the two first occurrences of X in $\mathrm{Occ}(X, Y, V^i)$, and $k = \#\mathrm{Occ}(X, Y, V^i)$ is the number of occurrences of X in $\mathrm{Occ}(X, Y, V^i)$. That is, we have $Y = Y_1 X Y_2$ with $||Y_1|| = n + is$, for all $0 \leq i < k$. As an example, consider the words $Y = aaabababababb$ and $X = ababab$. The arithmetic progression which corresponds to the occurrences of X in Y is $(2, 2, 3)$ (the first position in a word being 0).

Remark 2 Using the algorithm of [20] we immediately obtain that the equality of two SLPs M, N can be checked in time $O(|M|^2|N|^2)$, which improves the complexity provided by the algorithm proposed by Plandowski in [22].

Throughout the section we denote occurrences of projections $M|_i$ using superscripts. That is, $M|_i^1$ will correspond to a given starting position of $M|_i$ as pattern of $N|_i$. Suppose that M^i is a factor of $N|_i$ for all $i \in P$. We say that $(M^i)_{i=1}^{\wp}$ forms a factor of N if there exists M a factor of N such that $M|_i = M^i$ for all $i \in P$.

5.3 Pattern Matching for Connected Patterns

We turn now to the pattern matching problem for nMSCs M, N where the pattern M is connected. That is, we suppose in this section that M cannot be written as $M_1 M_2$, where M_1, M_2 are non-empty MSCs with no common process.

Definition 7. *Let $M|_i^1$ and $M|_j^2$ be occurrences of $M|_i$ in $N|_i$, resp. of $M|_j$ in $N|_j$. We say that $M|_i^1$ and $M|_j^2$ are compatible, if the first send (resp. receive) between the processes i and j on $M|_i^1$ matches the first receive (resp. send) on $M|_j^2$ (if i, j communicate in M). More generally, we call the indices corresponding to $M|_i^1$, $M|_j^2$ in a given arithmetic progression compatible.*

Lemma 1. *Let $(M|_i^0)_{i=1}^{\wp}$ be occurrences of $M|_i$ in $N|_i$. Then $(M|_i^0)_{i=1}^{\wp}$ forms a factor of N iff $(M|_i^0)_{i=1}^{\wp}$ are pairwise compatible.*

As in the previous section we will denote by $\mathrm{Occ}(M, Y)$ the set of occurrences M^0 of M in Y such that M^0 does not occur in any $Z < Y$. We denote by $\mathrm{Occ}(M, Y, V) \subseteq \mathrm{Occ}(M, Y)$ those occurrences that start within V (ending beyond V), where $V \in Y$ is a reference occurring literally in Y. It means that one event of the occurrence has to occur in V and one (not necessarily on the same process) has to occur not in V. Obviously, no event may occur before V.

Our search for compatible occurrences uses the following properties, that are easily shown using the fact that M is connected:

Fact 1 *1. Let Y be a variable of N and $h \neq j$ two processes. Then for each $M|_h^0 \in Occ(M|_h, Y)$ there can be at most one occurrence $M|_j^0$* **in** *Y that is compatible with $M|_h^0$.*

2. For each occurrence M^0 in $Occ(M, Y, V)$ there exists some process h such that $M^0|_h \in Occ(M|_h, Y, V)$. We call such a process h a leading process for M^0. Thus, any pairwise compatible tuple $(M^0|_k)_{k \neq h} \subseteq Y$ is determined by the occurrence $M^0|_h$, because of Fact 1.1.

Example 4. For the nMSC P in Figure 1 and the pattern N in Figure 5 we have $\mathrm{Occ}(N, P) = \emptyset$ and $\mathrm{Occ}(N, S)$ is a singleton, corresponding to the unique occurrence of N in S. The leading processes are 1 and 3, since e.g. $\mathrm{Occ}(N|_3, S|_3) = \{0\}$. Note that $\mathrm{Occ}(N|_2, S|_2) = \emptyset$ and $\mathrm{Occ}(N|_2, M|_2) = \{0\}$ is the arithmetic progression $(0, 0, 0)$.

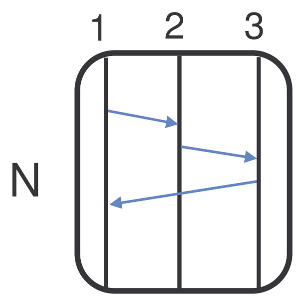


Fig. 5. Pattern MSC N

An **index** $i = n + js$, $j < k$, of an arithmetic progression (n, s, k) in Y is called *external*, if it is either the first or the last index of the progression, that is $i = n$ $(j = 0)$ or $i = n + (k-1)s$ $(j = k-1)$. Any non external index is called an *internal index*.

