The Modal Logic of Stepwise Removal

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Abstract

We investigate the modal logic of stepwise removal of objects, both for its intrinsic interest as a logic of quantification without replacement, and as a pilot study to better understand the complexity jumps between dynamic epistemic logics of model transformations and logics of freely chosen graph changes that get registered in a growing memory. After introducing this logic (MLSR) and its corresponding removal modality, we analyze its expressive power and prove a bisimulation characterization theorem. We then provide a complete Hilbert-style axiomatization for the logic of stepwise removal in a hybrid language enriched with nominals and public announcement operators. Next, we show that model-checking for MLSR is PSPACE-complete, while its satisfiability problem is undecidable. Lastly, we consider an issue of fine-structure: the expressive power gained by adding the stepwise removal modality to fragments of first-order logic.

1 Model change and quantification

Logical systems describing various kinds of model change come up when reasoning about scenarios of semantic interpretation that affect a current model, varieties of information update, or more general actions changing a local environment. A typical feature of such systems is the use of dynamic modalities that, when evaluated in a current model $M$, look at what is true in other models $N$, related to $M$ via some relevant cross-model relation.

Now, these dynamic logics come in a wide range of expressive power and computational complexity [Aucher et al., 2018]. Our aim in this small pilot study is to explore one significant border line, where the complexity of the satisfiability problem jumps from decidable to undecidable. In the process, we highlight some further issues, as well as some new proof techniques, as will be explained below.

Dynamic epistemic logics of information update. Here is one recent genre of dynamic logics that can describe model change. When modeling the effects of new information, a natural format changes a current epistemic model to a new one, suitably modified. Say, an event $!\varphi$ of reliable public information that $\varphi$ is the case changes a current pointed model $(M, s)$ to the definable sub-model $(M|\varphi, s)$, whose domain is the set of all points in $M$ that satisfy $\varphi$. Likewise, an event where all agents publicly lose all uncertainty about $\varphi$ takes $(M, s)$ to a model $(M/\varphi, s)$, where the domain stays the same, but the epistemic accessibility relation $\sim$ of $M$ gets replaced by the refinement $s \sim_\varphi t$: i.e., $s \sim t$ and, also, $M, t \models \varphi$ if and only if $M, s \models \varphi$. These and many other model transformations $F$ have matching modalities $[F]\psi$ in dynamic epistemic logics, whose key axioms for $[F]\psi$ give a recursive analysis of when the postconditions $\psi$ hold in terms of what was true before the
F-update (see the survey by van Benthem [2011]). Dynamic epistemic logics are usually decidable if their underlying static logics are: the recursion axioms reduce out the dynamic modalities, at least on full standard universes of epistemic models.

**Sabotage-style graph logics.** Here is a second natural genre of modal logics for describing model change. In the *sabotage game* of van Benthem [2005], arbitrary links in a graph are cut, one by one, by a Demon opposing a Traveler, who, in turn, moves across the graph along still available links. The winning positions of the Demon and the Traveler can be analyzed using standard modalities, together with additional modalities describing what holds in a pointed model after one link has been removed from the current accessibility relation. However, validity in modal logics for various graph games of this sort can be undecidable, and the resulting model theory is quite complex (see [Aucher et al., 2018] and [van Benthem and Liu, 2018]).

This difference in complexity calls for an explanation. The present paper locates its source in the contrast between, on the one hand, the simultaneous removal of points or links in dynamic epistemic logics and, on the other, the stepwise modifications captured by logics for sabotage and related graph games. In doing so, we explore the border between two system designs: dynamic epistemic logics of graph change that reduce effectively to a decidable static base language—and, hence, to what is true in the initial model, which already ‘pre-encodes’ the effects of changes—and, on the other hand, undecidable sabotage-type logics of graph change operations, whose effects are not pre-encoded in the original model, but rather depend on a growing ‘memory’ of previous changes.

To make this concrete, here is a simplest dynamic epistemic logic turned ‘stepwise’. For simplicity, we focus on point deletion, rather than link deletion.

**A stepwise update modality.** Consider the standard language of basic modal logic, augmented with a dynamic modality \(\langle \neg \varphi \rangle \psi\). Given a relational model \(\mathcal{M} = (W, R, V)\), with \(R \subseteq W \times W\) and \(V\) a valuation, the satisfaction clause for \(\langle \neg \varphi \rangle \psi\) is given by

\[
\mathcal{M}, s \models \langle \neg \varphi \rangle \psi \text{ iff there is a point } t \neq s \text{ in } \mathcal{M} \text{ with } \mathcal{M}, t \models \varphi \text{ and } \mathcal{M} - \{t\}, s \models \psi,
\]

where \(\mathcal{M} - \{t\}\) is the submodel of \(\mathcal{M}\) having just the point \(t\) removed from its domain. This system of what may be called stepwise point removal (MLSR) will be studied here as an intermediate case between the simplest dynamic epistemic logic of public announcements, where all points satisfying \(\varphi\) are removed simultaneously during an update, and a simple sabotage logic for stepwise graph change.

**Quantification without replacement.** The language introduced here has various further interpretations. For instance, it can be seen as a logic of ‘interventions’ that minimally change some given model to make some specified new properties true [Renardel de Lavalette, 2001]. But the system has an even more general logical motivation, which is not tied to information updates or any other specific application.

Consider the evaluation of restricted existential quantifiers \(\exists x \varphi(x) \cdot \psi(x)\) in first-order logic (FOL). One searches for an object \(d\) satisfying \(\varphi\) and then checks whether \(d\) also satisfies \(\psi\). In this second stage, the model has not changed: the witness \(d\) is still in the domain and it influences the evaluation of \(\psi\). Call this process ‘quantification with replacement’.
Now, it has been claimed [Hintikka and Sandu, 1997] that quantifiers in natural language can also behave differently: witness, for instance, the natural sense in which the distrust in “John distrusted everyone” does not apply to John himself. Even though this may be an idiosyncrasy of natural language, it clearly makes sense to explore quantification without replacement as a model for evaluation procedures that change domains [Gabbay, 2013]:

\[ \exists x (\varphi | \psi) \]

This quantifier form is clearly definable in FOL with identity, but, taken by itself, it suggests its own model theory and proof theory. Moreover, as we shall see, adding quantification without replacement to weaker fragments of the first-order language, such as monadic predicate logic or basic modal logic, produces much less simple effects.

The system MLSR. The system MLSR of stepwise object removal studied in this paper provides a simple modal setting for bringing all of this out. Its syntax is that of the basic modal language with proposition letters, \( \neg, \lor, \Box \), plus the additional modality \( \langle \neg \varphi \rangle \psi \), whose semantics was given above. Occasionally, we will also use this language extended with a ‘public announcement’, or relativization, modality \( \langle ! \varphi \rangle \psi \) describing what is true in restrictions to definable subdomains:

\[ \mathcal{M}, s \models \langle ! \varphi \rangle \psi \text{ iff } \mathcal{M}, s \models \varphi \text{ and } \mathcal{M}|s, s \models \psi, \]

with \( \mathcal{M}|s \) the submodel of \( \mathcal{M} \) consisting of all and only the points in \( \mathcal{M} \) where \( \varphi \) is true.

Outline of the paper. In this paper, we study the essential features of this modal system. In Section 2, we analyze the expressive power of MLSR by providing a first-order translation and a semantic characterization in terms of bisimulation invariance. This mainly requires straightforward adaptations of known techniques. Section 3 presents a complete axiomatization for MLSR, based on a new idea of mixing standard relativization with stepwise removal, which may very well be applicable to many other logics of graph change, for which Hilbert-style axiomatizations have long been an open problem. In Section 4, we first analyze the computational complexity of model checking for MLSR, which turns out to be PSPACE-complete. This analysis uses a reduction technique from Löding and Rohde [2003] which deserves to be better known in modal logic. Next, we prove that the satisfiability problem for MLSR is undecidable using a tiling argument familiar from the modal logic literature [Marx, 2006; Areces et al., 2015]. In Section 5, we then raise a more general issue: namely, what the addition of quantification without replacement does to various fragments of first-order logic. In particular, we show that, when added to monadic first-order logic, the modality \( \langle \neg \varphi \rangle \psi \) essentially allows us to count, boosting the expressive power of monadic first-order logic to that of monadic first-order logic with identity.

In summary, we locate the threshold of complexity in the stepwise character of the modality for point removal, leading to the need for a computational device for maintaining a memory of deleted points, whose complexity equals that of arbitrary tiling problems, and computations of Turing machines. In the process, we also raise new types of questions about modal logics of graph change, and we advertise and introduce some techniques that deserve to be better known among modal logicians.
2 Basics of expressive power

Some definable notions. The language of MLSR can define various modal operators from hybrid logic [Areces and ten Cate, 2006] that go beyond the basic modal language. For instance, the difference modality $D \varphi$ (‘$\varphi$ is true at some different point’) can be defined as $\langle -\varphi \rangle \top$, and this, in turn, allows to define the existential modality $E \varphi$ as $\varphi \lor D \varphi$. MLSR can also count all finite cardinalities, using suitably iterated formulas $\langle -\top \rangle \ldots \langle -\top \rangle \top$ $k$ times, which express that a model has at least $k$ objects different from the current point of evaluation. In addition, MLSR can define quite a few finite relational graphs up to isomorphism. For instance, let $\rho_2$ be the formula defining domain size 2, and let $U$ be the universal modality (i.e., $U \varphi = \varphi \land [-\neg \varphi] \bot$). The following observation requires an easy exercise in understanding what our language can express.

