Two series of formalized interpretability principles for weak systems of arithmetic

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Abstract

The provability logic of a theory $T$ captures the structural behavior of formalized provability in $T$ as provable in $T$ itself. Like provability, one can formalize the notion of relative interpretability giving rise to interpretability logics. Where provability logics are the same for all moderately sound theories of some minimal strength, interpretability logics do show variations.

The logic $\text{IL}(\text{All})$ is defined as the collection of modal principles that are provable in any moderately sound theory of some minimal strength. In this paper we raise the previously known lower bound of $\text{IL}(\text{All})$ by exhibiting two series of principles which are shown to be provable in any such theory. Moreover, we compute the collection of frame conditions for both series.

1 Introduction

Relative interpretations in the sense of Tarski, Mostowski and Robinson \cite{10} are widely used in mathematics and in mathematical logic to interpret one theory into another. Roughly speaking, such an interpretation between two theories is a translation from the language of one theory to the language of the other so that the translation preserves logical structure and theoremhood.

We shall write $U \triangleright V$ to denote that a theory $U$ interprets a theory $V$. Once we know that $U \triangleright V$, this provides us much information; for example the consistency of $U$ implies the consistency of $V$ and also, various definability results carry over from the one theory to the other. Famous examples of interpretations are abundant: the theory of the natural numbers into the theory of the integers, set theory plus the continuum hypothesis into ordinary set theory, non-Euclidean geometry into Euclidean geometry, etc.

Interpretability, being a syntactical notion, allows for formalization very much as one can formalize the notion of provability. As such, we can consider interpretability logics which actually extend the well-know provability logic $\text{GL}$ of Gödel Löb. We shall see that the interpretability logic of a theory is the collection of all structural properties of interpretability that it can prove.
Where all modestly correct theories of some minimal strength—let us call them *reasonable theories* in this paper—have the same provability logic $\text{GL}$, the situation is different in the case of interpretability and different theories have different logics. It is an open question to determine the logic of interpretability principles being provable in any reasonable theory. This paper reports on substantial progress on this open question by increasing the previously known lower bound.

## 2 Preliminaries

Let $U$ and $V$ denote theories with languages $\mathcal{L}_U$ and $\mathcal{L}_V$ respectively. A relative interpretation $j$ from $V$ into a theory $U$—we will write $j : U \sqsupset V$—is a pair $\langle \delta(x), t \rangle$ where $\delta(x)$ is a formula of $\mathcal{L}_U$ that specifies the domain in which $V$ will be interpreted and $t$ is a translation, mapping symbols of $\mathcal{L}_V$ to formulas of $\mathcal{L}_U$ providing a definition in $U$ of these symbols.

The translation $t$ is extended to a translation $j$ of formulas in the usual way by having $j$ commute with the connectives and relativize the quantifiers to the domain specifier $\delta(x)$ as follows: $(\forall x \varphi(x))^j := \forall x (\delta(x) \rightarrow \varphi^j(x))$. We will not go too much into details but the main point is that interpretations are primarily syntactical notions—especially for finite languages—and as such allow for an arithmetization/formalization very much as formal proofs do.

### 2.1 Arithmetic

In order to formalize the notion of interpretability within some base theory $T$ one needs to require some minimal strength conditions on $T$. In particular, we shall require that $T$ can speak of numbers where to code syntax and without loss of generality we shall assume that $T$ contains the language of arithmetic $\{+, \times, S, 0, 1, <, \leq\}$.

We will need that the main properties of the basic syntactical operations like substitution are provable within $T$. For reasonable coding protocols this implies that we need to require the totality of a function of growth-rate $\omega_1(x) := x \mapsto 2^{2|\log_2 x|}$ where $|\cdot|$ denotes the integral part of the binary logarithm of $x$.

Further, to perform basic arguments we need a minimal amount of induction and actually a surprisingly little amount of induction suffices. Buss’s theory $S^1_2$ has just the needed amount of induction and proves the totality of $\omega_1$ and this shall be our base theory (formulated in the standard language of arithmetic).

Alternatively, we could have taken as base theory $\text{I} \Delta_0 + \Omega_1$ which consists of Robinson’s arithmetic $\text{Q}$ together with induction for bounded formulas with parameters and the axiom $\Omega_1$ stating that the graph of $\omega_1$ defines a total function.

We refer the reader to [5] and [2] for further details.

A sharply bounded quantifier is one of the form $\forall x<|y|$ where $|y|$ denotes the integer value of the binary logarithm of $y$. The class $\Delta^b_0$ contains exactly the formulas where each quantifier is sharply bounded. The class $\Sigma^b_1$ arises by
allowing bounded existential quantifiers and sharply bounded universal quantifiers to occur over \( \Delta_0^b \) formulas. By \( \exists \Sigma^b_1 \) we denote those formulas that arise by allowing a single unbounded existential quantifier over a \( \Sigma^b_1 \) formula. The complexity classes \( \Pi_n, \Sigma_n \) and \( \Delta_n \) refer to the usual quantifier alternations hierarchies in the standard language of arithmetic.

In this paper we shall only be concerned with first order-theories in the language of arithmetic with a poly-time recognizable set of axioms extending \( S^1_2 \) and shall often refrain from repeating (some of) these conditions. We shall write \( \square_T \phi \) as the \( \exists \Sigma^b_1 \) formalization of \( \phi \) being provable in the theory \( T \) and refrain from distinguishing formulas from their Gödel numbers or even the numerals thereof. It is well known that we can express provable \( \Sigma^1_1 \) completeness using formalized provability.

**Lemma 2.1.** For any theory \( T \) extending \( S^1_2 \) we have that
\[
T \vdash \forall \alpha \, \square_T \alpha \rightarrow \square_T \square_T \alpha.
\]

We will use \( U \triangleright V \) to denote the formalization of “the theory \( V \) is interpretable in the theory \( U \)”. If we abbreviate the existential quantifier over numbers that code a pair \( \langle \delta(x), t \rangle \) defining an interpretation by \( \exists \text{int}_j \) we can write
\[
U \triangleright V := \exists \text{int}_j \forall \psi (\square_V \psi \rightarrow \square_U \psi^j), \quad (1)
\]

An interpretation \( j : U \triangleright V \) can be used as a uniform way to obtain a model of \( V \) inside any model of \( U \). If \( U \) satisfies full induction, then we see that actually the defined model of \( V \) is an end extension of the model of \( U \): we define \( f(0) := 0 \) and \( f(x + 1) := f(x) + 1 \) and by induction see that \( \forall x \exists y f(x) = y \). As such, we see that any \( \Sigma_1 \) consequence of \( U \) must necessarily also hold in \( V \). Since \( \square_T \varphi \) is a \( \Sigma_1 \) formula, the insight on end extensions is reflected in what is called *Montagna’s principle* \([\text{U} \triangleright V \rightarrow (\text{U} \cup \{ \square_T \varphi \}) \triangleright (\text{V} \cup \{ \square_T \varphi \})]\).

(2)

In case \( U \) does not have full induction, we can still define the graph \( F(x, y) \) of the function \( f \) from above, but we can no longer prove that the function is total. However, we can prove that \( \exists y \, F(x, y) \) is *progressive*, that is, we can prove
\[
\exists y \, F(0, y) \land \forall x \, (\exists y \, F(x, y) \rightarrow \exists y, \, F(x + 1, y)).
\]

In particular, the formula \( \exists y \, F(x, y) \) defines an initial segment within \( U \). A common trick in weak arithmetics is to use this initial segment as our natural numbers instead of applying induction (which is not necessarily available). By Solovay’s techniques on shortening initial segments we may assume that they obey certain closure properties giving rise to what is called a *definable cut*.

A formula \( J \) is called a \( T \)-cut whenever \( T \) proves all of
1. \( J(0) \land \forall x \, (J(x) \rightarrow J(x + 1)) \);
2. \( \forall x \, (J(x) \land J(y) \rightarrow J(x + y) \land J(xy) \land J(\omega_1(x))) \);
3. \( J(x) \land y < x \rightarrow J(y) \).  

Let \( \text{Cut}(J) \) denote the conjunction of these three requirements. Sometimes we want to quantify over cuts within \( T \) so that these cuts can then of course be non-standard. We shall use \( \forall \text{Cut} J \psi \) and \( \exists \text{Cut} J \psi \) to denote \( \forall J ( \Box T \text{Cut}(J) \rightarrow \psi) \) and \( \exists J ( \Box T \text{Cut}(J) \land \psi) \) respectively. Here the dot notation in \( \Box T \text{Cut}(J) \) is the standard way to abbreviate the formula with one free variable \( J \) stating that the formula \( \text{Cut}(J) \) is provable in \( T \). We will freely use the dot notation throughout the remainder of this paper. Sometimes we shall write \( x \in J \) instead of \( J(x) \).

For \( J \) a cut, let \( \psi^J \) denote the formula where all unrestricted quantifiers are now required to range over values in \( J \). That is, \( (\forall x \phi)^J := \forall x (J(x) \rightarrow \phi^J) \), \( (\exists x \phi)^J := \exists x (J(x) \land \phi^J) \). Moreover, that is the only thing that is done by this translation so that for example \( (\phi \land \psi)^J := \phi^J \land \psi^J \) etc. Instead of writing \( (\Box T \phi)^J \) we shall simply write \( \Box J T \phi \). We note that if \( \psi(J) \in \exists \Sigma^b_1 \), then \( \exists \text{Cut} J \psi(J) \) is again provably equivalent to an \( \exists \Sigma^b_1 \) formula.

Let us get back to the role of induction in Montagna’s principle. If \( j : U \triangleleft V \) and \( U \) does not prove full induction, then \( j \) will not define an end extension of any model of \( U \). However, it is easy to see that \( j \) does define, using the progressive formula \( \exists y F(x, y) \), a definable cut in \( U \) on which \( f \) is an isomorphism. This is reflected in a weakening of Montagna’s principle also referred to as Pudlák’s principle.