The next proposition provides the main argument that the search for a pairwise compatible tuple of occurrences $(M|_i)_{i \in P}$ can be done in polynomial time. Intuitively, we must show that the occurrences of $(M|_i)_{i \in P}$ can be located in the same variable Y of N, up to polynomially many exceptions. Without this property we would have to consider different variables Y^i for different processes $i \in P$. We recall that for every message (e, f) in an nMSC $N = (N_q)_{q=1}^n$ the events e and f appear literally in the same macro N_q.

Proposition 1 *Assume that $M^0 \in Occ(M, Y, V)$ with $M^0|_i \in Occ(M|_i, Y^i, V^i)$, where Y, Y^i, V^i are variables of N. Then we have one of the following two cases:*

1. *$Y^i = Y$ and $V^i = V$ for all i.*

2. For some leading process h for M^0 (i.e., $V^h = V$ and $Y^h = Y$), the occurrence $M^0|_h$ is an external index of $Occ(M|_h, Y^h, V^h)$.

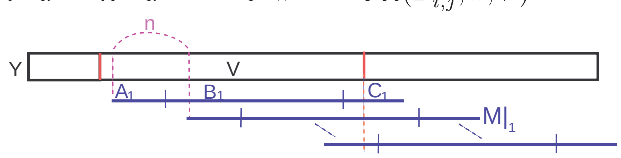
Proof. Suppose that there is no leading process h such that $M^0|_h$ is an external index of $\mathrm{Occ}(M|_h, Y^h, V^h)$. Assume also that there is a message from process i to process j in M. We decompose $M|_i = A_{i,j} B^s_{i,j} C_{i,j}$ such that $B^s_{i,j}$ begins with the first send from i to j, and ends with the last one. Similarly, we decompose $M|_j = A_{j,i} B^r_{j,i} C_{j,i}$ such that $B^r_{j,i}$ begins with the first receive on j from i, and ends with the last one. We need the next lemma to infer that if an occurrence M^0 is such that $M^0|_i \in \mathrm{Occ}(M|_i, Y^i, V^i)$ and $M^0|_j \in \mathrm{Occ}(M|_j, Y^j, V^j)$ are both internal indices, then we have $Y^i = Y^j$ and $V^i = V^j$. This will allow finishing the proof of the proposition.

Lemma 2. *Let $\pi = Occ(M|_i, Y, V)$ be an arithmetic progression consisting of at least three indices. Then each $B^s_{i,j}$ associated with some internal index of π belongs to $Occ(B^s_{i,j}, Y, V)$.*

Proof of lemma: Since $M|_i$ belongs to an arithmetic progression consisting of at least three indices, $M|_i$ is of the form $(a_1 \cdots a_n)^d (a_1 \cdots a_m)$, where $d \geq 2$ and $m < n$.

By assumption, there is a message from i to j in $M|_i$, hence $a_k = i!j$ for some k. Since $A_{i,j}$ and $C_{i,j}$ have no $i!j$, we obtain $A_{i,j} = a_1 \cdots a_{k-1}$ and $C_{i,j} = a_{l+1} \cdots a_n a_1 \cdots a_m$, with $l > m$.

In particular, we have $|A_{i,j}| < n$ and $|C_{i,j}| < n$. Since each $M|_i$ contains the last position of the word generated by V, the subword $B^s_{i,j}$ also contains this position, except possibly for the first and the last $B^s_{i,j}$. Hence, every $B^s_{i,j}$ associated with an internal index of π is in $\mathrm{Occ}(B^s_{i,j}, Y, V)$.



□

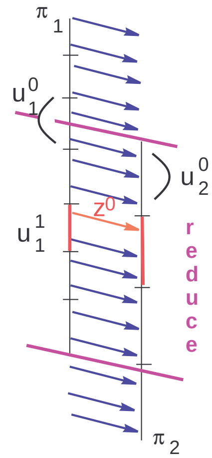
Let now h be a leading process, thus $Y^h = Y$ and $V^h = V$. Let also $j \neq h$ such that j, h communicate in M. Since $M^0|_h$ is an internal index of $\mathrm{Occ}(M|_h, Y, V)$ we can apply Lemma 2 and we obtain that $B^{s,0}_{h,j} \in \mathrm{Occ}(B^s_{h,j}, Y, V)$. Hence, we also have $B^{r,0} j, h \in \mathrm{Occ}(B^r_{j,h}, Y, V)$, since matching sends and receives always appear literally in the same variable. Recall that $M^0|_j \in \mathrm{Occ}(M|_j, Y^j, V^j)$ with $Y^j \leq Y$. Using $B^{r,0}_{j,h} \in \mathrm{Occ}(B^r_{j,h}, Y, V)$ we obtain that $Y \leq Y^j$, hence $Y^j = Y$. Applying the lemma again to $M^0|_j$ we obtain also $V^j = V$, that is j is a leading process too. The result follows for all processes j, due to M being connected. □