**Fact 2.1.** The MLSR-formula $\rho_2 \land U \langle -\top \rangle \Box \bot \land \Diamond \Diamond \top$ defines a two-point irreflexive loop.

However, not every finite graph is definable, as we shall soon see.

SR-bisimulation. The semantic invariance matching this language is as follows.

**Definition 2.2.** A relation $Z$ between a set of pointed relational models is an SR-bisimulation if it is a modal bisimulation in the ordinary sense, where the back and forth clauses stay inside the same models $M, N$, while, in addition,

(a) if $(M, s)Z(N, t)$ and $u \in M$ with $u \neq s$, then there is a $v \in N$ such that $v \neq t$ and $(M - \{u\}, s)Z(N - \{v\}, t)$,

(b) the analogous clause in the converse direction.

Note that this definition imposes some minimal closure conditions on the set of models involved in the above clauses that are easy to spell out. The following property is proved by a standard induction on formulas.

**Fact 2.3.** MLSR-formulas are invariant for SR-bisimulations.

Now we can give an example of two finite graphs that are not definable up to isomorphism and, in line with this, a first-order formula that is not in MLSR.

**Fact 2.4.** $\forall y (Rxy \lor Ryx)$ is not MLSR-definable.

**Proof.** Consider the model $M$ consisting of two isolated reflexive points and the model $N$ consisting of two points with the universal relation, plus all their submodels. By checking all clauses, one sees that the universal relation $Z$ between all pairs $(M, x)$ and $(N, y)$ plus all links between the 1-point pointed sub-models of $M$ and $N$ is an SR-bisimulation. But, clearly, connectedness holds in $N$, but not in $M$. □

This new logical system still lies inside standard first-order logic.
**Fact 2.5.** There is an effective meaning-preserving translation from MLSR into FOL.

*Proof.* We define the following compositional translation $\tau(\varphi, y, X)$ from MLSR-formulas $\varphi$ to first-order formulas, where $y$ is a free variable and $X$ a finite set of variables:

$$
\begin{align*}
\tau(p, y, X) &= Py, \\
\tau(\neg \varphi, y, X) &= \neg \tau(\varphi, y, X), \\
\tau(\varphi \lor \psi, y, X) &= \tau(\varphi, y, X) \lor \tau(\psi, y, X), \\
\tau(\Diamond \varphi, y, X) &= \exists z \left( Ryz \land \bigwedge_{x \in X} \neg (z = x) \land \tau(\varphi, z, X) \right), \\
\tau((\neg \varphi)\psi, y, X) &= \exists z \left( \neg (z = y) \land \bigwedge_{x \in X} \neg (z = x) \land \tau(\varphi, z, X) \land \tau(\psi, y, X \cup \{z\}) \right).
\end{align*}
$$

Let $(M, s)$ be any pointed model and $D = \{d_1, ..., d_k\}$ a finite set of points in $M$ of size $k$. The following equivalence is shown by a straightforward induction on MLSR-formulas $\varphi$ and sets of variables $X = \{x_1, ..., x_k\}$ of size $k$:

$$M - D, s \models \varphi \iff M, a[y/s, X/D] \models \tau(\varphi, y, X),$$

where $a[y/s, X/D](y) = s$ and $a[y/s, X/D](x_i) = d_i$. As a special case, there is an equivalence for MLSR-formulas in ordinary models $M$ with $D = \emptyset$. $\Box$

**Remark.** The set $X$ in this translation serves as a finite memory storing the points that have already been deleted. This is an essential difference with first-order translations for standard modal languages, which usually lie inside fixed finite-variable fragments.

A simple adaptation of a well-known model-theoretic argument for standard modal logic (cf. [Blackburn et al., 2011]) yields the following result.

**Theorem 2.6.** The following assertions are equivalent for all first-order formulas $\varphi(x)$ in the signature of our models, with one free variable:

(a) $\varphi(x)$ is invariant for SR-bisimulation;

(b) $\varphi(x)$ is equivalent to the translation of some MLSR-formula.

*Proof.* We merely outline the points that need attention in the non-trivial direction from (a) to (b). Let $\mathcal{SR}$ denote the MLSR-fragment of first-order logic (that is, all first-order formulas equivalent to translations of MLSR formulas via the translation $\tau$ from Fact 2.5). As usual, one shows that $\varphi(x)$ is a semantic consequence of the set $\mathcal{C}_x(\varphi)$ of its $\mathcal{SR}$-consequences and then applies Compactness to get an $\mathcal{SR}$-equivalent. We thus need to show that $\mathcal{C}_x(\varphi) \models \varphi(x)$. Suppose $M, s \models \mathcal{C}_x(\varphi)$. A standard compactness argument shows that there is a model $N$ and $t \in N$ such that $(M, s)$ and $(N, t)$ are $\mathcal{SR}$-equivalent, while $N, t \models \varphi(x)$. These models are then extended to $\omega$-saturated elementary extensions $(M^+, s)$ and $(N^+, t)$. We use first-order saturation allowing finite sets of parameters consisting of designated objects in the models; in turn, the finitely satisfiable sets of first-order formulas to be saturated can have a finite set of free variables (not just one, as in the argument for basic modal logic). This is needed for the saturation argument to follow.
Now we define a relation $Z$ between pointed models $(M^+ - D, u)$ and $(N^+ - E, v)$, with $E, D$ of the same finite size, which holds if $(M^+ - D, u)$ and $(N^+ - E, v)$ satisfy the same $\mathcal{SR}$-formulas. Using saturation, it can be shown that $Z$ is an $\mathcal{SR}$-bisimulation, where the argument for the modality $\Diamond \varphi$ is standard, while the one for $\langle -\varphi \rangle \psi$ in terms of removing single objects goes as follows. Take $(M^+ - D, u)$ and $w$. Now, let

$$\Gamma(y) := \{ \gamma(y) \in \mathcal{SR} \mid M^+ - D, w \models \gamma \}$$
$$\Delta(x) := \{ \delta(x) \in \mathcal{SR} \mid M^+ - (D \cup \{w\}), u \models \delta \}$$

and consider the set of first-order formulas

$$p(x, y) := \{ -(y = x) \} \cup \Gamma(y) \cup \Delta(x)$$

This set is finitely satisfiable in $(M^+ - D, u, w)$ (interpreting $x$ as $u$ and $y$ as $w$). For each of its finite subsets $\{ -(y = x) \} \cup \Gamma'(y) \cup \Delta'(x)$, we have

$$M^+ - D, u, w \models -(y = x) \land \bigwedge \Gamma'(y) \land \bigwedge \Delta'(x),$$

which means that

$$M^+ - D, u, w \models \exists y (-(y = x) \land \bigwedge \Gamma'(y) \land \bigwedge \Delta'(x)),$$

and this formula is in $\mathcal{SR}$ (it is equivalent to the translation of a $\langle -\varphi \rangle \psi$ formula). This means that the formula also holds in $(N^+ - E, v)$. Thus, every finite subset of $p(x, y)$ is satisfiable in $(N^+ - E, v)$ (interpreting $x$ as $v$). In other words, expanding the language with a new constant symbol $c$, the 1-type $p(c, y)$ is finitely satisfiable in $(N^+ - E, v)$ (fixing the interpretation of $c$ as $v$). Then, by saturation, the type is realized in $(N^+ - E, v)$: we can thus find an object in $N^+ - E$ matching the given $w$, as required for an $\mathcal{SR}$-bisimulation. □

Finally, the first-order translation for MLSR goes into a fragment of full FOL, and it can also be given directly into a hybrid logic [Areces and ten Cate, 2006].

**Fact 2.7.** MLSR can be translated into $H(@, \downarrow)$, the hybrid language corresponding to the bounded fragment of first-order logic.

**Proof.** For each formula of the form $\langle -\varphi \rangle \psi$ and sequence of nominals $\bar{n} = (n_1, \ldots, n_\ell)$, one can define the following translation clause:

$$\sigma(\langle -\varphi \rangle \psi)_{\vec{n}} = \downarrow_m \cdot E \downarrow_k \cdot \left( \neg m \land \bigwedge_{i=1}^\ell \neg n_i \land \sigma(\varphi)_{\vec{n}} \land @_m \sigma(\psi)_{\vec{n,k}} \right)$$

Then, the map sending each MLSR formula $\varphi$ to $\sigma(\varphi)$ (with $\bar{n}$ empty) is a translation from MLSR into the hybrid logic $H(@, \downarrow)$ defining the bounded fragment. □

However, even this map is not surjective. Our earlier first-order formula $\forall y (Rxy \lor Ryx)$ is in $H(@, \downarrow)$, but it is not definable in MLSR (Fact 2.4).

