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Infinitary and Cyclic Proof Systems for Transitive Closure Logic

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Abstract. We formulate an infinitary proof system for transitive closure logic, which is the logic obtained by augmenting first-order logics with a transitive closure operator. Our system is an infinite descent-style counterpart to the existing (explicit induction) proof system for the logic. We show that, as for similar systems for first order logic with inductive definitions, our infinitary system is complete for the standard semantics and subsumes the explicit system. Moreover, the uniformity of the transitive closure operator allows semantically meaningful complete restrictions to be defined using simple syntactic criteria. Consequently, the restriction to regular infinitary (i.e. cyclic) proofs provides the basis for an effective system for automating inductive reasoning.

1 Introduction

A core technique in mathematical reasoning is that of induction. This is especially true in computer science, where it plays a central role in reasoning about recursive data and computations. Formal systems for mathematical reasoning usually capture the notion of inductive reasoning via one or more inference rules that express the general induction schemes, or principles, that hold for the elements being reasoned over.

Increasingly, we are concerned with not only being able to formalise as much mathematical reasoning as possible, but also with doing so in an effective way. In other words, we seek to be able to automate such reasoning. Transitive closure (TC) logic has been identified as a potential candidate for a minimal, ‘most general’ system for inductive reasoning, which is also very suitable for automation\(^3\). TC adds to first order logic a single operator for forming binary relations: specifically, the transitive closures of arbitrary formulas. This modest addition affords enormous expressive power: namely it provides a uniform way of capturing inductive principles. If an induction scheme is expressed by a formula \(\varphi\), then the elements of the inductive collection it defines are those ‘reachable’ from the base elements \(x\) via the iteration of the induction scheme. That is, those \(y\)’s for which \((x, y)\) is in the transitive closure of \(\varphi\). Thus, bespoke induction principles do not need to be added to, or embedded within, the logic; instead, all induction schemes are available within a single, unified language.

\(^3\) In this work, for simplicity, we use a reflexive form of the operator.
Although the expressiveness of TC logic renders any effective proof system for it incomplete for the standard semantics, a natural, effective proof system which is sound for TC logic was shown to be complete with respect to a generalized form of Henkin-semantics [10]. In this paper, following similar developments in other formalizations for fixed point logics and inductive reasoning (see e.g. [5,6,17,19]), we present an infinitary proof theory for TC logic which, as far as we know, is the first system that is (cut-free) complete with respect to the standard semantics. The soundness of such infinitary proof theories is underpinned by the principle of infinite descent: proofs are permitted to be infinite (i.e. non-well-founded) trees, but subject to the restriction that every infinite path in the proof admits some infinite descent. The descent is witnessed by tracing terms or formulas for which we can give a correspondence with elements of a well-founded set. In the context of formalised induction, we can use formulas interpreted by the elements of inductive collections. For this reason, such theories are considered systems of implicit induction, as opposed to those which employ explicit rules for applying induction principles. While a full infinitary proof theory is clearly not effective, in the aforementioned sense, such a system can be obtained by restricting consideration to only the regular infinite proofs. These are precisely those proofs that can be finitely represented as (possibly cyclic) graphs.

These infinitary proof theories generally subsume systems of explicit induction in expressive power, but also offer a number of advantages. Most notably, they can ameliorate the primary challenge for inductive reasoning: finding an induction invariant. In explicit induction systems, this must be provided a priori, and is often much stronger than the goal one is ultimately interested in proving. However, in implicit systems the inductive arguments and hypotheses are encoded in the cycles of a proof, so cyclic proof systems seem better for automation. The cyclic approach has also been used to provide an optimal cut-free complete proof system for Kleene algebra [12], providing further evidence of its utility for automation.

In the setting of TC logic, we observe some further benefits over more traditional formal systems of inductive definitions and their infinitary proof theories (cf. LKID [6,15]). As previously mentioned, TC (with a pairing function) has all inductive definitions immediately ‘available’ within the language of the logic: as with inductive hypotheses, one does not need to ‘know’ in advance which induction schemes will be required. Moreover, the use of a single transitive closure operator provides a uniform treatment of all induction schemes. That is, instead of having a proof system parameterized by a set of inductive predicates and rules for them (as is the case in LKID), TC offers a single proof system with a single rule scheme for induction. This has immediate benefits in developing the metatheory: the proofs of completeness w.r.t. standard semantics and adequacy (i.e. subsumption of explicit induction) for the infinitary system presented in this paper are simpler and more straightforward. Furthermore, it allows a simple syntactic criterion (which we call normality) to define a cyclic subsystem that is complete for Henkin semantics. This suggests the possibility of more focussed proof-search strategies, further enhancing the potential for automation. TC logic seems more expressive in other ways, too. The transitive closure operator may
be applied to any formula, thus we are not restricted to induction principles corresponding only to monotone generation schemes (as in, e.g., [46]).