**Lemma 2.2.** Let \( T \) be a theory containing \( S^1_2 \) and let \( U \) and \( V \) be theories.

\[
T \vdash U \triangleleft V \rightarrow \exists \text{Cut} J \forall \psi \in \Delta_0 \left( U \cup \{ (\exists x \psi)^J \} \triangleleft V \cup \{ \exists x \psi \} \right).
\] (3)

### 2.2 The interpretability logic of a theory

Interpretability logics are designed to capture structural behavior of formalized interpretability just as provability logic captures the structural behavior of formalized provability. To this end we consider a propositional modal language with a unary modal operator \( \Box \) to model formalized provability and a binary modal operator \( \triangleright \) to model formalized interpretability of sentential extensions of some base theory. Let us make this more precise.

Let us fix an arithmetical theory \( T \); By \( * \) we will denote a realization, that is, any mapping from the set of propositional variables to sentences of \( T \). The map \( * \) is extended to the set of all modal formulas of interpretability logics as follows

\[
\begin{align*}
(\neg A)^* & := \neg A^* \\
(A \land B)^* & := A^* \land B^* \\
(\Box A)^* & := \Box T A^* \\
(A \triangleright B)^* & := (T + A^*) \triangleright (T + B^*).
\end{align*}
\]

We can now define the interpretability logic of a theory as those modal principles which are provable under any realization. With some liberal notation this is captured in the following.
Definition 2.3. Let $T$ be a theory containing $S_1^2$. We define the interpretability logic of $T$ as

$$\text{IL}(T) := \{ A | \forall^* T \vdash A^* \}.$$  

Further, we define the interpretability logic of all arithmetical theories extending $S_1^2$ by

$$\text{IL}(\text{All}) := \{ A | \forall T \forall^* T \vdash A^* \}.$$  

As a direct corollary to (2) –Montagna’s principle– we can conclude that

$$(A \triangleright B) \rightarrow ((A \land \square C) \triangleright (B \land \square C)) \in \text{IL}(T)$$

whenever $T$ proves full induction. However, there is no direct reflection of Pudlák’s principle on the level of interpretability logics since Pudlák’s principle would translate to

$$(A \triangleright B) \rightarrow ((A \land \square J C) \triangleright (B \land \square C)) \in \text{IL}(T)$$

for the particular cut $J$ corresponding to $j : A \triangleright B$ and this cannot be expressed in our modal language. In a sense, $\square J C$ corresponds to finding a small witness of the provability of $C$. As we shall see, there are various occasions where we can conclude that such small witnesses exist. The two main ingredients in obtaining such small witnesses are expressed by the following lemmas.

Lemma 2.4 (Outside big, inside small). For $T,U$ any theories extending $S_1^2$, we have that

$$T \vdash \forall \text{Cut} J \forall x \square U (\dot{x} \in \dot{J}).$$

Proof. Given $J$ and given $x$, not necessarily in $J$, we can construct a proof-object to the extent that $x \in J$ in the obvious way. First conclude $J(0)$ which holds since $J$ is a cut. Next, conclude that $J(1)$ from the progressiveness of $J$ and $J(0)$ and so all the way to $J(x)$. This proof object is not much bigger than $x$ itself. However it requires the totality of exponentiation. If this is not provable in $T$, the proof can be generalized by switching to dyadic numerals and we refer to e.g. [2, 7] for details.

Lemma 2.5 (Formalized Henkin construction). For theories $T,U$ and $V$ all extending $S_1^2$ we have

$$T \vdash \forall \text{Cut} J (U \cup \{ \text{Con}^I(V) \} \triangleright V).$$

Proof. (Sketch) The theory $T$ can verify that the usual Henkin construction can be formalized in $U$ without many problems where $J$ plays the role of the natural numbers. Instead of applying induction to obtain a maximal consistent set $M_V$ as a consistent branch of infinite length in Lindenbaum’s lemma, we can now only conclude that the length of the branch is within some cut $I$ which is a shortening of $J$ thereby yielding a set $M_V^I$ which is contradiction-free on $I$.

The set $M_V^I$ can be used to obtain a term model and we define an interpretation $j : (U \cup \{ \text{Con}^I(V) \} \triangleright V)$ from the term model as usual so that provably
\[ \phi^j \leftrightarrow (\phi \in M^j_{\psi}) \]. Note that since the interpretation of identity can be any equivalence relation, there is no need to move to equivalence classes in the construction of our term model. By construction we have \( \Box_U \forall \phi (\text{Con}^j(V) \land \Box^j_{\psi} \phi \rightarrow \phi^j) \).

By the outside big, inside small principle and the formalized deduction theorem we now conclude that

\[ \forall \phi \ (\Box_V \forall \varphi ightarrow \Box_{U \cup \{\text{Con}^j(V)\}} \forall \varphi) \]

which, by (1) is nothing but \((U \cup \{\text{Con}^j(V)\}) \rhd V\). We refer to [13] where one can see that the necessary induction for this argument is available in \( S^1_2 \). \( \square \)

Using these lemmas we can infer in various occasions the existence of small witnesses to provability.

**Lemma 2.6.** For any theory \( T \) we have \( T \vdash \neg(A \rhd \neg C) \rightarrow \forall \text{Cut} K \Diamond (A \land \Box^K C) \).

**Proof.** Reason in arbitrary \( T \) by contraposition and apply the Henkin construction on a cut. \( \square \)

As a corollary to this lemma, we see that \((A \rhd B) \rightarrow (\neg(A \rhd \neg C) \rhd (B \land \Box C)) \in \text{IL}(T)\) for any \( T \) extending \( S^1_2 \). It is an open problem to classify the modal principles that hold in any theory extending \( S^1_2 \). This paper raises the previously known lower bound.

We formulate some other direct corollaries of the outside-big inside-small principle in the following useful lemma.

**Lemma 2.7.** Let \( T \) be any theory containing \( S^1_2 \). We have that

1. \( T \vdash \forall A (\Box A \rightarrow \forall \text{Cut} K \Box \Box^K A) \);
2. \( T \vdash \sigma \rightarrow \forall \text{Cut} K \Box \Box^K \sigma \) for any formula \( \sigma \) in \( \exists \Sigma^b \);
3. \( T \vdash \forall C \in \exists \Sigma^b \forall \text{Cut} J (\exists x C \rightarrow \Box \exists x \in \dot{J} C) \).

One ingredient in proving interpretability principles arithmetically sound, is to find small witnesses. Another ingredient tells us how we can keep these witnesses small. A simple generalization of Pudlák’s lemma which was first proved in [6] and tells us how to do so.

**Lemma 2.8.** If \( j : \alpha \rhd \beta \) then, for every cut \( I \) there exists a definable cut \( J \) such that for every \( \gamma \) we have that

\[ T \vdash \forall \text{Cut} J \exists \text{Cut} J \exists j \left( j : (\alpha \land \Box^j \gamma) \rhd (\beta \land \Box^j \gamma) \right). \]
2.3 Modal interpretability logics

When working in interpretability logic, we shall adopt a reading convention that will allow us to omit many brackets. Thus, we say that the strongest binding ‘connectives’ are ¬, □ and ♦ which all bind equally strong. Next come ∧ and ∨, followed by ⊲ and the weakest connective is →. Thus, for example, $A \triangleright B \rightarrow A \land □ C \triangleright B \land □ C$ will be short for $(A \triangleright B) \rightarrow ((A \land □ C) \triangleright (B \land □ C))$.

If we do not disambiguate a formula of nested conditionals (→ or ⊲), then this should be read as a conjunction. For example, $A \triangleright B \triangleright C$ should be read as $(A \triangleright B) \land (B \triangleright C)$ and likewise for implications.

We first define the core logic **IL** which shall be present in any other interpretability logic. As before, we work in a propositional signature where apart from the classical connectives we have a unary modal operator □ and a binary modal operator ⊲.

**Definition 2.9 (IL).** The logic **IL** contains apart from all propositional logical tautologies, all instantiations of the following axiom schemes.

- **L1** $\square(A \rightarrow B) \rightarrow (\square A \rightarrow \square B)$
- **L2** $\square A \rightarrow \square □ A$
- **L3** $\square(\square A \rightarrow A) \rightarrow □ A$
- **J1** $\square(A \rightarrow B) \rightarrow A \triangleright B$
- **J2** $(A \triangleright B) \land (B \triangleright C) \rightarrow A \triangleright C$
- **J3** $(A \triangleright C) \land (B \triangleright C) \rightarrow A \lor B \triangleright C$
- **J4** $A \triangleright B \rightarrow (\Diamond A \rightarrow \Diamond B)$
- **J5** $\Diamond A \triangleright A$

The rules of the logic are Modus Ponens (from $A \rightarrow B$ and $A$, conclude $B$) and Necessitation (from $A$ conclude $\square A$).

It is not hard to see that **IL** ⊆ **IL}(All). By **ILM** we denote the logic that arises by adding Montagna’s axiom scheme

$$M : \quad A \triangleright B \rightarrow A \land □ C \triangleright B \land □ C$$

to **IL**. It follows from our earlier observations that **ILM** ⊆ **IL}(T) and the other inclusion can be proven too.

**Theorem 2.10** (Berarducci [2], Shavrukov [3]). *If T proves full induction, then **IL}(T) = **ILM*.
The logic **ILP** arises by adding the axiom scheme

\[ P : A \triangleright B \rightarrow \Box(A \triangleright B) \]

to the basic logic **IL**. If \( T \) is finitely axiomatizable it is easy to see that \( P \) is provably equivalent to a \( \Sigma_1 \) formula so that by provable \( \Sigma_1 \) completeness we see that \( \text{ILP} \subseteq \text{IL}(T) \) for any finitely axiomatized theory \( T \) that proves that exponentiation is a total function. If \( T \) can moreover prove the totality of superexponentiation \( \text{supexp} \) then the inclusion can be reversed too. Here, \( \text{supexp}(x) \) is defined as \( x \mapsto 2^x \) with \( 2^0 := n \) and \( 2^{n+1} := 2^{2^n} \).