Theorem 8 *Let M, N be two nMSCs with M connected. We can check whether M occurs in N in time $O(|M|^2 |N|^2)$.*

```
Pattern-Matching (nMSC M, N)
 For each variable X on the lowest level of hierarchy:
     If M ⊆ X at position pos then return (X, pos);
 For all variables Y,V of N with V ∈ Y:
     Compute Occ(M|_1, Y, V),...,Occ(M|_p, Y, V);
 For every variable Y of N:
     For every process h:
       For every pos(h) at the beginning or end of an
           arithmetic progression of Occ(M|_h, Y):
         Let (M|_h)^{pos(h)} be the corresponding occurrence of M|_h:
         If there exists ((M|_k)^{pos(k)})_{k≠h} compatible with (M|_h)^{pos(h)}
           where for all k, pos(k) ∈ Occ(M|_k, Z^k) with Z^k ≤ Y:
                               Return (Y, (pos(k))_{k ∈ P});
     For every V ∈ Y s.t. ∀i, π_i = Occ(M|_i, Y, V) ≠ ∅:
         For each i, let π_i = (n_i, s_i, k_i);
         Let (t_1,...,t_p, e_1,...,e_p) = Periods(Reduce(π_1,...,π_p));
         Let π'_i = (n_i + t_i s_i, s_i e_i, (k_i − t_i)/e_i)
         If (π'_i)_i ≠ ∅ then return (Y, (π'_i)_i)
```

Notice that we have to restrict $pos(k)$ to be inside Y for every k to ensure that h is leading, which ensures the uniqueness of $pos(k)$ for every k. For simplifying the presentation of the algorithm we will assume below that every process i in M sends at least one message to every other process $j > i$. The algorithm first computes the occurrences $M|_i$ process wise. Then, in the third for-loop, it first considers external indices, corresponding to the second case of Proposition 1. If no pattern is found, the algorithm looks for an occurrence corresponding to the first case of Proposition 1, where $M^0|_i \in \mathrm{Occ}(M|_i, Y, V)$ for every process i. The arithmetic progression $\mathrm{Occ}(M|_i, Y, V)$ is denoted by $\pi_i = (n_i, s_i, k_i)$ above. We denote by u_i the word consisting of the s_i first symbols of $M|_i$. By assumption, each u_i contains both symbols $i!j$ and $i?j$, for all $j > i$. For each $i < j$ we denote by $m_{i,j}$ the number of sends from i to j in u_i, and by $m_{j,i}$ the number of receives from i to j in u_j.

We describe now the subroutines `Reduce` and `Periods` and show that our algorithm returns only occurrences of M which are indeed factors of N. The subroutine `Reduce` restricts the arithmetic progressions $(\pi_1, \ldots, \pi_p)$ by adding an offset to each arithmetic progression π_i. This is done such that for all pairs of distinct processes i, j there exists a send to process j and a receive from j in every occurrence from π_i, such that the matching event belongs to π_j. For instance, in the example below the arithmetic progression π_1 will start after a call of `Reduce` with u_1^0, since the two copies of u_1 before have no send to process 2 such that the matching receive belongs to π_2. Thus, the first two occurrences of u_1 in π_1 will not be used for looking for compatible occurrences. It also reduces the number of occurrences of arithmetic progressions. Reduce takes a quadratic time by computing for every pair of processes i, j the first and the last event on i that sends or receives a message from an occurrence of π_j. We then compute the events which fulfills every constraint.