We show that the explicit and cyclic TC systems are equivalent under arithmetic, as is the case for LKID [31,38]. However, there are cases in which the cyclic system for LKID is strictly more expressive than the explicit induction system [2]. To obtain a similar result for TC, the fact that all induction schemes are available poses a serious challenge. For one, the counter-example used in [2] does not serve to show this result holds for TC. If this strong inequivalence indeed holds also for TC, it must be witnessed by a more subtle and complex counter-example. Conversely, it may be that the explicit and cyclic systems do coincide for TC. In either case, this points towards fundamental aspects that require further investigation.

The rest of the paper is organised as follows. In Section 2 we reprise the definition of transitive closure logic and both its standard and Henkin-style semantics. Section 3 presents the existing explicit induction proof system for TC logic, and also our new infinitary proof system. We prove the latter sound and complete for the standard semantics, and also derive cut-admissibility. In Section 4 we compare the expressive power of the infinitary system (and its cyclic subsytem) with the explicit system. Section 5 concludes and examines the remaining open questions for our system as well as future work.

2 Transitive Closure Logic and its Semantics

In this section we review the language of transitive closure logic, as well as two possible semantics for it: a standard one, and a Henkin-style one.

Definition 1 (The language \( \mathcal{L}_{RTC} \)). Let \( \sigma \) be some first-order signature, and let \( \mathcal{L} \) be the corresponding first-order language. The language \( \mathcal{L}_{RTC} \) is obtained from \( \mathcal{L} \) by the addition of the reflexive transitive closure operator (RTC), together with the following clause concerning the definition of a formula:

\[ (RTC_{x,y} \varphi)(s,t) \]

is a formula in \( \mathcal{L}_{RTC} \) for any formula \( \varphi \) in \( \mathcal{L}_{RTC} \), distinct variables \( x, y \), and terms \( s, t \). (The free occurrences of \( x \) and \( y \) in \( \varphi \) become bound in this formula.)

Definition 2 (Standard Semantics). Let \( M \) be a structure for \( \mathcal{L}_{RTC} \), and \( v \) an assignment in \( M \). The (standard) semantics of \( \mathcal{L}_{RTC} \) is defined as classical first-order logic, with the following additional clause for the satisfaction relation:

\[ (RTC_{x,y} \varphi)(s,t), \text{ if } v(s) = v(t), \text{ or there exist } a_0, \ldots, a_n \in D (n > 0) \text{ s.t. } v(s) = a_0, v(t) = a_n, \text{ and } M, v[x := a_i, y := a_{i+1}] \models \varphi \text{ for } 0 \leq i < n. \]

We next recall the concepts of frames and Henkin structures (see, e.g. [14]). A frame is a relational structure together with some subset of the powerset of its domain (called its set of admissible subsets).
Definition 3 (Frames). Let $\sigma$ be a first-order signature. A $\sigma$-frame $M$ is a triple $\langle D, I, D' \rangle$, where $D$ is a non-empty domain, $I$ is an interpretation function on $\sigma$ in $D$, and $D' \subseteq \wp(D)$.

Note that if $D' = \wp(D)$, the $\sigma$-frame is identified with a standard structure.

Definition 4 (Frame Semantics). Let $\mathcal{L}_{RTC}$ be the language based on the first-order signature $\sigma$. $\mathcal{L}_{RTC}$ formulas are interpreted in $\sigma$-frames as in standard structures, except for the following clause concerning the RTC operator:

$- M, v \models (RTC_{x,y}\varphi)(s,t)$ if for every $A \in D'$, if $v(s) \in A$ and for every $a,b \in D$: $a \in A$ and $M, v[x := a, y := b] \models \varphi$ implies $b \in A$, then $v(t) \in A$.