**Theorem 2.11** (Visser [12]). If \( T \) is finitely axiomatizable and proves the totality of \( \text{supexp} \), then \( \text{IL}(T) = \text{ILP} \).

It follows that \( \text{IL} \subseteq \text{IL}(\text{All}) \subseteq (\text{ILP} \cap \text{ILM}) \). In this paper we shall focus on these bounds.

### 2.4 Relational semantics

We can equip interpretability logics with a natural relational semantics often referred to as Veltman semantics.

**Definition 2.12.** A Veltman frame is a triple \( \langle W, R, \{S_x\}_{x \in W} \rangle \) where \( W \) is a non-empty set of possible worlds, \( R \) a binary relation on \( W \) so that \( R^{-1} \) is transitive and well-founded. The \( \{S_x\}_{x \in W} \) is a collection of binary relations on \( x \uparrow \) (where \( x \uparrow := \{y \mid xRy\} \)). The requirements are that the \( S_x \) are reflexive and transitive and the restriction of \( R \) to \( x \uparrow \) is contained in \( S_x \), that is \( R \cap (x \uparrow) \subseteq S_x \).

A Veltman model consists of a Veltman frame together with a valuation \( V : \text{Prop} \rightarrow P(W) \) that assigns to each propositional variable \( p \in \text{Prop} \) a set of worlds \( V(p) \) in \( W \) where \( p \) is stipulated to be true. This valuation defines a forcing relation \( \models \subseteq W \times \text{Form} \) telling us which formulas are true at which particular world:

\[
\begin{align*}
\text{for no } x \in W; \\
x \not\models A \rightarrow B & : \Leftrightarrow x \not\models A \text{ or } x \models B; \\
x \models \Box A & : \Leftrightarrow \forall y (xRy \rightarrow y \models A); \\
x \models A \triangleright B & : \Leftrightarrow \forall y (xRy \land y \models A \rightarrow \exists z (ySxz \land z \models B)).
\end{align*}
\]

For a Veltman model \( \mathcal{M} = \langle W, R, \{S_x\}_{x \in W}, V \rangle \), we shall write \( \mathcal{M} \models A \) as short for \( \forall x \in W \ \mathcal{M}, x \models A \).

The logic **IL** is sound and complete with respect to all Veltman models ([3]). Often one is interested in considering all models that can be defined over a frame. Thus, given a frame \( \mathcal{F} \) and a valuation \( V \) on \( \mathcal{F} \) we shall denote the corresponding model by \( \langle \mathcal{F}, V \rangle \). A frame condition for a modal formula \( P \) is a formula \( F \) (first or higher-order) in the language \( \{R, \{S_x\}_{x \in W}\} \) so that \( \mathcal{F} \models F \) (as a relational structure) if and only if \( \forall \text{valuation } V \langle \mathcal{F}, V \rangle \models P \).
It is easy to establish that the frame condition for $P$ is $xRyRzS_xu \rightarrow zS_yu$ where $xRyRzS_xu$ is short for $xRy \land yRz \land zS_xu$. Likewise, it is elementary to see that the frame condition for $M$ is given by $yS_xzRu \rightarrow yRu$. In this paper we shall compute the frame conditions for two new series of principles in $\text{IL(All)}$.

Often we shall denote a valuation $V$ directly by the induced forcing relation $\vdash$. Given a Veltman model $(\mathcal{F}, \vdash)$ we define a $C$-assuring successor $-$denoted by $R^C_{\vdash}-$ as follows

$$xR^C_{\vdash}y := (xRy \land y \vdash C \land \forall z (yS_xz \rightarrow z \vdash C)).$$

## 3 A slim hierarchy of principles

In this section we present a hierarchy of interpretability principles in $\text{IL(All)}$ of growing strength. For a well-behaved sub-hierarchy we shall compute the frame conditions and prove arithmetical soundness. There is no particular ‘slimness’ inherent to the hierarchy presented here. The main reason for our name is that we tend to depict the frame conditions (see Figure 1) in a slim way as opposed to the depicted frame conditions for the series of principles that we refer to as a broad series of principles (see Figure 2).

### 3.1 A slim hierarchy

Inductively, we define a series of principles as follows.

\begin{align*}
R_0 & := A_0 \triangleright B_0 \rightarrow \neg(A_0 \triangleright \neg C_0) \triangleright B_0 \land \Box C_0 \\
R_{2n+1} & := R_{2n}[\neg(A_n \triangleright \neg C_n) \land (E_{n+1} \triangleright \Box A_{n+1})] \\
R_{2n+2} & := R_{2n+1}[B_n / B_n \land (A_{n+1} \triangleright B_{n+1}) \\
& \quad \triangleright \Diamond A_{n+1} / (E_{n+1} \triangleright \Box A_{n+1})] \\
& \quad \land (E_{n+1} \triangleright B_{n+1}) \\
& \quad \land (E_{n+1} \triangleright B_{n+1} \land \Box C_{n+1})]
\end{align*}

As to illustrate how these substitutions work we shall calculate the first five principles.

\begin{align*}
R_0 & := A_0 \triangleright B_0 \rightarrow \neg(A_0 \triangleright \neg C_0) \triangleright B_0 \land \Box C_0 \\
R_1 & := A_0 \triangleright B_0 \rightarrow \neg(A_0 \triangleright \neg C_0) \land (E_1 \triangleright \Box A_1) \triangleright B_0 \land \Box C_0 \land (E_1 \triangleright A_1) \\
R_2 & := A_0 \triangleright B_0 \land (A_1 \triangleright B_1) \rightarrow \neg(A_0 \triangleright \neg C_0) \land (E_1 \triangleright \Box A_1) \land (E_1 \triangleright \neg C_1) \triangleright \\
& \quad B_0 \land (A_1 \triangleright B_1) \land \Box C_0 \land (E_1 \triangleright A_1) \land (E_1 \triangleright B_1 \land \Box C_1) \\
R_3 & := A_0 \triangleright B_0 \land (A_1 \triangleright B_1) \rightarrow \\
& \quad \neg(A_0 \triangleright \neg C_0) \land (E_1 \triangleright \neg A_1 \land \Box C_1) \land (E_1 \triangleright \neg C_1) \land (E_2 \triangleright \Box A_2) \land B_0 \land (A_1 \triangleright B_1) \land \Box C_0 \land (E_1 \triangleright A_1) \land (E_1 \triangleright B_1 \land \Box C_1) \\
R_4 & := A_0 \triangleright B_0 \land (A_1 \triangleright B_1 \land (A_2 \triangleright B_2)) \rightarrow \\
& \quad \neg(A_0 \triangleright \neg C_0) \land (E_1 \triangleright \neg A_1 \land \Box C_1) \land (E_2 \triangleright \neg A_2 \land \Box C_2) \land (E_1 \triangleright B_1) \land (E_2 \triangleright A_2) \land (E_2 \triangleright B_2 \land \Box C_2)
\end{align*}
It is easy to see that the hierarchy defines a series of principles of increasing strength as expressed by the following lemma.

**Lemma 3.1.** For each natural number \( n \) we have that \( \text{IL} \mathcal{R}_{n+1} \vdash \mathcal{R}_n \).

**Proof.** By an easy case distinction. We see that \( \vdash_{\text{IL}} \mathcal{R}_{2n+1} \rightarrow \mathcal{R}_{2n} \) by choosing \( E_{n+1} := \Diamond \top \) and \( A_{n+1} := \top \). To see that \( \vdash_{\text{IL}} \mathcal{R}_{2n+2} \rightarrow \mathcal{R}_{2n+1} \) we choose \( C_{n+1} := \top \) and \( B_{n+1} := A_{n+1} \).

Thus, to understand the hierarchy well, it suffices to study a well-behaved co-final subsequence of it. To this end we define the following hierarchy.

For any \( n \geq 0 \) define schemata \( X_n, Y_n \) and \( Z_n \) as follows.

\[
\begin{align*}
X_0 &= A_0 \triangleright B_0; \\
Y_0 &= \neg (A_0 \triangleright C_0); \\
Z_0 &= B_0 \land \square C_0; \\
X_{n+1} &= A_{n+1} \triangleright B_{n+1} \land (X_n); \\
Y_{n+1} &= \neg (A_{n+1} \triangleright \neg C_{n+1}) \land (E_{n+1} \triangleright Y_n); \\
Z_{n+1} &= B_{n+1} \land (X_n) \land \square C_{n+1} \land (E_{n+1} \triangleright A_n) \land (E_{n+1} \triangleright Z_n).
\end{align*}
\]

For any \( n \geq 0 \) define

\[
\mathcal{R}_n = X_n \rightarrow Y_n \triangleright Z_n.
\]

To see how this proceeds, let us evaluate the first couple of instances:

\[
\begin{align*}
\mathcal{R}_0 &:= A_0 \triangleright B_0 \rightarrow \neg (A_0 \triangleright \neg C_0) \triangleright B_0 \land \square C_0; \\
\mathcal{R}_1 &:= A_1 \triangleright B_1 \land (A_0 \triangleright B_0) \rightarrow \\
&\neg (A_1 \triangleright \neg C_1) \land (E_1 \triangleright \neg (A_0 \triangleright \neg C_0)) \triangleright \neg (E_1 \triangleright A_0) \land B_1 \land (A_0 \triangleright B_0) \land \square C_1 \land (E_1 \triangleright A_0) \land (E_1 \triangleright B_0 \land \square C_0); \\
\mathcal{R}_2 &:= A_2 \triangleright B_2 \land (A_1 \triangleright B_1 \land (A_0 \triangleright B_0)) \rightarrow \\
&\neg (A_2 \triangleright \neg C_2) \land (E_2 \triangleright \neg (A_1 \triangleright \neg C_1) \land (E_1 \triangleright \neg (A_0 \triangleright \neg C_0))) \triangleright \neg (E_2 \triangleright A_1) \land B_2 \land (A_1 \triangleright B_1 \land (A_0 \triangleright B_0)) \land \square C_2 \land (E_2 \triangleright A_1) \land (E_2 \triangleright B_1 \land (A_0 \triangleright B_0) \land \square C_1 \land (E_1 \triangleright A_0) \land (E_1 \triangleright B_0 \land \square C_0));
\end{align*}
\]

It is clear that the \( \mathcal{R}_k \) hierarchy is directly related to the \( \mathcal{R}_k \) hierarchy:

**Lemma 3.2.** For each natural number \( k \) we have \( \mathcal{R}_{2k} := \mathcal{R}_k [X_i / X_{k-i}; E_i / E_{k+1-i}] \), where \( X \in \{ A, B, C \} \).