6
3 Axiomatization

Thanks to the first-order translation, the valid formulas of MLSR are effectively axiomatizable. But more immediate information comes from explicit modal laws. For instance, the removal modality \( \langle - \varphi \rangle \psi \) distributes over disjunction in both of its arguments:

**Fact 3.1.** The following formulas are both valid:

\[
\langle - (\varphi_1 \lor \varphi_2) \rangle \psi \leftrightarrow (\langle - \varphi_1 \rangle \psi \lor \langle - \varphi_2 \rangle \psi)
\]

\[
\langle - \psi \rangle (\varphi_1 \lor \varphi_2) \leftrightarrow (\langle - \psi \rangle \varphi_1 \lor \langle - \psi \rangle \varphi_2)
\]

To obtain an explicit modal axiomatization, we extend the language of MLSR with a countable set \( \text{NOM} \) of nominals, each standing for either a unique point in the model, or not denoting at all (this small technical deviation from hybrid logic will be helpful later on.) We also add standard public announcement modalities \( \langle ! \varphi \rangle \psi \) from dynamic epistemic logic, whose interpretation was given in Section 1. This will turn out to be useful, even though the axiom system to follow feature no recursion axioms in the usual dynamic epistemic style for the removal modality. For simplicity, we retain the name MLSR for this logic.

**Remark.** There seem to be no modal recursion axioms inverting the operator order for combinations \( \langle - \varphi \rangle \langle ! \alpha \rangle \psi \) or \( \langle ! \alpha \rangle \langle - \varphi \rangle \psi \). For example, \( \langle ! \alpha \rangle \langle - \varphi \rangle \psi \) is not equivalent to \( \alpha \land \langle ! \alpha \rangle \langle - \varphi \rangle \psi \) (consider, for instance, the case where \( \alpha = \Diamond p \), \( \varphi = \Box \bot \) and \( \psi = \top \)). This feature of the modal language may be contrasted with how first-order logic augmented with an explicit syntactic operator of relativization would write this recursion:

\[
(\exists x (\varphi | \psi))^{(\alpha)}(y) \leftrightarrow \exists x (\alpha(x) \land \varphi^{(\alpha)}(y) \land \alpha^{(\alpha)}(\psi | \alpha)^{\alpha}(\cdot | y))
\]

**Definition 3.2.** The syntax of MLSR is given by

\[
\varphi := p \mid n \mid \neg \varphi \mid (\varphi \lor \varphi) \mid \Diamond \varphi \mid \langle ! \varphi \rangle \varphi \mid \langle - \varphi \rangle \varphi \mid U \varphi,
\]

with \( p \in \text{PROP} \), \( n \in \text{NOM} \). Dual modal operators \( \Box \), \( \langle ! \varphi \rangle \), \( \langle - \varphi \rangle \) and \( E \) are defined as usual.

Note that it is not necessary to add the \( @_n \) operator from hybrid logic as a primitive symbol, for it can be defined using the universal modality: in our setting with partial nominals, \( @_n \varphi \) is simply a shorthand for \( U(n \rightarrow \varphi) \).

**Definition 3.3.** The logic MLSR (see Figure 1) consists of:

- the axioms and rules of the minimal normal modal logic extended with the universal modality [Blackburn et al., 2011];
- the axioms and rules for public announcement logic PAL, with reduction axioms for nominals and the universal modality [van Benthem, 2011];
- the axiom \( E(n \land \varphi) \rightarrow U(n \rightarrow \varphi) \), which we denote by (II);

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\(^1\)For a further study of combining dynamic epistemic proof systems with hybrid logic, see [Hansen, 2011].
• the Name Rule and the Paste Rule from hybrid logic [Areces and ten Cate, 2006];

• the following two principles connecting the stepwise removal modality with the public announcement modality:

(Mix Axiom) $\langle E (n \land \alpha) \rangle \land \langle \neg n \rangle \varphi \rightarrow \langle \neg \alpha \rangle \varphi$;

(Mix Rule) If $\vdash (\sigma \land \langle \neg \varphi \rangle \neg n \land \langle \varphi \rangle E (n \land \alpha)) \rightarrow \neg (\neg \varphi) \langle \neg n \rangle \psi$, then $\vdash \sigma \rightarrow \neg (\neg \varphi) \langle \neg \alpha \rangle \psi$, for $n \notin \sigma, \varphi, \alpha, \psi$.

**Fact 3.4.** The Mix Axiom is valid, and the Mix Rule is semantically sound.

**Remark.** The axioms for the universal modality, together with the axiom (H) above, ensure that all of the axioms for the basic hybrid logic are derivable, except for the following three:

• $\@_n p \leftrightarrow \neg \@_n \neg p$ (Self-dual Axiom);

• $\@_n m \leftrightarrow \@_m n$ (Symmetry Axiom);

• $\@_n \@_m p \leftrightarrow \@_m p$ (Agree Axiom).

Crucially, the reason why MLSR does not include all of the axioms for the basic hybrid language is that nominals can fail to denote in models after an update. In particular, after the deletion of a state named by $n$, the formula $\@_n \bot$—which is equivalent to $\neg E n$—holds. The three axioms above are valid only in a weaker form with an additional premise $\neg \@_n \bot$ (or $\neg \@_n \bot \land \neg \@_m \bot$): each of these weakened forms, however, is already derivable from the rules in the system as it is. Note also that we do not need to add $\@_n$-generalization to the list of proof rules, for it already follows from the rules for the universal modality.

In contrast with hybrid logics, the language of MLSR is expressive enough to capture various global properties of our semantics, such as the fact that nominals hold at one state at most, expressed by the formula

$$k \rightarrow \neg (\neg k) \top.$$ 

It is worth noting that the Mix Rule is powerful enough to derive it.

**Observation 3.5.** The formula $k \rightarrow \neg (\neg k) \top$ is an MLSR validity for any $k \in \text{NOM}$.

**Proof.** Taking $\varphi, \psi = \top$ and $\sigma, \alpha = k$, the following rule of inference is (equivalent to) an instance of the Mix Rule:

$$\frac{(k \land \neg n \land E (n \land k)) \rightarrow \neg (\neg n) \top}{k \rightarrow \neg (\neg k) \top}$$

Thus, we only need to observe that $\vdash_{\text{MLSR}} (k \land \neg n \land E (n \land k)) \rightarrow \neg (\neg n) \top$. This follows by noting that $\vdash_{\text{MLSR}} (k \land E (n \land k)) \rightarrow n$, so that the antecedent $(k \land \neg n \land E (n \land k))$ is an MLSR-contradiction. To see this, consider that $E (k \land n) \rightarrow U (k \rightarrow n)$ is an instance of the (H) axiom, and simple laws for $E$ and $U$ give us $\vdash_{\text{MLSR}} (k \land U (k \rightarrow n)) \rightarrow n$. So, we have $\vdash_{\text{MLSR}} \neg (k \land \neg n \land E (n \land k))$ and the Mix Rule yields $\vdash_{\text{MLSR}} k \rightarrow \neg (\neg k) \top$. 

We now proceed to prove completeness.
### System MLSR

**Axioms:**

- **K axioms for □**

- **Axioms for the universal modality:**
  
  \[
  U\varphi \rightarrow \varphi \\
  U(\varphi \rightarrow \psi) \rightarrow (U\varphi \rightarrow U\psi) \\
  \varphi \rightarrow UE\varphi \\
  U\varphi \rightarrow UU\varphi \\
  U\varphi \rightarrow □U\varphi
  \]

- **Axioms for PAL:**
  
  \[
  ⟨!\varphi⟩p \leftrightarrow \varphi \land p \ (p \in PROP) \\
  ⟨!\varphi⟩n \leftrightarrow \varphi \land n \ (n \in NOM) \\
  ⟨!\varphi⟩¬\psi \leftrightarrow (\varphi \land ¬⟨!\varphi⟩\psi) \\
  ⟨!\varphi⟩(\psi \lor α) \leftrightarrow (⟨!\varphi⟩\psi \lor ⟨!\varphi⟩α)
  \]

- **Hybrid Axiom:**
  
  \[
  (H) \ E(n \land \varphi) \rightarrow U(n \rightarrow \varphi)
  \]

- **Axiom for the deletion modality:**
  
  \[
  (Mix) \ (E(n \land α) \land ⟨!n⟩\varphi) \rightarrow ⟨¬α⟩\varphi
  \]

**Inference rules:**

- **(MP)** \[
  \varphi, \varphi \rightarrow \psi \\
  \]

- **(Nec)** \[
  \varphi \rightarrow U\varphi
  \]

- **(Mix Rule)** \[
  \frac{(σ \land ⟨!\varphi⟩\neg n \land ⟨!\varphi⟩E(n \land α)) \rightarrow ¬⟨!\varphi⟩⟨!n⟩\psi}{σ \rightarrow ¬⟨!\varphi⟩⟨¬α⟩\psi} \quad (n \not\in σ, \varphi, α, ψ)
  \]

- **(Name)** \[
  m \rightarrow \varphi \quad (m \not\in \varphi)
  \]

- **(Paste)** \[
  \frac{(U(n \rightarrow ♦m) \land U(m \rightarrow \varphi)) \rightarrow α}{U(n \rightarrow ♦\varphi) \rightarrow α} \quad (m \not\in \varphi, α)
  \]

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Figure 1: The Hilbert-style proof system for MLSR.
**Theorem 3.6.** The system MLR is complete for validity in the given semantics.