Let $\pi_1, \ldots, \pi_\wp$ be arithmetic progressions of occurrences of $M|_1, \ldots, M|_\wp$, such that for each pair $i \neq j$ there exists a message between i, j from each u_i in π_i to some u_j in π_j, and vice-versa. That is, $\pi_1, \ldots, \pi_\wp$ is the result of a call of `Reduce`. Let u_i^0 be the first index of each arithmetic progression. The only problem that remains for deciding whether there exist compatible occurrences $M|_i, M|_j$ is that the existence of messages from u_i in π_i to u_j in π_j does not mean that the events match correctly w.r.t. M. We will look for tuples of occurrences of the $M|_i$ that are pairwise compatible by considering sub progressions of the π_i.

From now on we want to determine all tuples $(u_1, \ldots, u_\wp)$ corresponding to the starting positions of pairwise compatible tuples $((M|_1)^0, \ldots, (M|_\wp)^0)$. As we show later, such tuples occur periodically, hence we just need to determine some periods $(\mu_1, \ldots, \mu_\wp) \in \mathbb{N}^\wp$ and the first positions $(u_1^1, \ldots, u_\wp^1)$ from which we can apply these periods.

For all $i < j$ let $z_{i,j} < m_{i,j}$ be the number of events $i!j$ in u_i^0 before the first one that has a matching receive in π_j. Let $z_{j,i} < m_{j,i}$ be the number of $j?i$ in u_j^0 before the first that has a matching send in π_i. In the figure aside, $m_{1,2} = 2$, $m_{2,1} = 3$, $z_{1,2} = 1$ and $z_{2,1} = 0$. Let $z_{i,j}^0$ be such that after reading the first $z_{i,j}^0 + z_{i,j}$ sends from π_i to π_j we arrive at a message consisting of the first $i!j$ of some u_i and the first $j?i$ of some u_j. In the example, we marked as z^0 the earliest message consisting of the first $1!2$ of some u_1 and the first $2?1$ of some u_2, and $z_{1,2}^0 = 3$. So $z_{i,j}^0 + z_{i,j} \equiv 0 \pmod{m_{i,j}}$ and $z_{i,j}^0 + z_{j,i} \equiv 0 \pmod{m_{j,i}}$. Using the Chinese Remainder Theorem the subroutine `Periods` first computes the least solutions $z_{i,j}^0$ modulo $\mathrm{lcm}(m_{i,j}, m_{j,i})$ to the above equations in time $O(\min(|M|_i|, |M|_j|)^3)$. We perform this computation for each pair of processes in overall time $O(|M|^3)$ for obtaining the new period μ_i and the new offset u_i^1. Notice that μ_i divides $\mathrm{lcm}\{m_{i,j} \mid i < j\}$. The restriction of the arithmetic progression π_i according to μ_i, u_i^1 is denoted π_i'.

The first $i!j$ of each u_i in the restricted arithmetic progression π_i' corresponds to the first $j?i$ of some u_j of the unrestricted arithmetic progression π_j. The final step of `Periods` is to compute occurrences of M from $(\pi_i')_i$. Let $x_{i,j}$ be an integer denoting the number of u_j between u_j^1 and the reception of the first message from u_i^1. We want to compute all tuples $(u_i)_{i=1,\wp}$ such that the first $i!j$ of u_i matches the first $j?i$ of u_j. That is, we need a solution $(t_i)_{i=1,\wp}$ of the following system of $\wp(\wp-1)$ linear equations:

$$\mu_i m_{i,j} t_i = x_{i,j} m_{j,i} + \mu_j m_{j,i} t_j$$

Thus, the value of t_1 determines each t_i, modulo some value e_i depending on the $(m_{i,j})_{i,j}$. We can combine the equation for $(1, i)$ with the equation for (i, j) to obtain a system of $\wp(\wp-1)$ equations :

$$\delta_{i,j} t_1 = y_{i,j} + \nu_{i,j} t_j$$

Let $j \leq \wp$. Notice that several of these equations (for different i)concerns the same t_1 and t_j: either all these equation are equivalent, or there exists a unique or no solution at all (we just combine two equations by multiplying per $\nu_{i,j}$ one and by $\nu_{i',j}$ the other and substracting one equation with the other one). If there is a unique solution, we stop the procedure and test this solution in each equation. If this is indeed a solution of the system, we return its value, else we will not find an occurrence of M in this level. Hence, we can assume for the following that there is a unique equation (since all are equivalent) for each j, that is we have a system of $\wp$ equations, where i is fixed.

If $\gcd(\delta_{i,j}, \nu_{i,j})$ does not divide $y_{i,j}$, there is no solution to our system. Else, we can divide $\delta_{i,j}, y_{i,j}, \nu_{i,j}$ by $\gcd(\delta_{i,j}, \nu_{i,j})$, and thus consider only the case where $\gcd(\delta_{i,j}, \nu_{i,j}) = 1$.