**Proof.** By visual inspection we see that it holds for \( k = 0, 1 \). It is proven in full generality by an easy induction. To prove the lemma, it is best to consider the place-holders like \( A_i \) etc. as propositional variables since otherwise in principle, for example, \( A_i \) could contain \( E_i \) as a subformula.

For the remainder of this section, we shall focus on the \( \mathcal{R}_k \) hierarchy and begin by computing a collection of frame conditions.
3.2 Frame conditions

For any \( n \geq 0 \) we define a ternary relation \( G_n(x, y, z) \) on Veltman-frames as follows.

\[
G_0(x, y, z) = \forall u \ (zRu \Rightarrow yS_xu),
\]
\[
G_{n+1}(x, y, z) = \forall u \ (zRu \Rightarrow yS_xu \land \forall v (uS_xv \Rightarrow G_n(z, u, v))).
\]

For every \( n \geq 0 \) we define the first-order frame condition \( F_n \) as follows.

\[
F_n = \forall w, x, y, z \ (wRx Ry S_wz \Rightarrow G_n(x, y, z)).
\]

The main result of this subsection is that \( F_{2n} \) is the frame correspondence of \( R_n \). For \( n = 0 \) this has been established in [4]. It is easy to see that \( G_{n+1}(x, y, z) \) implies \( G_n(x, y, z) \) so that \( F_{n+1} \) also implies \( F_n \). The frame conditions \( F_k \) are depicted in Figure 1 for the first three values of \( k \).

![Figure 1: From left to right we have depicted \( F_0 \) to \( F_2 \). Since \( F_{k+1} \) implies \( F_k \) we have only depicted the content of \( F_{k+1} \) which is new w.r.t. \( F_k \). As such we should read the pictures as: “if all un-dashed relations are as in the picture, then also the dashed relation should be present”.](image)

In what follows we let \( F = \langle W, R, S \rangle \) be an arbitrary Veltman-frame. With a forcing relation \( \models \) we will always mean a forcing relation on \( F \). For our convenience we define

\[
A_{-1} \equiv X_{-1} \equiv Z_{-1} \equiv \top.
\]
Before we can prove a frame correspondence we first need a technical lemma.

**Lemma 3.3.** For all \( k \geq 0 \) and all \( x, y, z \in W \). If \( \mathcal{G}_{2k}(x, y, z) \) then for any forcing relation \( \Vdash \) for which

\[
x \Vdash Y_k \quad \text{and} \quad xR^C_{\nu} y \quad \text{and} \quad z \Vdash X_{k-1},
\]

we also have

\[
z \Vdash □C_k \land (E_k \Vdash A_{k-1}) \land (E_k \Vdash Z_{k-1}).
\]

**Proof.** We shall write \( xR^C_k y \) as short for \( xR^C_k \Vdash y \) and prove the claim by induction on \( k \). With the convention that \( A_{-1} \equiv Z_{-1} \equiv \top \) the lemma is trivial for \( k = 0 \). So assume \( k > 0 \). Let \( \Vdash \) be a forcing relation and take \( x, y \) and \( z \) such that

\[
x \Vdash \neg(A_k \Vdash \neg C_k) \land (E_k \Vdash Y_{k-1}), \quad (4)
\]

\[
xR^C_k y, \quad (5)
\]

\[
z \Vdash X_{k-1}, \quad (6)
\]

\[
\mathcal{G}_{2k}(x, y, z). \quad (7)
\]

Take an arbitrary \( u \in W \) with \( zRu \). By (7) we have \( yS_x u \) and thus by (5) we have \( u \Vdash C_k \). This shows \( z \Vdash □C_k \).

To show that also the other two conjuncts hold at \( z \) assume that \( u \Vdash E_k \).

By (4) we find some \( v \) with \( uS_x v \) and

\[
v \Vdash Y_{k-1}. \quad (8)
\]

In order to show \( z \Vdash E_k \Vdash A_{k-1} \) we have to find some \( a \) with \( uS_x a \Vdash A_{k-1} \). Remark that \( Y_{k-1} \) implies \( □A_{k-1} \) thus there exists some \( a \) with \( vRa \Vdash A_{k-1} \).

By (7) we have \( \mathcal{G}_{2k-1}(z, u, v) \) and thus \( uS_x a \).

In order to show that also \( z \Vdash E_k \Vdash Z_{k-1} \) we have to find some \( b \) with \( uS_x b \Vdash Z_{k-1} \). We just used that \( Y_{k-1} \) implies \( □A_{k-1} \), but remark that \( Y_{k-1} \) implies the stronger statement that \( \neg(A_{k-1} \Vdash \neg C_{k-1}) \). Thus there exists some \( a \) with \( a \Vdash A_{k-1} \) and

\[
vR^C_{k-1} a. \quad (9)
\]

As above, by (7) we have \( \mathcal{G}_{2k-1}(z, u, v) \) and thus \( uS_x a \) and \( vRa \). By (9) there exists a \( b \) with \( aS_x b \) and

\[
b \Vdash B_{k-1} \land (X_{k-2}). \quad (10)
\]

Since \( uS_x a \) whence also \( uS_x b \) holds, we will be done if we show that \( b \Vdash Z_{k-1} \). To show that the remaining conjuncts of \( Z_{k-1} \) hold at \( b \) (that is \( b \Vdash □C_{k-1} \land (E_{k-1} \Vdash A_{k-2}) \land (E_{k-1} \Vdash Z_{k-2})) \) simply observe that \( \mathcal{G}_{2k-2}(v, a, b) \) and use \( 8, 9 \) and \( 10 \) to invoke the (IH) on \( v, a \) and \( b \).

**Corollary 3.4.** If \( F \Vdash \mathcal{F}_{2k} \), then \( F \Vdash \tilde{R}_k \).
Proof. Fix a forcing relation $\models$ and let $w, x \in W$ such that $w \models X_k$ and $wRx \models Y_k$. Then for some $y$ we have $xR^c y \models A_k$. Thus there exists $z$ with $yS_w z$ and

$$z \models B_k \wedge (X_{k-1})$$  \hfill (11)

(recall $X_{-1} \equiv \top$). Since $F \models F_{2k}$ we have $G_{2k}(x, y, z)$. Thus by Lemma 3.3 we get

$$z \models \Box C_k \wedge (E_k \triangleright A_{k-1}) \wedge (E_k \triangleright Z_{k-1})$$  \hfill (12)

Combining (11) and (12) gives $z \models Z_k$.

The reversal of this corollary is again preceded by a technical lemma. We shall denote by $a_k, b_k, c_k,$ and $e_k$, propositional variables that shall play the role of the $A_k, B_k, C_k$ and $E_k$ respectively in the principles $R_n$. Likewise, by $\nabla_k$ we shall denote the formula that arises by substituting $a_j$ for $A_j$ in $X_k$ and $b_j$ for $B_j$. The formulas $Y_k$ and $Z_k$ are defined similarly.

Lemma 3.5. For any $k \geq 0$ and all $x, y, z \in W$. If for all forcing relations $\models$ for which $x \models Y_k$ and $xR^c y$ and $z \models X_{k-1}$ we also have $z \models \Box c_k \wedge (e_k \triangleright a_{k-1}) \wedge (e_k \triangleright Z_{k-1})$, then $G_{2k}(x, y, z)$.

Proof. Induction on $k$. Let $x, y, z \in W$ and assume the conditions of the lemma. Unfolding the definition of $G_{2k}(x, y, z)$ shows us that we have to show that

1. for all $u$ with $zRu$ we have $yS_x u$ $(k \geq 0)$;
2. and for all $v$ and $a$ with $uS_{xv}$ and $vRa$ we have $uS_z a$ $(k > 0)$;
3. and for all $b$ with $aS_x b$ we have $G_{2(k-1)}(v, a, b)$ $(k > 0)$.

We will show 1 and 2 'by hand' and invoke the (IH) for 3. In each of the three cases we will choose similar but different forcing relations $\models$.

We first show 1. So let $zRu$. Define

$$w \models c_k \leftrightarrow yS_x w \quad \text{and} \quad w \models a_k \leftrightarrow w = y.$$

And let all the other variables be false everywhere. Then $xR^c y$ and $x \models \neg (a_k \triangleright \neg c_k)$. Since none of the $e_i$ nor $a_j$ with $j \neq k$ holds anywhere in the model, we trivially have $x \models \nabla_k$ and $z \models \nabla_{k-1}$ and thus according to the conditions of the lemma in particular $z \models \Box c_k$. By definition of $\models$ we thus have $yS_x u$ which proves 1. Note that for $k = 0$ we only have to look after 1 hence we have now dealt with the base case of our induction.

Now we continue to show 2 assuming $k > 0$. Choose any $v$ and $a$ with $uS_{xv}$ and $vRa$. As above define

$$w \models c_k \leftrightarrow yS_x w \quad \text{and} \quad w \models a_k \leftrightarrow w = y.$$

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We now also define

\[
\begin{align*}
  & w \models e_k \iff w = u \quad \text{and}, \\
  & w \models a_{k-1} \iff w = a \iff w \models b_{k-1} \quad \text{and}, \\
  & w \models c_{k-1} \iff aS_k w.
\end{align*}
\]

Let all the other propositional variables be false everywhere. Now \( v \models \nabla_{k-1} \) and thus \( x \models \nabla_k \). It is not hard to see that we also have \( z \models \nabla_{k-1} \) and thus according to the condition of the lemma we have in particular \( z \models e_k \supset a_{k-1} \). Since \( zRu \models e_k \) there must be an \( a' \) with \( uS_z a' \models a_{k-1} \). Since \( a \) is the only world that forces \( a_{k-1} \) we must have \( uS_z a \).