The Henkin-style completeness proof follows standard modal and hybrid lines [Blackburn et al., 2011], but there are some interesting new features, that will be highlighted in what follows. We begin with a Lindenbaum Lemma.

**Definition 3.7 (Named, Pasted, Mixed).** An MLR-maximally consistent set (MCS) of formulas \( \Gamma \) is named if and only if it contains a nominal. It is pasted if and only if \( U(n \rightarrow \diamond \varphi) \in \Gamma \) implies that there is some nominal \( m \) such that \( (U(n \rightarrow m) \land U(m \rightarrow \varphi)) \in \Gamma \). Lastly, it is mixed if and only if \( (\langle \diamond \varphi \rangle \neg n \land (\langle \diamond \varphi \rangle E(n \land \alpha) \land (\langle \diamond \varphi \rangle \langle \neg n \rangle \psi) \in \Gamma \).

**Lemma 3.8 (Lindenbaum Lemma).** Every MLR-consistent set of formulas can be extended to an MLR-MCS that is named, pasted and mixed.

**Proof.** Naming and pasting work in exactly the same way as in the context of the completeness proof for the basic hybrid logic. To mix, we only have to ensure that, throughout the inductive construction, whenever we consistently add a formula of the form \( (\langle \diamond \varphi \rangle \langle \neg \alpha \rangle \psi) \) to a consistent, named set of formulas \( \Sigma \), the formulas \( (\langle \diamond \varphi \rangle \neg n, \langle \diamond \varphi \rangle E(n \land \alpha) \land (\langle \diamond \varphi \rangle \langle \neg n \rangle \psi) \) are also added to \( \Sigma \)—where \( n \) is the first nominal in the new nominal enumeration that occurs in neither \( \Sigma \) nor \( (\langle \diamond \varphi \rangle \langle \neg \alpha \rangle \psi) \). Clearly, the set \( \Sigma \cup \{\langle \diamond \varphi \rangle \neg n, \langle \diamond \varphi \rangle E(n \land \alpha), \langle \neg n \rangle \psi\} \) is consistent, given that \( \Sigma \) is consistent. For if not, then for some conjunction \( \sigma \) of formulas from \( \Sigma \), the implication \( (\sigma \land (\langle \diamond \varphi \rangle \neg n \land (\langle \diamond \varphi \rangle E(n \land \alpha))) \rightarrow \neg (\langle \diamond \varphi \rangle \langle \neg n \rangle \psi) \) would be provable. But then, by the Mix Rule, the negation \( \neg (\langle \diamond \varphi \rangle \langle \neg \alpha \rangle \psi) \) is provable from \( \Sigma \), contradicting our initial assumption that \( \Sigma \cup \{\langle \diamond \varphi \rangle \langle \neg \alpha \rangle \psi\} \) is consistent. \( \Box \)

Given a set of formulas \( \Gamma \), let \( \Delta_n := \{\varphi \in \text{MLSR} \mid U(n \rightarrow \varphi) \in \Gamma\} \). Now, take an MLR-MCS \( \Gamma \) that is named, pasted, and mixed. Let

\[ \mathcal{W} = \{\Gamma\} \cup \{\Delta_n \mid n \in \text{NOM}, E_n \in \Gamma\}. \]

The relations are defined as usual:

\[ R_{\diamond}(w, v) \iff \text{for all } \varphi, \text{ if } \varphi \in v \text{ then } \diamond \varphi \in w; \]

\[ R_{E}(w, v) \iff \text{for all } \varphi, \text{ if } \varphi \in v \text{ then } E\varphi \in w. \]

We now define the upper Henkin model.

**Definition 3.9 (Upper Henkin Model).** The upper Henkin model \( \mathcal{M} \) generated by \( \Gamma \) is defined as the structure \( (W, R_{\diamond}, R_{E}, V) \), where

- \( W := [\Gamma]_{R_{E}}, \) the equivalence class of \( \Gamma \) in \( \mathcal{W} \) under \( R_{E}; \)
- the relations \( R_{\diamond} \) and \( R_{E} \) are, respectively, \( R_{\diamond} \) and \( R_{E} \) restricted to \( W; \)
- the valuation \( V \) is given by \( V(p) = \{w \in W \mid p \in w\} \) for all proposition letters \( p \) and, for all nominals \( n, \)

\[ V(n) = \begin{cases} \Delta_n \text{ if } E_n \in \Gamma; \\ \emptyset \text{ otherwise.} \end{cases} \]
Setting the domain to be the equivalence class of $\Gamma$ under $R_E$ ensures that $R_E$ is the universal relation in the model. A simple further argument shows that the valuation is well-defined for nominals, and each $w \in W$ is an MLSR-MCS that contains a nominal. Then, the Existence Lemma, stated below, follows easily, with a minor twist to take care of non-referring nominals.

**Lemma 3.10 (Existence Lemma).** Let $\Gamma$ be an MLSR-MCS and $M = (W, R_\delta, R_E, V)$ the upper Henkin model yielded by $\Gamma$. Suppose that $u \in W$ and $\varphi \in u$. Then, there is an object $v \in W$ such that $R_\delta uv$ and $\varphi \in v$. Similarly, if $E \varphi \in u$, then there is some $v \in W$ such that $R_E uv$ and $\varphi \in v$.

Next, we define derived Henkin models, which capture the effects of (sequences of) updates on the upper Henkin model $M$.

**Definition 3.11 (Derived Henkin Model).** For each finite sequence of MLSR-formulas $\mathfrak{F} = (\varphi_1, \ldots, \varphi_k)$, viewed as successive truthful updates, the derived Henkin model $M : \mathfrak{F}$ is defined as the structure $(W^\mathfrak{F}, R_\delta^\mathfrak{F}, R_E^\mathfrak{F}, V^\mathfrak{F})$, where

- $W^\mathfrak{F} := \{(w, \mathfrak{F}) \mid w \in W$ and $\langle \varphi_1 \rangle \cdots \langle \varphi_k \rangle \top \in w\}$;
- for $(w, \mathfrak{F}), (v, \mathfrak{F}) \in W^\mathfrak{F}$, $R_\delta^\mathfrak{F}((w, \mathfrak{F}), (v, \mathfrak{F}))$ (resp., $R_E^\mathfrak{F}((w, \mathfrak{F}), (v, \mathfrak{F}))$) if and only if $R_\delta uv$ (resp., $R_E uv$) in the upper Henkin model $M$;
- $V^\mathfrak{F}(p) := \{(w, \mathfrak{F}) \mid p \in \text{PROP}$ and $V^\mathfrak{F}(n) := \{(w, \mathfrak{F}) \mid n \in w\}$ for $n \in \text{NOM}$.

Given $\mathfrak{F} = (\varphi_1, \ldots, \varphi_k)$, the notation $\langle \mathfrak{F} \rangle$ stands for $\langle \varphi_1 \rangle \cdots \langle \varphi_k \rangle$. Points in the derived Henkin model yielded by $\mathfrak{F}$ are thus sequences $(w, \varphi_1, \ldots, \varphi_k)$, where $w$ is a MCS in the upper Henkin model that contains the pre-condition formula $\text{pre}(\mathfrak{F}) = \langle \mathfrak{F} \rangle \top = \langle \langle \varphi_1 \rangle \cdots \langle \varphi_k \rangle \rangle$. The accessibility relations stay as they were for the initial points of the sequences in the upper Henkin model. Likewise, the valuation at each sequence stays the same as that for its initial point. To each point $(w, \varphi_1, \ldots, \varphi_k)$ in the derived Henkin model $M : \mathfrak{F}$, we associate the set of formulas

$$\Phi(M, \mathfrak{F}, w) := \{\alpha \mid \langle \langle \varphi_1 \rangle \cdots \langle \varphi_k \rangle \rangle \alpha \in w\}.$$ 

These sets can be seen as determining the relations between states.\(^2\)

**Proposition 3.12.** Given $(w, \mathfrak{F})$ and $(v, \mathfrak{F})$ in $W^\mathfrak{F}$, the following are equivalent:

(a) $R_\delta^\mathfrak{F}((w, \mathfrak{F}), (v, \mathfrak{F}))$;

(b) for all formulas $\alpha$, if $\alpha \in \Phi(M, \mathfrak{F}, v)$, then $\Diamond \alpha \in \Phi(M, \mathfrak{F}, w)$.