Let $\gamma_{i,j}$ be the inverse of $\delta_{i,j}$ modulo $\nu_{i,j}$. Hence the equations are reduced to $\wp$ trivial equations of the form $t_1 \equiv y_{i,j}\gamma_{i,j} \pmod{\nu_{i,j}}$. The subroutine `Periods` finally computes a solution $(t_1, \ldots, t_p)$ using again the Chinese Remainder Theorem and returns $(t_i + u_i^1 - u_i^0, e_i)_i$.

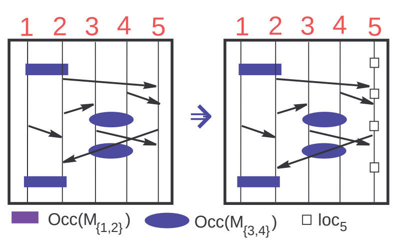
Since the intersection of an arithmetic progression with the periodic set is still an arithmetic progression, in the end we have arithmetic progressions of periods increased by a factor of e_i, that contains only compatible occurrences. A call of `Periods` costs time $O(|M|^3)$.

Remark 3 We can slightly adapt the algorithm for computing all occurrences of M in N. Note that the number of occurrences might be exponential (as in the word case), thus the representation of all occurrences will be implicit.

5.4 Pattern Matching for Non-Connected Patterns

We turn now to the general case where the nMSC pattern M is not connected. We show that the complexity of the algorithm increases just by a factor $O(|\mathcal{C}_M|^2) \leq O(\wp^2)$, namely the square of the number of weakly connected components of M.

It will be helpful in the following to have all processes of N appear in M. This can be enforced by a simple modification of M, N, as depicted below. For each reference Y of N and each process $i \in P_N \setminus P_M$ we add a local action loc_i on process i in Y before each message or reference on i, and before the end of Y. Let $M' = M \cdot \prod_{i \in P_N \setminus P_M}(\mathrm{loc}_i)$. Obviously, M' occurs in N' iff M occurs in N.



Let M, N be nMSCs. For each reference X of M or N, let $\mathcal{C}_X \subseteq 2^{P_M}$ be the set of maximal connected components of the communication graph of X (this is the graph with vertices corresponding to processes and edges between communicating processes). We will denote by $X|_C$ the projection of X over the processes in $C \in \mathcal{C}_X$. In other words, $X = (X|_C)_{C \in \mathcal{C}_X}$ represents the decomposition of the MSC associated with X into weakly connected nMSCs. It follows from the previous section that we can compute in time $O(|M|^2|N|^2)$ a compact representation of all occurrences of $M|_C$ in N, for each $C \in \mathcal{C}_M$. The next definition states when a tuple of occurrences $(M|_C)_{C \in \mathcal{C}_M}$ of the weakly connected components of M corresponds to an occurrence of M in N.

Definition 8. *Let $a \in Occ(M|_C, Y), b \in Occ(M|_D, Y)$ be two occurrences of weakly connected components of M, where $C, D \in \mathcal{C}_M$ and $C \neq D$. Then a, b are called compatible if there is no message in Y from some process in C to some process in D that is sent after a and received before b (or vice versa).*

Lemma 3. *Let $a_C \in Occ(M|_C, Y)$, for all $C \in \mathcal{C}_M$. Then $(a_C)_{C \in \mathcal{C}_M}$ is an occurrence of M in Y iff a_C, a_D are compatible for all $C, D \in \mathcal{C}_M$, $C \neq D$.*

Proof. The implication from left to right follows directly from the definition of pattern. For the converse assume that $(a_C)_{C \in \mathcal{C}_M}$ is not an occurrence of M in Y. This means that there is some chain of messages $(s_k, r_k)_{k=1}^m$ with $P(s_1) \in C$, $P(r_m) \in D$, $P(r_k) = P(s_{k+1})$ for all k, and such that a_C precedes s_1, r_i precedes s_{i+1}, and r_m precedes a_D. Since all processes appear in M, there exist some k and $C', D' \in \mathcal{C}_M$ such that $P(s_k) \in C'$, $P(r_k) \in D'$, $a_{C'}$ precedes s_k and r_k precedes $a_{D'}$. But this means that $a_{C'}, a_{D'}$ are not compatible, contradiction. □

Let $C \in \mathcal{C}_M$. Note that the occurrences of the weakly connected components $M|_C$ in Y are totally ordered by the visual order of Y. This justifies the use of min and max on occurrences of the same weakly connected component in the proposition below.