To finish and show \( \square \) choose \( b \) such that \( aS_z b \). We want to show that \( \mathcal{G}_{2(k-1)}(v, a, b) \). Invoking the (IH) it is enough to show that for any forcing relation \( \models \) for which

\[
\text{we have}\quad \text{(13)}
\]

we also have

\[
\text{we also have}\quad \text{(14)}
\]

Our strategy in proving this is as follows. We slightly tweak \( \models \) to obtain \( \models' \). This \( \models' \) is similar to \( \models \) in that \( (13) \) still holds and moreover

\[
\text{we have}\quad \text{(15)}
\]

However, it is (possibly) different in that we now know that \( x \models' \nabla_k \) and \( xR_{k-1}^c y \) and, \( z \models' \nabla_{k-1} \) so that we may apply the main assumption of the lemma to \( \models' \) concluding \( z \models' \square c_k \land (e_k \supset a_{k-1}) \land (e_k \supset \nabla_{k-1}) \). The latter will help us conclude \( (13) \).

Thus we consider an arbitrary forcing relation \( \models \) that satisfies \( (13) \). We modify \( \models \) to obtain \( \models' \) such that it satisfies

\[
\begin{align*}
  & w \models' a_k \quad \iff \quad w = y; \\
  & w \models' e_k \quad \iff \quad w = u; \\
  & w \models' c_k \quad \iff \quad yS_z w; \\
  & w \models' a_{k-1} \quad \iff \quad w = a; \\
  & w \models' b_{k-1} \quad \iff \quad w = b.
\end{align*}
\]

Apart from these modifications, \( \models' \) will coincide with \( \models \). It is a straightforward check to see that we have \( (13) \) for \( \models' \) and that moreover \( (15) \) holds. In addition, by the definition of \( \models' \) we now also have

\[
\text{we have}\quad \text{(16)}
\]

Thus, we see that \( \models' \) satisfies the antecedent of the condition of the lemma. Consequently, we have \( z \models' e_k \supset \nabla_{k-1} \). Since \( zRu \models' e_k \), there must exist some \( b' \) with \( uS_z b' \models' \nabla_{k-1} \). But now, since \( b_{k-1} \) is a conjunct of \( \nabla_{k-1} \) and \( b' \) is the only world that \( \models' \) forces \( b_{k-1} \), we must have \( b \models' \nabla_{k-1} \). In particular, we conclude \( b \models' (e_{k-1} \supset a_{k-2}) \land (e_{k-1} \supset \nabla_{k-2}) \land \square c_{k-1} \); by \( (15) \) the same holds for \( \models \) and we are done.
Putting this all together gives us the frame correspondence for \( \tilde{R}_k \).

**Theorem 3.6.** For any Veltman frame \( F \) and any natural number \( k \geq 0 \) we have

\[
F \models \mathcal{F}_{2k} \iff F \models \tilde{R}_k \iff F \models R_{2k}.
\]

**Proof.** The second equivalence is a direct consequence of Lemma 3.2 so we focus on the first equivalence.

The \( \Rightarrow \) direction is just Corollary 3.4. For the other direction, fix some \( k \), assume that \( F \models R_k \) and let \( wR RxRyS wz \). We have to show that \( \mathcal{G}_{2k}(x, y, z) \). Now consider any forcing relation \( \models \) that satisfies \( x R_{\mathcal{G}_k} y \), and \( x \models \mathcal{V}_k \) and, \( z \models \mathcal{V}_{k-1} \). By Lemma 3.5 it is enough to show that

\[
z \models \Box c_k \land (e_k \triangleright a_{k-1}) \land (e_k \triangleright \mathcal{Z}_{k-1}). \tag{17}
\]

Now consider a forcing relation \( \models ' \) where \( \models ' \) is like \( \models \) except that

\[
v \models ' a_k \iff v = y \quad \text{and} \quad v \models ' b_k \iff v = z.
\]

Notice that \( x R_{\mathcal{G}_k} y \) and thus also \( x \models \mathcal{V}_k \). But now we have \( w \models ' \mathcal{V}_k \) as well and thus \( w \models ' \mathcal{V}_k \triangleright \mathcal{Z}_k \). Thus there must be some \( z' \) with \( x S_w z' \models ' \mathcal{Z}_k \). Since \( b_k \) is a conjunct of \( \mathcal{Z}_k \) and \( z \) is the only world where \( b_k \) is forced we must have \( z \models ' \mathcal{Z}_k \). Since \( \Box c_k \land (e_k \triangleright a_{k-1}) \land (e_k \triangleright \mathcal{Z}_{k-1}) \) does not involve \( a_k \) nor \( b_k \) we have (17). \( \square \)

### 3.3 Arithmetical soundness

Via a series of lemmata we shall prove Theorem 3.7 to the effect that the hierarchy \( \{R_i\}_{i \in \omega} \) is arithmetically sound in any reasonable arithmetical theory.

**Theorem 3.7.** Each of the \( R_i \) is arithmetically sound in any theory extending \( S_2^1 \).

It is sufficient to prove that each of the \( R_{2m} \) is arithmetically sound in any reasonable arithmetical theory whence we shall focus on the principles \( R_i \). We shall first exhibit a soundness proof of \( R_1 \) and then indicate how this is generalized to the rest of the hierarchy. And before proving \( R_1 \) we need some auxiliary lemmas.

**Lemma 3.8.** Let \( T \) be any theory extending \( S_2^1 \). We have that for any arithmetical sentences \( E_1, A_0, B_0 \) and \( C_0 \) that

\[T \vdash E_1 \Rightarrow \neg(A_0 \Rightarrow \neg C_0) \rightarrow \exists \text{Cut} \ J \ \Diamond (E_1 \rightarrow \forall \text{Cut} \ K \in J \ \Diamond (A_0 \land \Box K C_0)).\]

**Proof.** Reason in \( T \) and assume \( E_1 \Rightarrow \neg(A_0 \Rightarrow \neg C_0) \). Note that by Lemma 2.6 we have \( E_1 \Rightarrow \forall \text{Cut} \ K \Diamond (A_0 \land \Box K C_0) \). Consequently, by Pudlák’s Lemma, Lemma 2.2 we get \( \exists J \ (E_1 \land \exists \text{Cut} \ K \in J \ \Diamond \neg(A_0 \land \Box K C_0) \Rightarrow \bot) \). But this is provably the same as \( \exists \text{Cut} \ J \ \Diamond (E_1 \rightarrow \forall \text{Cut} \ K \in J \ \Diamond (A_0 \land \Box K C_0)) \) as was to be shown. \( \square \)
Lemma 3.9. Let $T$ be any theory extending $S^1_2$. We have that for any arithmetical sentences $E_1, A_0, B_0$ and $C_0$ that

$$T \vdash \exists \text{Cut} \, j \, \Box \left( E_1 \rightarrow \forall \text{Cut} \, k \in \bar{j} \, \Diamond^j (A_0 \land \Box^k C_0) \right) \rightarrow E_1 \rhd A_0.$$

Proof. Reasoning in $T$. From the assumption we get in particular that $\exists \text{Cut} \, j \, \Box \left( E_1 \rightarrow \Diamond^j A_0 \right)$ so that $\exists \text{Cut} \, j \, E_1 \rhd \Diamond^j A_0 \rhd A_0$. \hfill $\square$

Lemma 3.10. Let $T$ be any theory extending $S^1_2$. We have that for any arithmetical sentences $E_1, A_0, B_0$ and $C_0$ that

$$T \vdash (A_0 \rhd B_0) \land \exists \text{Cut} \, j \, \Box \left( E_1 \rightarrow \forall \text{Cut} \, k \in \bar{j} \, \Diamond^j (A_0 \land \Box^k C_0) \right) \rightarrow E_1 \rhd B_0 \land \Box C_0.$$

Proof. Reasoning in $T$ we get from $\exists \text{Cut} \, j \, \Box \left( E_1 \rightarrow \forall \text{Cut} \, k \in \bar{j} \, \Diamond^j (A_0 \land \Box^k C_0) \right)$ that $\forall \text{Cut} \, K \left( E_1 \rhd A_0 \land \Box^k C_0 \right)$. We combine this with $A_0 \rhd B_0 \rightarrow \exists \text{Cut} \, j \left( A_0 \land \Box^j C_0 \rhd B_0 \land \Box C_0 \right)$ to conclude $E_1 \rhd B_0 \land \Box C_0$. \hfill $\square$

With these technical lemmas we can prove soundness of $\bar{R}_1$.