**Proof.** (a) $\Rightarrow$ (b) Since $R_\delta^\mathfrak{F}((w, \mathfrak{F}), (v, \mathfrak{F}))$, $R_\delta uv$. Now, suppose that $\alpha \in \Phi(M, \mathfrak{F}, v)$. Then, $\langle \mathfrak{F} \rangle \alpha \in v$. This, together with $R_\delta uv$, yields $\Diamond \langle \mathfrak{F} \rangle \alpha \in w$. Next, by repeatedly applying the

\(^2\)Alternatively, one can think of the sets $\Phi(M, \mathfrak{F}, w)$ as themselves being the worlds in the derived Henkin models, but our notation provides unique names.
PAL-equivalence $\vdash_{\text{MLSR}} \langle \alpha \rangle \beta \leftrightarrow \alpha \land \langle \alpha \rangle \beta$ and the distributivity of the public announcement modality over conjunctions, we have that

$$
\vdash_{\text{MLSR}} \langle \varphi_1 \rangle \ldots \langle \varphi_k \rangle T \leftrightarrow \varphi_1 \land \langle \varphi_1 \rangle \varphi_2 \land \langle \varphi_1 \rangle \langle \varphi_2 \rangle \varphi_3 \land \ldots \land \langle \varphi_1 \rangle \ldots \langle \varphi_{k-1} \rangle \varphi_k
$$

$$
\leftrightarrow \bigwedge_{i=1}^{k} \langle \varphi_1 \rangle \ldots \langle \varphi_{i-1} \rangle \varphi_i
$$

Secondly, by repeatedly applying the PAL recursion axioms for the public announcement modality, we get that

$$
\vdash_{\text{MLSR}} \left( \bigwedge_{i=1}^{k} \langle \varphi_1 \rangle \ldots \langle \varphi_{i-1} \rangle \varphi_i \land \Diamond \langle \langle \varphi \rangle \alpha \rangle \right) \leftrightarrow \langle \langle \varphi \rangle \Diamond \alpha \rangle
$$

So, $\vdash_{\text{MLSR}} (\operatorname{pre}(\varphi) \land \Diamond \langle \langle \varphi \rangle \alpha \rangle) \leftrightarrow \langle \langle \varphi \rangle \Diamond \alpha \rangle$. Since $\operatorname{pre}(\varphi) \in w$ and $\Diamond \langle \langle \varphi \rangle \alpha \rangle \in w$, we then have that $\langle \langle \varphi \rangle \Diamond \alpha \rangle \in w$. Hence, $\Diamond \alpha \in \Phi(\mathcal{M}, \varphi, w)$.

(b) $\Rightarrow$ (a) We first observe the following, for any $w, v \in W$. Let $n \in \text{NOM}$ such that $n \in v$ (recall that all MCSs in the upper Henkin model are named). Then $\Diamond n \in w$ entails (a). To see this, suppose $\Diamond n \in w$, and let $\alpha \in v$. We show that $\Diamond \alpha \in w$. We have $n \land \alpha \in v$, which entails $E(n \land \alpha) \in w$ (by the definition of the $R_E$ relation). Using the (H) axiom, we derive $U(n \rightarrow \alpha) \in w$. Thus we have $U(n \rightarrow \alpha) \land \Diamond n \in w$. But $\vdash_{\text{MLSR}} (U(n \rightarrow \alpha) \land \Diamond n) \rightarrow \Diamond \alpha$, and so $\Diamond \alpha \in w$, as desired.

With this at hand, suppose now that (a) fails. By the above, this means that $\Diamond n \notin w$, where $n \in v$. Thus, we have $\neg \Diamond n \in w$. Now, $n \in \Phi(\mathcal{M}, \varphi, v)$ (using $\vdash_{\text{MLSR}} \langle \langle \varphi \rangle \alpha \rangle \leftrightarrow \operatorname{pre}(\varphi) \land n$). But we also have $\vdash_{\text{MLSR}} \neg \Diamond n \rightarrow \neg \langle \langle \varphi \rangle \Diamond \alpha \rangle$ and, since $\neg \Diamond n \in w$, we conclude $\langle \langle \varphi \rangle \Diamond \alpha \rangle \notin w$. Therefore, $\Diamond n \notin \Phi(\mathcal{M}, \varphi, w)$, and (b) fails.

Note that the same proof goes through for the relation $R_E^\varphi$ (for $E\alpha$ instead of $\Diamond \alpha$).

In the derived Henkin models defined in this way, all points are still named by nominals of the sort introduced earlier, but some nominals may now fail to denote (this explains our earlier change in the hybrid base logic). Intuitively, the models $\mathcal{M} : \varphi$ are meant to be isomorphic to submodels of $\mathcal{M}$ after the sequence of consecutive updates $\varphi$, but the precise sense in which this is true will become clear in the following key property of the above construction, whose proof clarifies the above choice of definitions.

**Lemma 3.13** (Truth Lemma). For all formulas $\psi$, finite sequences $\varphi$ and points $w$,

$$
\mathcal{M} : \varphi, (w, \varphi) \models \psi \text{ if and only if } \psi \in \Phi(\mathcal{M}, \varphi, w)
$$

**Proof.** The proof is by induction on the formulas $\psi$, using various well-known properties of proofs in public announcement logic PAL [van Benthem, 2011]. For convenience, when the context is clear, we write $w$ instead of $\langle w, \varphi \rangle$.

(a) For the equivalence of truth and membership for atoms $p$, it suffices to observe that $p \in \Phi(\mathcal{M}, \varphi, w)$ iff $\langle \langle \varphi \rangle \rangle p \in w$ iff $\operatorname{pre}(\varphi) \land p \in w$, by the recursion axiom for atomic formulas in PAL. The same argument applies to nominals, by using the axiom $\langle \langle \varphi \rangle \rangle n \leftrightarrow \varphi \land n$. 


(b) For negations, we have \( \neg \psi \in \Phi(M, \mathcal{P}, w) \) iff \( \langle \mathcal{P} \rangle \neg \psi \in w \), which, by the PAL recursion axiom for negation, is provably equivalent to \( \text{pre}(\mathcal{P}) \in w \) and \( \neg \langle \mathcal{P} \rangle \psi \in w \).\(^3\) Given that \( \text{pre}(\mathcal{P}) \in w \), we have that \( \langle \mathcal{P} \rangle \neg \psi \in w \) iff \( \langle \mathcal{P} \rangle \psi \notin w \). This is equivalent to \( \psi \notin \Phi(M, \mathcal{P}, w) \), which, by the inductive hypothesis for \( \psi \), holds if and only if \( M : \mathcal{P}, w \vdash \neg \psi \).

(e) The inductive step for disjunctions \( \psi_1 \lor \psi_2 \) is similar to the preceding one, by using the distribution of \( \langle \mathcal{P} \rangle \) modalities over disjunctions repeatedly, as well as the fact that maximally consistent sets split disjunctions.

(d) Next, we consider the case of the basic modality \( \lozenge \psi \).

- If \( M : \mathcal{P}, w \vdash \lozenge \psi \), then, for some \( v \) with \( R^w_\mathcal{P}v \), \( M : \mathcal{P}, v \vdash \psi \). So, by the inductive hypothesis, \( \psi \in \Phi(M, \mathcal{P}, v) \). Hence, in the upper Henkin model \( M \), we have that \( \langle \mathcal{P} \rangle \psi \in v \) and also \( R_\alpha v \). This entails that, in \( M \), \( \langle \mathcal{P} \rangle \psi \in w \). Now, using the fact that the pre-condition \( \text{pre}(\mathcal{P}) \) for the finite sequence \( \varphi \) is in \( w \), and by repeatedly appealing to the PAL recursion axiom for the diamond modality, we get that \( \langle \mathcal{P} \rangle \lozenge \psi \in w \).\(^4\) By the definition of derived Henkin models, it then follows that \( \lozenge \psi \in \Phi(M, \mathcal{P}, w) \).

- Conversely: \( \lozenge \psi \in \Phi(M, \mathcal{P}, w) \) entails that \( \langle \mathcal{P} \rangle \lozenge \psi \in w \) in the upper Henkin model \( M \). Since \( \text{pre}(\mathcal{P}) \in w \), using the converse direction of the PAL recursion axiom, we obtain \( \lozenge \langle \mathcal{P} \rangle \psi \in w \). By the Existence Lemma, there is some \( v \in M \) with \( R_\alpha v \) and \( \langle \mathcal{P} \rangle \psi \in v \). Since \( \vdash_{\text{MLSR}} \langle \mathcal{P} \rangle \psi \leftrightarrow \text{pre}(\mathcal{P}) \lor \langle \mathcal{P} \rangle \psi \), it follows that \( \text{pre}(\mathcal{P}) \in v \). Hence, \( v \) is in the derived Henkin model. So, we have that \( \psi \in \Phi(M, \mathcal{P}, v) \) and, by the inductive hypothesis, we get \( M : \mathcal{P}, v \vdash \psi \). Since \( R_\alpha v \), we have \( R^w_\mathcal{P}v \), and so \( M : \mathcal{P}, w \vdash \lozenge \psi \).

(e) The reasoning for the existential modality \( \exists \psi \) is just like the preceding argument, except that it relies on the reduction axiom \( \langle \mathcal{P} \rangle \exists \psi \leftrightarrow (\varphi \land E(\langle \mathcal{P} \rangle \psi)) \).

(f) The analysis for PAL modalities \( \langle \alpha \rangle \psi \) proceeds as follows.