Proposition 2 *Let $a = (a_C)_{C \in \mathcal{C}_M}, b = (b_C)_{C \in \mathcal{C}_M} \in (Occ(M|_C, Y))_{C \in \mathcal{C}_M}$ be two occurrences of M in Y. Then $(\min(a_C, b_C))_{C \in \mathcal{C}_M}$ and $(\max(a_C, b_C))_{C \in \mathcal{C}_M}$ are also occurrences of M in Y.*

Proof. By Lemma 3 it suffices to check that $\min(a_C, b_C)$, $\min(a_D, b_D)$ are compatible, for all $C, D \in \mathcal{C}_M$, $C \neq D$. The only case to verify is when $\min(a_C, b_C) = a_C < b_C$ and $\min(a_D, b_D) = b_D < a_D$. Assume by contradiction that there is a message from C to D that is sent after a_C and received before b_D. Then a_C and $a_D > b_D$ are not compatible, a contradiction. The case where a message is sent after b_D and received before a_C is symmetrical. □

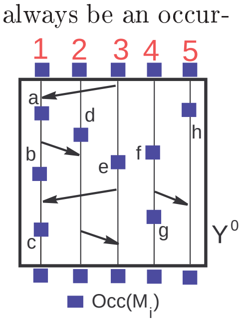
We describe the pattern matching algorithm in a simpler case where the following two conditions hold. First, we assume that every message is on the lowest hierarchical level. This means that macros either consist of references (and local actions) only, or they are MSCs. In other words, we forbid messages crossing references in N. Second, for all references Y, Z with $Z \in Y$ and each occurrence of $M|_C$ in Y either $M|_C$ is included in Z, or it has an empty intersection with Z. That is, we assume that no occurrence of $M|_C$ in Y is split between several references $Z \in Y$. If N satisfies these conditions w.r.t. M, then we call the pair (M, N) **nice**. The general case is technically more involved, but it does not require new ideas.

If M occurs as a pattern of N, then Proposition 2 ensures that there is a unique minimal occurrence of M in N (minimal with respect to the component wise ordering of tuples from $(\mathrm{Occ}(M|_C, N))_{C \in \mathcal{C}_M}$. In order to find the minimal occurrence of M in a reference X of N, we look for compatible minimal occurrences in each reference $Y \in X$. If Y does not contain the complete M, then we need more information about possible components $M|_C$ that are outside Y and that are compatible with the components within Y. Since there may be several references X with $Y \in X$ we encode this additional information by imaginary occurrences denoted $\downarrow_C$ and $\uparrow_C$, for each component $C \in \mathcal{C}_M$. The occurrence $\downarrow_C$ for component C means an occurrence of $M|_C$ after Y, while $\uparrow_C$ for C means an occurrence of $M|_C$ before Y. Thus, we let $\uparrow_C < a_C < \downarrow_C$ for all $a_C \in \mathrm{Occ}(M|_C, Y)$. For $C \neq D$, we say that $\uparrow_C, a_D \in \mathrm{Occ}(M|_C, Y)$ are compatible if there is no message from C to D that is received before a_D in Y (symmetrically for $\downarrow$). The precise definition follows:

Definition 9. *Let Y be a reference of N. Let $E \subseteq \{\neq\uparrow_C, =\downarrow_C \mid C \in \mathcal{C}_M\}$ be a set of constraints. We define $Min_E^Y = (a_C)_{C \in \mathcal{C}_M}$ as the minimal tuple satisfying the following conditions:*

1. *For each $C \in \mathcal{C}_M$, $a_C \in Occ(M|_C, Y) \cup \{\uparrow_C, \downarrow_C\}$.*
2. *The occurrences $(a_C)_{C \in \mathcal{C}_M}$ are pairwise compatible.*
3. *$(a_C)_{C \in \mathcal{C}_M}$ satisfies the constraint E. That is, $(\neq\uparrow_D) \in E$ implies that $a_D \neq\uparrow_D$ and $(=\downarrow_D) \in E$ implies that $a_D =\downarrow_D$.*

Note that the minimal occurrence in the previous definition is well defined, since there exists at least one tuple $(a_C)_{C \in \mathcal{C}_M}$ satisfying the three conditions above, namely $a_C =\downarrow_C$ for all C. In other words, there may always be an occurrence of M after Y.