Lemma 3.11. Let $T$ be any theory extending $S^1_2$. We have that for any arithmetical sentences $E_1, A_1, B_1, A_0, B_0$ and $C_0$ that

$$T \vdash A_1 \rhd B_1 \land (A_0 \rhd B_0) \rightarrow \neg (A_1 \rhd \neg C_1) \land (E_1 \rhd \neg (A_0 \rhd \neg C_0)) \rhd B_1 \land (A_0 \rhd B_0) \land \Box C_1 \land (E_1 \rhd A_0) \land (E_1 \rhd B_0 \land \Box C_0).$$

Proof. We reason in $T$. Using our new technical lemma and Lemma 2.6 we get

$$\neg (A_1 \rhd \neg C_1) \land (E_1 \rhd \neg (A_0 \rhd \neg C_0)) \rightarrow \forall \text{Cut} \, K \left( (A_1 \land \Box^k C_1) \land \exists \text{Cut} \, j \, \Box \left( E_1 \rightarrow \forall \text{Cut} \, L \in \bar{j} \, \Diamond^j (A_0 \land \Box^l C_0) \right) \right) \rightarrow \forall \text{Cut} \, K \left( (A_1 \land \Box^k C_1) \land \exists \text{Cut} \, j \, \Box \left( E_1 \rightarrow \forall \text{Cut} \, L \in \bar{j} \, \Diamond^j (A_0 \land \Box^l C_0) \right) \right).$$

The last step is due to the principle of outside-big inside-small (Lemma ??) and allows us to conclude

$$\forall \text{Cut} \, K \left( (A_1 \rhd \neg C_1) \land (E_1 \rhd \neg (A_0 \rhd \neg C_0)) \rhd A_1 \land \Box^k C_1 \land \exists \text{Cut} \, j \, \Box \left( E_1 \rightarrow \forall \text{Cut} \, L \in \bar{j} \, \Diamond^j (A_0 \land \Box^l C_0) \right) \right).$$

This can be combined with the fact that

$$A_1 \rhd B_1 \land (A_0 \rhd B_0) \rightarrow \exists \text{Cut} \, K \left( A_1 \land \Box^k \lhd B_1 \land \Box \left( A_0 \rhd B_0 \right) \right)$$

for this particular $K$ holds for any $\sigma \in \Sigma_1$ to conclude

$$A_1 \rhd B_1 \land (A_0 \rhd B_0) \rightarrow \neg (A_1 \rhd \neg C_1) \land (E_1 \rhd \neg (A_0 \rhd \neg C_0)) \rhd B_1 \land (A_0 \rhd B_0) \land \Box C_1 \land \exists \text{Cut} \, j \, \Box \left( E_1 \rightarrow \forall \text{Cut} \, L \in \bar{j} \, \Diamond^j (A_0 \land \Box^l C_0) \right).$$
Let \( \mathcal{R}_k \) be arbitrary. For any number \( k \) we have that

\[
T \vdash E_{k+1} \supset Y_k \rightarrow \mathcal{H}_{k+1}.
\]
Proof. By an external induction on $k$. For $k = 0$ this is simply Lemma 3.8.

For the inductive case we reason in $T$ and see that $E_{k+2} \triangleright Y_{k+1} \equiv E_{k+2} \triangleright \neg (A_{k+1} \triangleright \neg C_{k+1}) \land (E_{k+1} \triangleright Y_k)$. By the inductive hypothesis we have that $E_{k+1} \triangleright Y_k \rightarrow \mathcal{H}_{k+1}$ so that $E_{k+2} \triangleright Y_{k+1} \rightarrow E_{k+2} \triangleright \neg (A_{k+1} \triangleright \neg C_{k+1}) \land \mathcal{H}_{k+1}$. Since $\mathcal{H}_{k+1}$ is equivalent to an $\Sigma^b_1$ formula, by Lemma 3.12 we see that

$$E_{k+2} \triangleright \neg (A_{k+1} \triangleright \neg C_{k+1}) \land \mathcal{H}_{k+1} \rightarrow \mathcal{H}_{k+2}$$

as was to be shown. \hfill \QED

Moreover, the $\mathcal{H}_{k+1}$ formulas contain all the information to get the induction going as shown by the following lemma.

**Lemma 3.14.** Let $T$ be a theory containing $S^b_2$ and let the formulas $E_i, A_i, B_i,$ and $C_i$ be arbitrary. For any number $k$ we have that

$$T \vdash (X_k) \land \mathcal{H}_{k+1} \rightarrow E_{k+1} \triangleright Z_k.$$  

**Proof.** By induction on $k$ where the case $k = 0$ is just Lemma 3.10. For the inductive case, we reason in $T$ and assume $(X_k) \land \mathcal{H}_{k+1}$.

From the definition of $\mathcal{H}_{k+2}$ we get

$$\exists \text{Cut} J_{k+2} \ □ (E_{k+2} \rightarrow \forall \text{Cut} K_{k+2} \in J_{k+2} \退出_jk+2 (A_{k+1} \land \Box (K_{k+2} C_{k+1} \land \mathcal{H}_{k+1})))$$

so that $\exists \text{Cut} J_{k+2} \forall \text{Cut} K_{k+2} \ □ (E_{k+2} \rightarrow \退出_jk+2 (A_{k+1} \land \Box (K_{k+2} C_{k+1} \land \mathcal{H}_{k+1})))$.

whence

$$\forall \text{Cut} K_{k+2} \ (E_{k+2} \triangleright A_{k+1} \land \Box (K_{k+2} C_{k+1} \land \mathcal{H}_{k+1})). \quad (18)$$

From $X_{k+1}$ –which is by definition equal to $A_{k+1} \triangleright B_{k+1} \land (X_k)$– we find via Pudlák’s lemma, Lemma 2.22 a specific cut $K_{k+2}$ such that for any formula $\sigma$ in $\Sigma_1$ we obtain $A_{k+1} \land \sigma \rightarrow B_{k+1} \land (X_k) \land \sigma$. We can plug in this cut $K_{k+2}$ to (18) to obtain via transitivity of $\triangleright$ that

$$E_{k+2} \triangleright B_{k+1} \land (X_k) \land \Box (C_{k+1} \land \mathcal{H}_{k+1}).$$

We are almost done but $B_{k+1} \land (X_k) \land \Box (C_{k+1} \land \mathcal{H}_{k+1}$ is not quite equal to $Z_{k+1}$ as was needed. The missing conjuncts are $E_{k+1} \triangleright A_k$ and $E_{k+1} \triangleright Z_k$. The first is easily seen to follow from $\mathcal{H}_{k+1}$ and the second follows from the inductive hypothesis applied to $(X_k) \land \mathcal{H}_{k+1}$. \hfill \QED

We are now ready to prove Theorem 3.7 that the whole hierarchy is arithmetically sound.

**Theorem 3.15.** Let $T$ be a theory containing $S^b_2$ and let $A_i, B_i, C_i$ and $E_i$ be arbitrary arithmetical formulas. We have for each number $k$ that

$$T \vdash \bar{R}_k \ id \ est \ T \vdash X_k \rightarrow Y_k \triangleright Z_k.$$
Proof. By an external induction on \( k \) where the base case is the soundness of \( \tilde{R}_0 \) which has been proven in \([1]\). Thus, we reason in \( T \) assuming \( A_{k+1} \vdash B_{k+1} \wedge (X_k) \).

We need to conclude that \( Y_{k+1} \vdash Z_{k+1} \). But \( Y_{k+1} \) is nothing but \( \neg(A_{k+1} \vdash \neg C_{k+1}) \wedge (E_{k+1} \vdash Y_k) \). By Lemma 3.13 we know that \( (E_{k+1} \vdash Y_k) \rightarrow \mathcal{H}_{k+1} \).

Using this and reasoning as before we obtain

\[
\neg(A_{k+1} \vdash \neg C_{k+1}) \wedge (E_{k+1} \vdash Y_k) \rightarrow \forall \text{Cut} \ K \, \Diamond (A_{k+1} \wedge \Box^K C_{k+1}) \wedge (E_{k+1} \vdash Y_k)
\]

Consequently,

\[
\forall \text{Cut} \ K \, \Diamond (A_{k+1} \wedge \Box^K C_{k+1}) \wedge (E_{k+1} \vdash Y_k) \doteq A_{k+1} \wedge \Box^K C_{k+1} \wedge \mathcal{H}_{k+1}^K.
\]

This can be combined with Pudlák’s Lemma on \( A_{k+1} \vdash B_{k+1} \wedge (X_k) \) to obtain

\[
\neg(A_{k+1} \vdash \neg C_{k+1}) \wedge (E_{k+1} \vdash Y_k) \doteq B_{k+1} \wedge (X_k) \wedge \Box C_{k+1} \wedge \mathcal{H}_{k+1}.
\]

It is easy to see that \( \mathcal{H}_{k+1} \) implies \( E_{k+1} \vdash A_k \). Moreover, Lemma 3.14 tells us that \( (X_k) \wedge \mathcal{H}_{k+1} \rightarrow E_{k+1} \vdash Z_k \) so that we may conclude

\[
\neg(A_{k+1} \vdash \neg C_{k+1}) \wedge (E_{k+1} \vdash Y_k) \doteq B_{k+1} \wedge (X_k) \wedge \Box C_{k+1} \wedge (E_{k+1} \vdash A_k) \wedge (E_{k+1} \vdash Z_k)
\]

as was to be shown. \( \square \)

4 A broad series of principles

In this section we present a different series of principles. We refer to this series as the broad series since the frame-conditions—see Figure 2—are typically represented over a broader area than the slim hierarchy as discussed above.

4.1 A broad series

In order to define the second series we first define a series of auxiliary formulas. For any \( n \geq 1 \) we define the schemata \( U_n \) as follows.

\[
U_1 := \Diamond \neg (D_1 \vdash \neg C),
\]

\[
U_{n+2} := \Diamond ((D_{n+1} \vdash D_{n+2}) \wedge U_{n+1}).
\]

Now, for \( n \geq 0 \) we define the schemata \( R^n \) as follows.

\[
R^0 := A \vdash B \rightarrow \neg (A \vdash \neg C) \vdash B \wedge \Box C,
\]

\[
R^{n+1} := A \vdash B \rightarrow U_{n+1} \wedge (D_{n+1} \vdash A) \vdash B \wedge \Box C.
\]

As an illustration we shall calculate the first four principles.

\[
R^0 := A \vdash B \rightarrow \neg (A \vdash \neg C) \vdash B \wedge \Box C
\]

\[
R^1 := A \vdash B \rightarrow \Diamond \neg (D_1 \vdash \neg C) \wedge (D_1 \vdash A) \vdash B \wedge \Box C
\]

\[
R^2 := A \vdash B \rightarrow \Diamond \left( (D_1 \vdash D_2) \wedge \Diamond \neg (D_1 \vdash \neg C) \right) \wedge (D_2 \vdash A) \vdash B \wedge \Box C
\]

\[
R^3 := A \vdash B \rightarrow \Diamond \left( (D_2 \vdash D_3) \wedge \Diamond \left( (D_1 \vdash D_2) \wedge \Diamond \neg (D_1 \vdash \neg C) \right) \right) \wedge (D_3 \vdash A) \vdash B \wedge \Box C
\]
While the series $R_i$ did define a hierarchy in that $R_{i+1} \vdash R_i$, we shall see that no such relation holds for the series $R^i$.