First note that, by the inductive hypothesis, the truth of \( \alpha \) at any point \( (v, \mathcal{P}) \) is equivalent to \( \alpha \) belonging to \( \Phi(M, \mathcal{P}, v) \). Hence, restricting the model \( M : \mathcal{P} \) to \( (M : \mathcal{P})|\alpha \) in the usual semantic sense yields exactly the derived Henkin model \( M : (\mathcal{P}^\alpha) \).\(^5\) Then, we have the following equivalences:

\[
M : \mathcal{P}, w \vdash \langle \alpha \rangle \psi \quad \text{iff} \quad M : \mathcal{P}, w \vdash \alpha \quad \text{and} \quad (M : \mathcal{P})|\alpha, w \vdash \psi \\
\text{iff} \quad M : (\mathcal{P}^\alpha), w \vdash \psi \\
\text{iff} \quad \psi \in \Phi(M, \mathcal{P}^\alpha, w) \quad \text{(by the inductive hypothesis)} \\
\text{iff} \quad \langle \alpha \rangle \psi \in \Phi(M, \mathcal{P}, w) \quad \text{(since} \quad \langle \mathcal{P}^\alpha \rangle \psi = \langle \mathcal{P} \rangle \langle \alpha \rangle \psi) \]

\(^3\)This claim requires some computation by iterating the relevant PAL recursion axioms. For instance, \( \langle \mathcal{P}_1 \rangle \langle \mathcal{P}_2 \rangle \neg \psi \leftrightarrow \langle \mathcal{P}_1 \rangle (\mathcal{P}_2 \lor \neg \langle \mathcal{P}_2 \rangle \psi) \leftrightarrow (\langle \mathcal{P}_1 \rangle \mathcal{P}_2 \lor \langle \mathcal{P}_1 \rangle \neg \langle \mathcal{P}_2 \rangle \psi) \leftrightarrow (\langle \mathcal{P}_1 \rangle \mathcal{P}_2 \lor \neg \langle \mathcal{P}_1 \rangle \langle \mathcal{P}_2 \rangle \psi) \) is provable. Similar iterated calculations apply to the further cases below.

\(^4\)Here is a concrete calculation with the reduction axiom showing the operator inversion: \( \langle \mathcal{P}_1 \rangle \langle \mathcal{P}_2 \rangle \lozenge \psi \leftrightarrow \langle \mathcal{P}_1 \rangle (\mathcal{P}_2 \lor \lozenge \langle \mathcal{P}_2 \rangle \psi) \leftrightarrow (\langle \mathcal{P}_1 \rangle \mathcal{P}_2 \lor \langle \mathcal{P}_1 \rangle \lozenge \langle \mathcal{P}_2 \rangle \psi) \leftrightarrow \mathcal{P}_1 \lor (\langle \mathcal{P}_1 \rangle \mathcal{P}_2 \lor \langle \mathcal{P}_1 \rangle \langle \mathcal{P}_2 \rangle \psi). \)

\(^5\)Thanks to the inductive hypothesis, \( M : (\mathcal{P}^\alpha) \) contains exactly the states \( w \) for which \( \text{pre}(\mathcal{P}^\alpha) \in w \). But the derived Henkin model \( M : (\mathcal{P}^\alpha) \) is restricted to exactly the states \( w \) such that \( \text{pre}(\mathcal{P}^\alpha) \in w \), which is equivalent to \( \langle \mathcal{P} \rangle \alpha \in w \).
Having analyzed expressive power and axiomatization, we now turn to matters of computational complexity and undecidability.

4 Complexity and undecidability

Having analyzed expressive power and axiomatization, we now turn to matters of computational complexity for the core notions of our system MLSR as defined in Section 2.

(g) Finally, the MLSR deletion modality $(\neg \alpha)\psi$ is analyzed as follows.

- If $M : \varphi, w \models (\neg \alpha)\psi$, then, for some $v \neq w$, $M : \varphi, v \models \alpha$ and $(M : \varphi) - \{v\}, w \models \psi$; equivalently, using a nominal $n$ denoting $v$, $(M : \varphi)|- n, w \models \psi$. Here, by our construction, the model $(M : \varphi)|- n$ equals $M : (\varphi \equiv n)$. Then, by the inductive hypothesis, we have that (i) $\psi \in \Phi(M, \varphi \equiv n, w)$, and also that (ii) $\alpha \in \Phi(M, \varphi, v)$. The former means that, looking in $M : \varphi$, we have $(!n)\psi \in \Phi(M, \varphi, w)$. The latter means that $E(n \land \alpha) \in \Phi(M, \varphi, w)$. To see this, first note that $n \in v$ entails $(!n)n \in v$, and so we have $(!n)(n \land \alpha) \in v$. We also have $\text{pre}(\varphi) \in v$, and using the reduction axiom for $E$ we have

$$\vdash_{\text{MLSR}} (!n)E(n \land \alpha) \leftrightarrow (\text{pre}(\varphi) \land E(!n)(n \land \alpha)) \tag{1}$$

This allows us to conclude $(!n)E(n \land \alpha) \in w$. So, we have $E(n \land \alpha) \land (\neg n)\psi \in \Phi(M, \varphi, w)$ and, by the Mix Axiom, $(\neg \alpha)\psi \in \Phi(M, \varphi, w)$.

- If $(\neg \alpha)\psi \in \Phi(M, \varphi, w)$, then $(!n)(\neg \alpha)\psi \in w$. Since $w$ is mixed, there is some nominal $n$ such that $(!n)n \in w$, $(!n)(n \land \alpha) \in w$, and $(!n)(\neg n)\psi \in w$. Using (1), we get $E(n \land \alpha) \in w$. By the Existence Lemma, this means there is some $v \in M$ with $(!n)(n \land \alpha) \in v$. So, $n \land \alpha \in \Phi(M, \varphi, v)$. By the induction hypothesis, this entails that $M : \varphi, v \models \alpha$.

Since $\text{pre}(\varphi) \land (\neg n)\psi \in w$, using $\vdash_{\text{MLSR}} \text{pre}(\varphi) \rightarrow (\neg n)\psi \leftrightarrow \neg n$, we get $n \notin w$, while $(!n)(n \land \alpha) \in v$ entails $n \in v$, so $w \neq v$.

Lastly, $(!n)(\neg n)\psi \in w$ means $(\neg n)\psi \in \Phi(M, \varphi, w)$. This is equivalent to $\psi \in \Phi(M, \varphi \equiv n, w)$, which, by the induction hypothesis, yields $M : \varphi \equiv n, w \models \psi$. By the induction hypothesis for nominals, $M : (\varphi \equiv n) = (M : \varphi) - \{v\}$, and so we get $(M : \varphi)|- n, w \models \psi$. But note that $(M : \varphi)|- n = (M : \varphi) - \{v\}$, and so we have $(M : \varphi) - \{v\}, w \models \psi$.

Taking these three facts together, we conclude $M : \varphi, w \models (\neg \alpha)\psi$.

This concludes the proof of the Truth Lemma, and completeness follows. □

Remark. One way of understanding the mechanics of this modal completeness proof is doing a parallel standard Henkin-style completeness proof for a first-order language with explicit operations of quantification without replacement and definable relativization.

4 Complexity and undecidability

Having analyzed expressive power and axiomatization, we now turn to matters of computational complexity for the core notions of our system MLSR as defined in Section 2.
4.1 Model checking

We begin by showing that model checking for MLSR is PSPACE-complete. We do so by providing a reduction from the quantified Boolean formula problem (QBF) [Stockmeyer and Meyer, 1973], in the style of Rohde [2005] and Löding and Rohde [2003].

**Theorem 4.1.** Model checking for MLSR is PSPACE-complete.

*Proof.* An upper bound is established as follows. The translation into first-order logic given earlier (Fact 2.5) only has a polynomial size increase, and it is known that model checking for first-order logic is in PSPACE.

The lower bound is demonstrated by a reduction from QBF into model checking for MLSR. Take any QBF formula $\varphi$: that is, a formula of the form

$$Q_1 x_1 \ldots Q_n x_n \bigwedge_{1 \leq i \leq k} C_i,$$

where $Q_j \in \{\exists, \forall\}$, and each $C_i$ is a disjunction of literals $\pm x_j$ (w.l.o.g., we can assume that the quantifiers alternate between $\exists$ and $\forall$). Given such a formula $\varphi$, we construct a finite pointed model $(M_\varphi, s)$ and an MLSR formula $\gamma_\varphi$ such that $\varphi$ is true if and only if $(M_\varphi, s) \models \gamma_\varphi$. The construction will ensure that the model $M_\varphi$ and the formula $\gamma_\varphi$ both have a size that grows linearly in the number of quantifiers and clauses of $\varphi$, which gives the desired reduction from QBF.

To increase intuitive understanding, in what follows the model $(M_\varphi, s)$ is constructed so that the truth of $\varphi$ can be captured by a ‘traveling’ game on the model between two players: Traveler and Demon. The formula $\varphi$ is true if and only if Traveler has a winning strategy in the traveling game on $(M_\varphi, s)$, while the MLSR formula $\gamma_\varphi$ states the existence of a winning strategy for Traveler.