Example 5. The two extreme constraints correspond to $\mathrm{Min}_\emptyset = (\uparrow_C)_{C \in \mathcal{C}_M}$ and $\mathrm{Min}_{(=\downarrow_C)_{C \in \mathcal{C}_M}} = (\downarrow_C)_{C \in \mathcal{C}_M}$. In the figure to the right we also have:

- $\mathrm{Min}_{\{\neq\uparrow_1\}} = (a, \uparrow_2, e, \uparrow_4, \uparrow_5) = \mathrm{Min}_{\{\neq\uparrow_1, \neq\uparrow_3\}}$.
- $\mathrm{Min}_{\{=\downarrow_2\}} = (b, \downarrow_2, e, \uparrow_4, \uparrow_5)$.
- $\mathrm{Min}_{\{\neq\uparrow_4, =\downarrow_5\}} = (\uparrow_1, \uparrow_2, \uparrow_3, g, \downarrow_5)$.

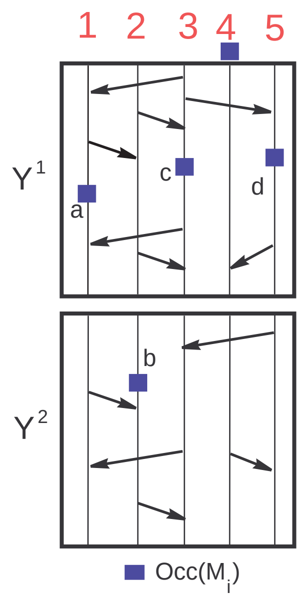
The next lemma shows that it suffices to compute (recursively) the tuples Min_E^Y, for suitable constraints E and references Y of N.

Lemma 4. *Let $(b_C)_{C \in \mathcal{C}_M} = Min^N_{(\neq\uparrow_C)_{C \in \mathcal{C}_M}}$. Then M is a pattern of N iff $b_C \neq\downarrow_C$, for all $C \in \mathcal{C}_M$.*

The problem is that we might need the tuples Min_E^Y for *arbitrary* sets E of constraints (and there are exponentially many). Fortunately, we can avoid the exponential blow-up by computing Min_E^Y only for singletons $E = \{\neq\uparrow_C\}$ and $E = \{\downarrow_C\}$, $C \in \mathcal{C}_M$. We first show that these tuples suffice for computing in polynomial time Min_E^Y for arbitrary E. In a second step, we show that we will need only a polynomial number of constraints E in the recursive step.

Lemma 5. *Let $E, F \subseteq \{\neq\uparrow_C, =\downarrow_C \mid C \in \mathcal{C}_M\}$ be two sets of constraints. Then $Min_{E\cup F}^Y = \max(Min_E^Y, Min_F^Y)$.*

Proof. Let $b = (b_C)_C = \max(\mathrm{Min}_E^Y, \mathrm{Min}_F^Y)$. We have of course $\mathrm{Min}_{E\cup F}^Y \geq \mathrm{Min}_E^Y$ and $\mathrm{Min}_{E\cup F}^Y \geq \mathrm{Min}_F^Y$, hence $\mathrm{Min}_{E\cup F}^Y \geq b$. But $\mathrm{Min}_{E\cup F}^Y$ is the minimal tuple that satisfies the three properties which b satisfies, too: the tuple b has pairwise compatible components b_C and it satisfies the constraints in $E \cup F$. Therefore, $b = \mathrm{Min}_{E\cup F}^Y$. □



$$\mathrm{Min}_{\{\neq\uparrow_5\}}^{Y^1Y^2} = (a, b, c, \uparrow_4, d).$$
$$\mathrm{Min}_{\{\neq\uparrow_5\}}^{Y^1} = (a, \downarrow_2, c, \uparrow_4, d) = \mathrm{Min}_{\{=\downarrow_2, \neq\uparrow_5\}}^{Y^1}.$$
$$\mathrm{Min}_{\{\neq\uparrow_2\}}^{Y^2} = (\uparrow_1, b, \uparrow_3, \uparrow_4, \uparrow_5).$$

Proposition 3 *Assume that the pair (M, N) is nice and consider some reference Y of N and a component $D \in \mathcal{C}_M$. Then $Min_{\{\neq\uparrow_D\}}^Y$ and $Min_{\{=\downarrow_D\}}^Y$ can be computed in time $O(|Y|\wp^2)$ from the tuples $(Min_{\{\neq\uparrow_C\}}^Z)_{C\in\mathcal{C}_M}$ and $(Min_{\{=\downarrow_C\}}^Z)_{C\in\mathcal{C}_M}$, where $Z \in Y$.*

Proof. Assume that any reference Y of N that is not on the lowest hierarchy level has exactly two subreferences, that is $Y = Y^1Y^2$.