### 4.2 Frame conditions

It is not hard to determine the frame condition for the first couple of principles in this series and in Figure 2 we have depicted the first three frame-conditions. In this section we shall prove that the correspondence proceeds as expected. Informally, the frame condition for $R^n$ shall be the universal closure of

$$x_{n+1}Rx_n \ldots Rx_0y_0S_{x_1}y_1 \ldots S_{x_n}y_nS_{x_{n+1}}y_{n+1}Ru \rightarrow y_0S_{x_0}u. \quad (19)$$

![Figure 2: From left to right, this figure depicts the frame conditions $F^0$ through $F^2$ corresponding to $R^0$ through $R^2$. The reading convention is as always: if all the un-dashed relations are present as in the picture, then also the dashed relation should be there.](image)

In order to make this frame condition precise and prove it, we shall first recast it in a recursive fashion. In writing (19) recursively we shall use those variables that will emphasize the relation with (19). Of course, free variables can be renamed at the readers liking.

First, we start by introducing a relation $B_n$ that captures the antecedent of (19). Note that this antecedent says that first there is a chain of points $x_i$
related by $R$, followed by a chain of points $y_i$ related by different $S$ relations. The relation $B_n$ will be applied to the end-points of both chains where the condition on the intermediate points is imposed by recursion.

$$
B_0(x_1, x_0, y_0, y_1) := x_1 R x_1 R y_0 S x_1 y_1,
$$

$$
B_{n+1}(x_{n+2}, x_0, y_0, y_{n+2}) := \exists x_{n+1}, y_{n+1} \left( x_{n+2} R x_{n+1} \land B_n(x_{n+1}, x_0, y_0, y_{n+1}) \land y_{n+1} S x_{n+2} y_{n+2} \right).
$$

For every $n \geq 0$ we can now define the first order frame condition $F^n$ as follows.

$$
F^n := \forall x_{n+1}, x_0, y_0, y_{n+1} \left( B_n(x_{n+1}, x_0, y_0, y_{n+1}) \Rightarrow \forall u (y_{n+1} Ru \Rightarrow y_0 S x_0) \right).
$$

Sometimes we shall write $x_{n+1} B_n[x_0, y_0] y_{n+1}$ conceiving the quaternary relation $B_n$ as a binary relation indexed by the pair $x_0, y_0$. In what follows we let $F = \{W, R, S\}$ be an arbitrary Veltman-frame. The next lemma follows from an easy induction on $n$.

**Lemma 4.1.** For each number $n$ we have that $B_n[x_0, y_0] \subseteq R$, that is, if $x_{n+1} B_n[x_0, y_0] y_{n+1}$, then $x_{n+1} R y_{n+1}$.

To prove that $F \models F^n$ implies $F \models R^n$ we first need a technical lemma.

**Lemma 4.2.** Let $w \in W$ and $\models$ be a forcing relation on $F$. If

$$
x_{k+1} \models U_{k+1} \land (D_{k+1} \rhd A),
$$

then there exist $x_0, y_0$ and $y_{k+1}$ such that $B_k(x_{k+1}, x_0, y_0, y_{k+1})$, $x_0 R^C y_0$ and $y_{k+1} \models A$.

**Proof.** Induction on $k$. If $k = 0$ then $U_{k+1} = \emptyset \models \neg (D_1 \rhd \neg C)$ and the statement is easily checked. For the inductive case, we assume

$$
x_{k+2} \models U_{k+2} \land (D_{k+2} \rhd A).
$$

Recall that $U_{k+2} := \emptyset (D_{k+1} \rhd D_{k+2} \land U_{k+1})$. Thus, there exists some $x_{k+1}$ with $x_{k+2} R x_{k+1}$ and

$$
x_{k+1} \models (D_{k+1} \rhd D_{k+2}) \land U_{k+1}.
$$

Applying the (IH) (with $D_{k+2}$ substituted for $A$) we find $x_0, y_0$ and $y_{k+1}$ with $B_k(x_{k+1}, x_0, y_0, y_{k+1})$, $x_0 R^C y_0$ and $y_{k+1} \models D_{k+2}$. As $B_k(x_{k+1}, x_0, y_0, y_{k+1})$ we get $x_{k+1} R y_{k+1}$ (Lemma [4.1]). Since we had $x_{k+2} R x_{k+1}$ we see that $x_{k+2} R y_{k+1}$ $\models D_{k+2}$, and since $x_{k+2} \models D_{k+2} \rhd A$, we find some $y_{k+2}$ with $y_{k+1} S x_{k+2} y_{k+2}$ and $y_{k+2} \models A$. By definition of $B_{k+1}$ we have $B_{k+1}(x_{k+2}, x_0, y_0, y_{k+2})$. \qed

**Corollary 4.3.** If $F \models F^n$ then $F \models R^n$.  

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Proof. Induction on $n$. For $n = 0$ this is known (see [4]), so we assume $n > 0$. Let $\vdash$ be a forcing relation, let $x_{n+1}, x_n \in W$ and assume $x_{n+1} \vdash A > B$, $x_{n+1}R x_n$ and $x_n \vdash U_n \cap (D_n \vdash A)$. By Lemma 12 we find $x_0$, $y_0$ and $y_n$ such that $\mathcal{B}_{n-1}(x_n, x_0, y_0, y_n)$, with $x_0 R^C y_0$ and $y_n \vdash A$. We have that $\mathcal{B}_{n-1}(x_n, x_0, y_0, y_n)$ implies $x_n R y_n$ (Lemma 4.1) and thus since $x_{n+1}R x_n$, we also have $x_{n+1}R y_n \vdash A$. By assumption $x_{n+1} \vdash A > B$ so that for some $y_{n+1}$ we have $y_n S x_{n+1} y_{n+1} \vdash B$. Clearly, we also have $x_n S x_{n+1} y_{n+1}$ so that we are done if we have shown that $y_{n+1} \vdash \Box C$. To this extent, we choose some $u$ with $y_{n+1}Ru$. Since we have that $\mathcal{B}_n(x_{n+1}, x_0, y_0, y_{n+1})$, by $\mathcal{F}^n$ we have also $y_0 S x_0 u$. But $x_0 R^C y_0$ and thus we have $u \vdash C$, as required.

To prove the converse implication, we start again with a technical lemma. As before we shall denote by $a$, $b$, $c$, and $d_k$, propositional variables that shall play the role of the $A$, $B$, $C$ and $D_k$ respectively in the principles $\mathcal{R}^n$. Let $\overline{U}_k$ denote the formula that arises by simultaneously substituting $c$ for $C$ and $d_k$ for $D_k$ in $U_k$.

Lemma 4.4. Let $\{a, c, d_1, \ldots, d_{k+1}\}$ be a collection of distinct propositional variables. If $\mathcal{F} \vdash \mathcal{B}_k(x_{k+1}, x_0, y_0, y_{k+1})$, then there exists a forcing relation $\vdash$ on $\mathcal{F}$ such that

1. $x_{k+1} \vdash \overline{U}_{k+1} \land (d_{k+1} \vdash a)$;
2. $x \vdash c$ iff $y_0 S x_0 x$;
3. $x \vdash a \iff x = y_{k+1}$;
4. $x \not\vdash p$ for any $p \notin \{d_1, \ldots, d_{k+1}, c, a\}$.

Proof. The idea is very simple using the informal description of $\mathcal{B}_k$ being the antecedent of (14). We define a valuation $\vdash$ so that $d_{k+1}$ is only true at $y_i$ and $a$ is only true at $y_{k+1}$. Moreover, we define $x \vdash c$ iff $y_0 S x_0 x$ and $x \not\vdash p$ for any $p \notin \{d_1, \ldots, d_{k+1}, c, a\}$. It is not hard to see that $x_{k+1} \vdash \overline{U}_{k+1} \land (d_{k+1} \vdash a)$ for this valuation $\vdash$.

To make the argument precise, we proceed by induction on $k$. If $k = 0$ then $\mathcal{B}_k(x_1, x_0, y_0, y_1)$ simply means $x_1 R x_0 R y_0 S x_1 y_1$ and we define

$$x \vdash a \iff x = y_1, \quad x \vdash c \iff y_0 S x_0 x \quad \text{and,} \quad x \vdash d_1 \iff x = y_0.$$ 

The lemma is easily checked if we further define $x \not\vdash p$ for any $p \notin \{d_1, c, a\}$.

For the inductive case we consider $k > 0$. Then $\mathcal{B}_k(x_{k+1}, x_0, y_0, y_{k+1})$ implies that there are $x_k$ and $y_k$ such that

$$x_{k+1} R x_k \mathcal{B}_{k-1}[x_0, y_0] y_k S x_{k+1} y_{k+1}.$$ 

The (IH) (with $d_{k+1}$ substituted for $a$) gives a forcing relation $\vdash$ such that

$$x_k \vdash \overline{U}_k \land (d_k \vdash d_{k+1}), \quad x_0 R^C y_0, \quad x \vdash d_{k+1} \iff x = y_k$$ 

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and $x \not\models p$ for $p \notin \{d_1, \ldots, d_{k+1}, c\}$. So we have $x_{k+1} \models \Diamond (\bigcup_k \land (d_k \triangleright d_{k+1}))$; in other words $x_{k+1} \models \bigcup_k$. We now define $\models'$ as follows

$$x \models' a \iff x = y_{k+1} \quad \text{and} \quad x \models' p \iff x \models p \quad \text{for} \quad p \neq a.$$  

Clearly, the properties $x_k \models \bigcup_k \land (d_k \triangleright d_{k+1}), aR_d b, x \models d_{k+1} \iff x = y_k$ simply extend to $\models'$ and likewise we have that $x \not\models' p$ for any $p \notin \{d_1, \ldots, d_{k+1}, c, a\}$. Moreover, we now have $x_{k+1} \models' d_{k+1} \triangleright a$ as well.

As a corollary to this lemma, we can now obtain the full frame conditions for the principles $R^n$.