The model $(M_\varphi, s)$ is built out of $n + 1$ vertically concatenated ‘modules’: one initial module for the first quantifier in $\varphi$, one module for each of the remaining $n - 1$ quantifiers, plus one final verification module. Each of these modules is depicted in Figure 2.

The construction of $M_\varphi$ is as follows: starting with the initial module, we concatenate successive $\forall x_i$- and $\exists x_j$-modules corresponding to the order of quantifiers in $\varphi$ (we treat the top nodes labeled by $x_j$ and $\neg x_j$ as the end nodes of the previous module). The *goal points* are those to which the valuation assigns the proposition letter $g$ (as depicted in Figure 2). Once all $n$ quantifier modules have been added, we append the final verification module. For each clause $C_i$, we use a distinct proposition letter $c_i$, which holds at exactly one node in the verification module, called a *clause vertex*. Each clause vertex $c_i$ has an outgoing edge to all and only the duals of literals that make $C_i$ true.

Now, the traveling game proceeds in the following manner. At the beginning, Traveler is positioned at the starting vertex $s$. When Demon plays, she deletes a node in the graph. When Traveler plays, she can travel along one of the remaining edges to an adjacent vertex. Traveler wins if she manages to reach a goal point, marked with the proposition letter $g$. Demon wins otherwise.

More in detail, if $\varphi$ starts with $\exists$, Traveler goes first. If $\varphi$ starts with $\forall$, Demon goes first: in the first move, she can only delete a vertex marked by $p_1$—that is, she can only delete a point adjacent to the starting vertex. From then on, Traveler and Demon alternate their
(a) Initial module.

(b) \(\exists x_j\)-module.

(c) \(\forall x_j\)-module.

(d) Final verification module.

Figure 2: The shape of the initial module (a) does not depend on which quantifier \(\varphi\) begins with. In (b), (c) and (d), the top nodes labeled by \(x_j\) and \(\neg x_j\) are the end nodes of the previous module. In (d), each clause vertex \(c_i\) has an outgoing edge to a vertex labeled by a literal \(\pm x_j\) exactly if the dual literal \(\mp x_j\) makes clause \(C_i\) true.

turns, with turns being either traveling one edge further or deleting one node, respectively, where the Demon’s second move at each \(\forall x_j\)-module is restricted to nodes marked with \(p_j\) (see Figure 2). This continues in this manner until Traveler reaches a node that sees a clause node. At this point, Demon has \(k - 1\) moves, which she must use to delete all but one clause node. Then, we allow Traveler two successive moves (once Demon has restricted her choices to one clause node). Then, Demon and Traveler once again alternate single moves until the game is resolved. See Figure 3 for an example.

The game adequately captures the truth of \(\varphi\):

**Observation 4.2.** Traveler has a winning strategy for the game on \((M_\varphi, s)\) if and only if the initial QBF formula \(\varphi\) is true.
\[ \varphi = \forall x_1 \exists x_2 \forall x_3 (C_1 \land C_2 \land C_3) \]

where \( C_1 = \neg x_1 \lor x_2 \), \( C_2 = \neg x_1 \lor x_2 \lor \neg x_3 \), and \( C_3 = x_1 \lor x_2 \lor x_3 \)

Figure 3: An example. The proposition letter \( g \) marks the goal points. Each \( \forall \)-module forces Traveler to the literal \( \pm x_j \) point chosen by Demon, while each \( \exists \)-module leaves the choice to Traveler. The letters \( p_1, \ldots, p_k \) are level markers that restrict Demon’s moves. In the final module, Demon forces Traveler into some clause. For each literal \( \pm x_j \) that makes such clause true, Traveler can then go to the node labeled by dual literal \( \mp x_j \) above.
Proof. Each travel path to the verification module yields a valuation. Say that a truth value is selected for $x_i$ if Traveler’s path passes through the node labeled $x_i$. At $\forall x_i$ modules, Demon selects a truth value for $x_i$. At $\exists x_j$ modules, Traveler selects a truth value for $x_j$. Once Traveler reaches a clause node, an assignment has been chosen for all variables. The design of the above modules guarantees that, at that stage, the deleted goal points are all and only those seen by visited vertices, and also, all unvisited $\pm x_i$ vertices still have two adjacent goal points.

If $\varphi$ is true, then the assignment chosen in this manner (with Demon controlling $\forall$ and Traveler controlling $\exists$) makes all (disjunctive) clauses true. So, for every clause $C_i$, there is some visited $\pm x_j$ node, for some $\pm x_j$ that entails $C_i$. By design of the final clause module, no matter which clause Traveler is at, there is some unvisited vertex labeled $\mp x_j$ accessible from this clause vertex. Traveler can then travel to this vertex, where she sees two goal points. Demon can remove at most one of them at her next move, and Traveler therefore wins. Conversely, if $\varphi$ is false, the final assignment makes at least one clause false. Demon forces Traveler to the corresponding clause vertex: because the current assignment makes the disjunctive clause false, all the accessible $\pm x_j$ vertices have already been visited and, thus, do not see any goal points. Demon therefore wins.

Lastly, to conclude the proof of Theorem 4.1, we make sure that MLSR can express the existence of a winning strategy for Traveler. When $\varphi$ starts with $\forall$ and has $n$ quantifiers and $k$ clauses, the general form of the corresponding MLSR formula $\gamma_\varphi$ is

$$((\neg \alpha_j) \otimes^{f(n)} [\delta]^{k-1} \otimes^2 [\neg \top] \otimes g)$$

Here, $f(n)$ is a function counting the total number of rounds played in the game up to the final module: $f$ is linear in $n$ (it is in fact easy to see that $f(n) \leq 3n$). The symbol $\alpha_j$ denotes $p_j$ whenever Traveler sees a $p_j$-point in the corresponding $\forall x_j$-module; it stands for $\top$ otherwise. The formula $\delta$, on the other hand, is a Boolean combination of $c_i$’s expressing that exactly one of the clauses is true. The repeated modalities capture exactly the structure of the game and the restrictions on the players’ moves. The game goes on for $f(n)$ rounds until the penultimate stage is reached. The $\neg p_j$ modalities force Demon to remove only $p_j$-points during the middle round played on a $\forall x_j$-module. Then, $[\delta]^{k-1}$ quantifies over all ways in which Demon can remove $k - 1$ clauses (all but one). The formula $\gamma_\varphi$ expresses that Traveler can ensure that such a sequence of moves results in reaching a goal point, and thus holds exactly if Traveler has a winning strategy: equivalently, it holds if and only if the initial QBF formula $\varphi$ is true.

Note on game equivalence. While not strictly necessary for the proof, the above traveling game with point removal over this structure is appealing, and it suggests links with the graph games we started from [van Benthem and Liu, 2018]. Moreover, the traveling game is virtually identical to the standard logical evaluation game for the given quantified Boolean formula. Can this game equivalence be made precise?

4.2 Satisfiability

Next, we show that, despite the recursive axiomatizability shown in Section 3, stepwise removal has a complex theory. The satisfiability problem for the logic MLSR is undecidable,
which we establish by a reduction from the tiling problem, a standard technique in modal logic (cf. [Blackburn et al., 2011; Marx, 2006], to which we refer for details).

**Theorem 4.3.** The satisfiability problem for MLSR with two binary accessibility relations $R_u$ and $R_r$ is undecidable.

**Proof.** Let $\mathcal{T} = \{T_1, \ldots, T_n\}$ be a finite set of tile types. Given a tile type $T_i$, $u(T_i), r(T_i), d(T_i)$ and $l(T_i)$ will represent the colors of the upper, right, lower and left edges of $T_i$, respectively. For each tile type $T_i$, we fix a proposition letter $t_i$ that is going to encode $T_i$.

We will now define an MLSR formula $\varphi_\mathcal{T}$ such that the following holds:

$\varphi_\mathcal{T}$ is satisfiable if and only if $\mathcal{T}$ tiles the discrete quadrant $\mathbb{N} \times \mathbb{N}$.

The formula $\varphi_\mathcal{T}$ is the conjunction of the following MLSR formulas. The first three describe the relational structure of a grid, the last three encode the behavior of a tiling of the grid:

- **(Succ)** $U(\Diamond_u T \land \Diamond_r T)$
- **(Func)** $U(\neg T)(\Diamond_u \perp \land \Diamond_r T)$
- **(Conf)** $U(\neg T)(\Diamond_r \Diamond_u \perp \land \Diamond_u \Diamond_r \perp)$
- **(Unique)** $U\left(\bigvee_{1 \leq i \leq n} t_i \land \bigwedge_{1 \leq i < j \leq n} (t_i \rightarrow \neg t_j)\right)$
- **(Vert)** $U \bigwedge_{1 \leq i \leq n} \left(t_i \rightarrow \Diamond_u \bigvee_{1 \leq j \leq n, u(T_i) = d(T_j)} t_j\right)$
- **(Horiz)** $U \bigwedge_{1 \leq i \leq n} \left(t_i \rightarrow \Diamond_r \bigvee_{1 \leq j \leq n, r(T_i) = l(T_j)} t_j\right)$

$\Rightarrow$ It is easy to see that any tiling of $\mathbb{N} \times \mathbb{N}$ induces a model for $\varphi_\mathcal{T}$.