We will compute the set of components $E_\downarrow \subseteq \mathcal{C}_M$ that consists of all C such that $M|_C$ has no occurrence in Y^1 which is compatible with the constraints, thus

$M|_C$ must occur either in Y^2 or after Y. In order to do this, we start with $E_\downarrow = \emptyset$ and we augment $E_\downarrow$ as long as there exist a, b with the following properties:

- $(a_C)_C$ is an occurrence in Y^1 with $a_C = \downarrow_C$ iff $C \in E_\downarrow$,
- $(b_C)_C$ is an occurrence in Y^2 with $b_C = \uparrow_C$ iff $C \notin E_\downarrow$.

The algorithm for computing $\mathrm{Min}^Y_{\{\neq\uparrow_D\}}$ is described below (for $\mathrm{Min}^Y_{\{=\downarrow_D\}}$ the reasoning is similar):

```
(1)  Let E↓ = ∅
(2)  Compute (a_C)_C = Min_E^{Y^1}, with E = {≠↑_D} ∪ {=↓_C | C ∈ E↓}
(3)  Let E↓ = { C | a_C = ↓_C}
  // For all C ∈ E↓, M|_C must be in Y^2 or after Y.
(4)  Compute (b_C)_C = Min_{(≠↑_C)_{C∈ E↓}}^{Y^2}
(5)  Let E↓ = { C | b_C ≠ ↑_C}. If E↓ changes, then goto (2).
(6)  Let d_C = b_C if C ∈ E↓, and d_C = a_C, otherwise.
(7)  Return (d_C)_C.
```

Note that each time the set $E_\downarrow$ changes at step (3), it increases by at least one component. Hence, we return to step (2) at most $O(\wp)$ times.

For the running time let us denote by $E^t_\downarrow$ the value of $E_\downarrow$ after t iterations. The t-th iteration needs time $\wp(|E^t_\downarrow| - |E^{t-1}_\downarrow|)$, thus the overall running time is at most $O(\wp^2)$.

If an nMSC has more than two references, then we define several sets $E^i_\downarrow$ to explain the minimal reference Y^i where the occurrence of the projection should be. Considering that for each step, one set $E^i_\downarrow$ has to change, the running time is $\wp^2|Y|$. □

Theorem 9 *We can test whether M occurs as pattern of N in time $O(\mathcal{C}^2_M(|M|^2 |N|^2))$.*

Proof. We show the theorem only for the case where (M, N) is a nice pair. The general case is technically more involved, but does not require new ideas.

Theorem 8 is used for computing first the implicit representation of all occurrences of $M|_C$ in Y, for all components $C \in \mathcal{C}_M$ of M and all references Y of N. For each Y we need then only the position of the minimal occurrence of each $M|_C$ in Y (if any). We compute then $\mathrm{Min}^Y_{\neq\uparrow_C}$ and $\mathrm{Min}^Y_{=\downarrow_C}$ for all components $C \in \mathcal{C}_M$ and references Y of N. We apply Proposition 3 to compute $\mathrm{Min}^Y_{\neq\uparrow_C}$ and $\mathrm{Min}^Y_{=\downarrow_C}$. The time costs are $O(|M|^2|N|^2)$ for the connected components and $O(\wp^3|N|) \le O(|M|^2|N|^2)$ for the additional algorithms looking for compatible components. The overall running time is thus $O(|M|^2|N|^2)$. In the general case we get an additional factor $\mathcal{C}^2_M$, where $\mathcal{C}_M$ is the number of connected components of M, expressing additional constraints due to components $M|_C$ that might be split over several references of N. □

6 Conclusion

In developing new techniques for algorithms on hierarchical MSCs, we proved that algorithms can benefit from the redundancy provided by the use of macros. Namely, it is not a good idea to unfold the hierarchical system since the redundancy is lost. Moreover, we use the hierarchy to lower the running time. We showed that pattern matching and membership can efficiently use the hierarchy, together with techniques stemming from combinatorics, arithmetics and dynamic programming. We believe that similar results can be stated for many other problems on hierarchical MSCs, such as model-checking against properties expressed by template MSCs [10].

Acknowledgment: We wish to thank Marc Zeitoun for his insightful comments on previous versions of this paper, as well as Markus Lohrey and the referees of LATIN'02 for several suggestions for improvement.

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