**Theorem 4.5.** For each number $n$ we have $F \models F^n$ iff $F \models R^n$.

**Proof.** The $\Rightarrow$ direction is just Corollary 4.3 so we focus on the other direction. Thus, we suppose that $F \models R^n$, consider any $x_{n+1}, x_0, y_0, y_{n+1} \in W$ with $E_n(x_{n+1}, y_0, y_{n+1})$ and set out to show that for any $u$ with $y_{n+1}Ru$ we have $y_0Sx_0u$. We now apply Lemma 4.4 and simultaneously substitute $a$ for $d_{n+1}$ and $b$ for $a$ to see that there exists a forcing relation $\models$ such that

$$x_{n+1} \models \bigcup_n [d_{n+1}/a] \land (a \triangleright b), \quad x \models c \iff y_0Sx_0x \quad \text{and} \quad x \models b \iff x = y_{n+1}.$$

Since $n = 0$ is known, we assume $n > 0$. Thus, we find $x_n$ with $x_{n+1}Rx_n$ and $x_n \models \bigcup_n \land d_n \triangleright a$ (note that $\bigcup_n [d_{n+1}/a] = \bigcup_{n-1}$). Using $F \models R^n$ we see that there must exist some $x$ with $x \models b \land \Box c$. But $y_{n+1}$ is the only world that forces $b$ thus necessarily $y_{n+1} \models \Box c$. By the choice of $\models$ we thus have that if $y_{n+1}Ru$ then $y_0Sx_0u$. □

Using the frame condition we readily see that the broad series of principles does not define a hierarchy.

**Corollary 4.6.** For $n \neq m$ we have $ILR^n \not\models ILR^m$.

**Proof.** For each $m \neq n$ it is easy to exhibit a frame $F$ so that $F \models F^n$ but $F \not\models F^m$. □

### 4.3 Arithmetical soundness

We will now see that all the principles $R^n$ are arithmetically sound and begin with a simple lemma.

**Lemma 4.7.** For any theory $T$ extending $S^1_2$ and any natural number $n > 0$, we have that

$$T \models \bigcup_n \rightarrow \forall \text{Cut} \kappa \Diamond (D_n \land \Box^K C).$$

**Proof.** We proceed by induction on $n$ and first consider $n = 1$. Thus, we reason in $T$ and assume $U_1$, that is, $\Diamond \neg (D_1 \triangleright \neg C)$. We conclude $\Diamond \forall \text{Cut} \kappa \Diamond (D_1 \land \Box^K C)$, whence $\forall \text{Cut} \kappa \Diamond (D_1 \land \Box^K C)$ and also $\forall \text{Cut} \kappa \Diamond (D_1 \land \Box^K C)$ as was to be shown.
Next, we consider the inductive case, again reasoning in $T$ and assuming
$U_{n+1}$ which is $\Diamond((D_n \triangleright D_{n+1}) \land U_n)$. By the (IH) we conclude from $U_n$ that

$$\forall \text{Cut} \ J \ (D_n \land \Box^j C). \quad (20)$$

By Lemma 2.8 we obtain from $D_n \triangleright D_{n+1}$ that

$$\forall \text{Cut} \ K \exists \text{Cut} \ J \ D_n \land \Box^j C \triangleright D_{n+1} \land \Box^K C. \quad (21)$$

Combining $D_n \land \Box^j C \triangleright D_{n+1} \land \Box^K C \rightarrow (\Diamond(D_n \land \Box^j C) \rightarrow \Diamond(D_{n+1} \land \Box^K C))$ with (20) and (21) under a $\Diamond$ we conclude that

$$\Diamond((D_n \triangleright D_{n+1}) \land U_n) \rightarrow \Diamond(\forall \text{Cut} \ K \Diamond(D_{n+1} \land \Box^K C)) \rightarrow \forall \text{Cut} \ K \Diamond(\Diamond(D_{n+1} \land \Box^K C)) \rightarrow \forall \text{Cut} \ K \Diamond(\Diamond(D_{n+1} \land \Box^K C))$$

as was to be shown. □

With this lemma, we can now prove the soundness of the series $R^n$.

**Theorem 4.8.** For each natural number $n$ we have that $R^n$ is arithmetically sound in any theory $T$ extending $S^1_2$.

**Proof.** Since we already know that $R^0$ is sound, we consider $n > 0$. We reason in $T$, assume $A \triangleright B$ and set out to prove $U_n \land (D_n \triangleright A) \triangleright B \land \Box C$. By Pudlák’s Lemma we get

$$\exists \text{Cut} \ J \ A \land \Box^j C \triangleright B \land \Box C. \quad (22)$$

On the other hand, by the generalization of Pudlák’s Lemma (Lemma 2.8) applied to $D_n \triangleright A$ we obtain that $\forall \text{Cut} \ J \exists \text{Cut} \ K \ D_n \land \Box^K C \triangleright A \land \Box^j C$ so that $\forall \text{Cut} \ J \exists \text{Cut} \ K \ (\Diamond(D_n \land \Box^K C) \rightarrow \Diamond(A \land \Box^j C))$. By Lemma 4.7 we see that $U_n \rightarrow \forall \text{Cut} \ K \Diamond(A \land \Box^K C)$. Combining these last two observations, we see that $U_n \land (D_n \triangleright A) \rightarrow \forall \text{Cut} \ J \Diamond(A \land \Box^j C)$ so that $\forall \text{Cut} \ J \ U_n \land (D_n \triangleright A) \triangleright A \land \Box^j C$. Combining this with (22) yields $U_n \land (D_n \triangleright A) \triangleright B \land \Box C$ as was to be shown. □

5 **On the core interpretability logic IL(All)**

Apart from the principles mentioned earlier in this paper the literature has considered various other principles too. Some of those are

- **W:** $A \triangleright B \rightarrow A \triangleright B \land \Box \neg A$
- **W**\*: $A \triangleright B \rightarrow B \land \Box C \triangleright B \land \Box C \land \Box \neg A$
- **P**\_0: $A \triangleright \Diamond B \rightarrow \Box(A \triangleright B)$
- **R:** $A \triangleright B \rightarrow \neg(A \triangleright \neg C) \triangleright B \land \Box C$
In [11], IL(All) was conjectured to be ILW. In [13] this conjecture was falsified and strengthened to a new conjecture, namely that ILW∗, which is a proper extension of ILW, is IL(All). In [8] it was proven that the logic ILW∗P0 is a proper extension of ILW∗, and that ILW∗P0 is a subsystem of IL(All) (we write IL{W∗, P0} instead of ILW∗P0). This falsified the conjecture from [13].

In [8] it is also conjectured that ILW∗P0 is not the same as IL(All).

In [7] it is conjectured that ILW∗P0 = IL(All) and this conjecture was refuted in [4] by proving that the logic ILRW is a subsystem of IL(All) and a proper extension of ILW∗P0.

It is easy to see that A ⊲ ♦ B → □(A ⊲ ♦ B) ∈ ILP ∩ ILM. In [14] it was shown however that A ⊲ ♦ B → □(A ⊲ ♦ B) /∈ IL(All) thereby lowering the upper bound IL(All) ⊆ ILP ∩ ILM. Since A ⊲ ♦ B → □(A ⊲ ♦ B) is reminiscent of the modally incomplete principle P0, we remark here that the principle

A ⊲ ♦ B → ¬(A ⊲ ♦ C) ⊲ B ∧ □¬C

implies A ⊲ ♦ B → □¬(A ⊲ ♦ B) so that it cannot be in IL(All) either.

The current paper raises the previously known lower bound of IL(All). However, it seems unlikely that this will be the end of the story and the two series presented here seem amenable for interactions. Just by mere inspection of the frame conditions we observe that

F_n = ∀w, x, y, z (B_0(w, x, y, z) ⇒ G_n(x, y, z)),
F^n = ∀w, x, y, z (B_n(w, x, y, z) ⇒ G_0(x, y, z)).

suggesting possible interactions. For example, a combination of R^1 and R_1 could yield

A ⊲ B → (C ⊲ A) ∧ ◊¬(C ⊲ ¬D) ∧ (E ⊲ ◊F) ⊲ B ∧ □D ∧ (E □ F).

We note that the two series presented in this paper only spoke of S relations that were imposed by the frame conditions. This suggests that a new conjecture can be formulated.

Let F be a class of IL-frame. By IL[F] we shall denote the interpretability logic corresponding to this class. That is,

IL[F] := \{ A | ∀F ∈ F \forall \text{valuation} V \langle F, V \rangle \models A}.\)

We now define the class of frames All to be the set of frames where any S relation that is implied both by the ILM and the ILP frame condition is present. To make this more precise, let P denote the first-order frame condition of P and let M denote the first-order frame condition of M. Let F(x, y, z) denote any sentence – first or higher order – in the language \{R, \{S_x\}_{x ∈ W}\}. We write ILP ⊨ F(x, y, z) → yS_xz to denote that for any Veltman frame F for which \langle F, V \rangle we also have \langle F, V \rangle \models F(x, y, z) → yS_xz. Likewise, we shall speak of
\[
\text{ILM} \models F(x, y, z) \to yS_x z. \text{ With this notation, we define }
\]

\[
\mathbb{A} := \{ F | \left( \text{ILP} \models (F(x, y, z) \to yS_x z) \land \text{ILM} \models (F(x, y, z) \to yS_x z) \Rightarrow F \models (F(x, y, z) \to yS_x z) \right) \}.
\]

The second author poses the new conjecture

**Conjecture 5.1.** \ IL(\mathbb{A}) = \ IL[\mathbb{A}] .

It is easy to formulate the conjecture where the antecedent \( F(x, y, z) \) is replaced by a set of sentences rather than a single sentence yet it seems hard to imagine that this is needed. Note that the conjecture only speaks of principles related to imposed \( S \) relations. For example, this will leave out a principle like \( A \triangleright B \rightarrow (\Diamond A \land \Box \Box C \triangleright B \land \Box C) \) as formulated in [7].

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