$\Rightarrow$ For the other direction, suppose that $\mathcal{M}, w \models \varphi_\mathcal{T}$, for some LSR-model $\mathcal{M} = (W, R_u, R_r, V)$ and $w \in W$. The formulas (Succ) and (Func) ensure that the relations $R_u$ and $R_r$ are functions, and that for every point $x$, that $R_u[x] \neq R_r[x]$ (the $R_u$ and $R_r$-images of $x$ are different). The formula (Conf) then guarantees that the functions commute: $R_u \circ R_r = R_r \circ R_u$. This ensures the existence of an embedding $f : \mathbb{N}^2 \rightarrow W$ that preserves the structure of vertical and horizontal successors: that is, for all $(n, m) \in \mathbb{N}^2$, we have $R_u(f(n, m), f(n, m + 1))$ and $R_r(f(n, m), f(n + 1, m))$. Now, tile the point $(n, m)$ in $\mathbb{N}^2$ with tile $T_i$ exactly if $\mathcal{M}, f(n, m) \models t_i$. This gives a tiling of the discrete quadrant of the plane. □

The two standard modalities used in this proof can be reduced to one using standard techniques [Kracht and Wolter, 1999], but we forego details here because of the syntactic cost involved in writing the formulas.

The above undecidability argument will also work with languages less expressive than MLSR. In particular, one can replace the universal modality by an extra standard modality that can survey the domain by employing the well-known ‘spypoint technique’ from hybrid logic. A detailed syntactic construction of this sort for modal logics of graph games can be found in [Zaffora Blando et al., 2019].

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Having determined the complexity of model checking and satisfiability, one task would remain, concerning definability and expressive power. However, we leave this open here.

**Open problem.** *What is the complexity of testing for SR-bisimulation?*

Given any two finite models $M, N$, it is easy to find an EXPSPACE upper bound. One considers the space of all models arising from $M, N$ by deleting finite sequences of different points, and then tests for ordinary modal bisimulation over this space with respect to MLSR, now viewed as a standard bimodal language. But, just as with standard modal bisimulation [Kanellakis and Smolka, 1983], one can probably do better.

## 5 Stepwise removal over first-order fragments

Having established the complexity of adding quantification without replacement to the basic modal language, we can also consider other fragments of first-order logic. Perhaps the simplest case is adding the removal modality $\langle - \varphi \rangle \psi$ to monadic first-order logic $MFOL$. As it turns out, this yields exactly the formulas in $MFOL^{=}$: that is, all formulas with one free variable $x$ in monadic first-order logic with identity.

**Theorem 5.1.** $\text{MLSR}(MFOL) = MFOL^{=}$.

**Proof.** Fix finitely many unary predicates $P_1, \ldots, P_k$. We define standard normal forms for the whole language $MFOL$. Local state descriptions $sd$ are conjunctions of $\pm P_i$ with $1 \leq i \leq k$. There are $2^k$ of these, and they can be applied to arbitrary variables. Global state descriptions $SD$ of depth $N$ are then conjunctions $\bigwedge_j SD_j$ where, for each local state description $sd_j$, $SD_j$ is either the statement that exactly $m_j$ objects satisfy $sd_j$, where we have $m_j < N$, or the statement that at least $N$ objects satisfy $sd_j$.

**Claim.** Each formula in $MFOL$ of quantifier depth $N$ and $m$ free variables $x_1, \ldots, x_m$ is equivalent to a disjunction of conjunctions $NF$, each consisting of (a) local state descriptions for each of the variables $x_i$ plus a complete set of equalities and inequalities for all pairs of variables from $x_1, \ldots, x_m$, plus (b) a global state description that is consistent with (a) in an obvious syntactic sense.

This can be proved by induction on formulas via a syntactic argument.\(^8\)

**Claim.** $MFOL$ is closed under the modality $\langle - \varphi \rangle \psi$.

**Proof.** Using the disjunction axioms for existential modalities and $\langle - \varphi \rangle \psi$ stated in Section 3, in proving closure, one can restrict attention to conjunctive forms $NF$ and special removal modalities $\langle -sd \land SD \rangle NF$. Closure can be shown here by a simple argument, driven by the following two key facts:

- the equivalence $\langle -sd \land SD \rangle NF \leftrightarrow (SD \land \langle -sd \rangle NF)$ is valid.\(^9\)

\(^8\)Alternatively, $NF$ describes a model $M$ in such a way that, for any model $N$ that satisfies $NF$, Duplicator has a winning strategy in the Ehrenfeucht game over $N$ rounds between $M$ and $N$ starting with the partial isomorphism between the objects on both sides satisfying the atomic diagram (a).

\(^9\)Here is a general useful principle that is easy to state in first-order syntax. When we take out a point satisfying $\varphi(x) \land \psi$, where $x$ does not occur free in $\psi$, then we can just put $\psi$ outside in a conjunction.
the formula $\langle -sd_i \rangle (sd' \land SD)$ is equivalent to $sd' \land SD[i := i + 1]$, where $SD[i := i + 1]$ replaces the quantification in the $i$-th conjunct of $SD$ by a quantifier stating the existence of one more point satisfying the relevant local state description.

Arguments like this are available for other languages that admit of simple normal forms of modal depth 1. Here is one obvious question.

**Open problem.** What fragment of first-order logic results from adding the dynamic operators of MSLR to the language of modal $S5$?

We have some initial results, but the combinatorics get considerably more complex, since the logic can now also distinguish between different equivalence classes in $S5$ models.

The general question suggested by the specific case analyzed in Theorem 5.1 is the following: what is the boost in expressive power when we close fragments of first-order logic under various model-changing modalities?

6 Conclusion and further directions

The logic of stepwise removal of objects lies in between modal logics of definable model change and logics for graph games with arbitrary moves, and it may well be the most intuitive example of a modal system that crosses the line from decidable to undecidable.\(^{10}\)

We have established its main properties, proving a bisimulation characterization theorem and other results on expressive power, a completeness theorem, and two basic complexity results. Most of the techniques that we used are well-known, others less so, and we also introduced a new technique for proving completeness. The resulting style of thinking can be applied to a wide range of modal systems of this sort.

Among the issues still to be addressed is the complexity problem for SR-bisimulation (Section 4), as well as the expressive closure problem for $S5$. Moreover, all of our questions return for some obvious extensions and variations.

**Simultaneous versions of MSLR.** It is natural to add a modality for removing a fixed finite number of points, either in a conjunctive unary version $\langle -(\varphi_1, \ldots, \varphi_k) \rangle \psi$ or with truly polyadic operators $\langle \neg \varphi \rangle \psi$, where the formulas $\varphi$ can be evaluated in a tuple of indices. These modalities seem undefinable as iterations of our unary $\langle -\varphi \rangle \psi$. Even so, we conjecture that all of our results go through.

Another immediate question concerns other extended modal logics.

**Connections with hybrid logic.** MSLR seems closely related to hybrid modal formalisms such as memory logics [Areces et al., 2008]. For instance, is there an inverse translation to the one into $H(\@, \downarrow)$ given in Section 2?

Then, there is the question of the scope of our methods.

**Axiomatizing logics of graph games.** It is a long-standing open problem how to axiomatize the validities of sabotage-style modal logics and related ones [Aucher et al., 2018].

\(^{10}\)Another contender is the ‘modal fact change logic’ of Thompson [2019].
Does our axiomatization technique for LSR employing added dynamic epistemic modalities work for these logics, as well?

Next, returning to the issue of undecidability, a few questions arise naturally.

**Other sources of undecidability.** In addition to the undecidability induced by stepwise removal, there is the undecidability induced by local link-cutting or local definable point removal, taking place only at the current point of evaluation [Li, 2018]. Both modifications of dynamic-epistemic logics block the usual recursion axioms, both allow for tiling encodings, but the connection remains to be clarified.

But there are also other perspectives on complexity that we have found.

**Lowering the complexity of MLSR.** Can MLSR be shifted back into the decidable modal fold? For many logical systems, one can lower the complexity by a Henkin-style change in the semantics [Andréka et al., 2016]. One could restrict the removal of points to those that are accessible from the current point in some global relation $Axy$ and, if this does not suffice for decidability, one might use further guarding, so that the earlier first-order translations of LSR formulas end up inside guarded, or loosely guarded, fragments of FOL.

Yet, moves like this make most sense when connected to a principled view of computation. We believe that modal logics like MLSR, but also hybrid memory logics or related systems, offer an interesting alternative take on the sources of computational complexity. In the usual automata hierarchy, Turing machine power arises when we have an active memory that can be rewritten. In our logics, however, a simple device that merely stores the set of deleted or visited points suffices. The reason must be the interplay of memory and expressiveness of the language for constructing models around that memory, suggesting a sort of descriptive complexity theory complementary to that of Immerman [1999].

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**References**