We now restrict our set of structures to $\sigma$-Henkin structures, which are $\sigma$-frames whose set of admissible subsets satisfies some closure conditions.

Definition 5 (Henkin structures). Let $\mathcal{L}_{RTC}$ be the language based on $\sigma$. A $\sigma$-Henkin structure is a $\sigma$-frame $M$ closed under parametric definability, i.e., for each $\mathcal{L}_{RTC}$ formula $\varphi$ and assignment $v$ in $M$, $\{a \in D \mid M, v[x := a] \models \varphi\} \in D'$.

We refer to the logical validity induced by considering only Henkin structures as the Henkin semantics. Note that under both the standard semantics and Henkin semantics for the RTC operator, equality is definable as follows:

$$s = t := (RTC_{x,y}\bot)(s,t).$$

Thus, we do not include it explicitly in our logical languages.

To obtain the full inductive expressivity we must allow the formation of the transitive closure of not only binary relations, but any $2n$-ary relation. In [1] it was shown that taking such a $RTC^n$ operator for every $n$ (instead of just for $n = 1$) results in a more expressive logic, namely one that captures all finitary inductive definitions and relations. Since having just one $RTC$ operator is more convenient from a proof theoretical point of view, we here instead incorporate the notion of ordered pairs and use it to encode such operators. For example, writing $\langle x, y \rangle$ for the application of the pairing function $\langle \rangle$, the formula $(RTC^2_{x_1,x_2,y_1,y_2}\varphi)(s_1,s_2,t_1,t_2)$ can be encoded by:

$$(RTC_{x,y}\exists x_1,x_2,y_1,y_2 . x = \langle x_1, x_2 \rangle \wedge y = \langle y_1, y_2 \rangle \wedge \varphi((s_1,s_2),(t_1,t_2)).$$

Accordingly, we assume languages that explicitly contain a pairing function, and restrict to structures that interpret it as such (i.e. the admissible structures).

Definition 6 (Admissible Structures). Let $\sigma$ be a signature that contains some constant $c$, and a binary function symbol, denoted by $\langle \rangle$. Let $\mathcal{L}_{RTC}$ be a language based on $\sigma$. A structure $M$ for $\mathcal{L}_{RTC}$ is called admissible if:

$$M \models \langle x_1, x_2 \rangle = \langle y_1, y_2 \rangle \rightarrow x_1 = y_1 \wedge x_2 = y_2 \quad (2)$$

$$M \models \langle x, y \rangle \neq c \quad (3)$$

For such languages we consider two induced semantics: admissible standard semantics and admissible Henkin semantics, obtained by restricting the (first-order part of the) structures to be admissible.
3 Proof Systems for $\mathcal{L}_{\text{RTC}}$

In this section, we define two proof systems for $\mathcal{L}_{\text{RTC}}$. The first is a finitary proof system with an explicit induction rule for $\text{RTC}$ formulas. The second is an infinitary proof system, in which $\text{RTC}$ formulas are simply unfolded, and inductive arguments are represented via infinite descent-style constructions. We show the soundness and completeness of these proof systems, and also compare their provability relations.

The systems for $\mathcal{L}_{\text{RTC}}$ defined below are extensions of the sequent calculus for classical first-order logic, $\mathcal{L}K$ (see, e.g., [13,20]). A sequent is an expression of the form $\Gamma \Rightarrow \Delta$, where we here take $\Gamma$ and $\Delta$ to be finite sets of formulas. Here we also take $\mathcal{L}K$ to include the substitution rule, but we do not include the following standard rules for equality since, as we show below, they are admissible in our systems:

$\Gamma \Rightarrow t \equiv t, \Delta$

(paramodulation): $\Gamma \Rightarrow \varphi \{ \frac{s}{x}, \frac{t}{y} \}, \Delta$

$\Gamma, s \equiv t \Rightarrow \varphi \{ \frac{t}{x}, \frac{s}{y} \}, \Delta$

$\Gamma, \psi \{ \frac{s}{x}, \frac{t}{y} \}, \Delta \Rightarrow \Delta, \psi \{ \frac{t}{x} \}$

(4)

$\Gamma \Rightarrow \Delta, (\text{RTC}_{x,y} \varphi)(s,t)$

$\Gamma \Rightarrow \Delta, (\text{RTC}_{x,y} \varphi)(s,r)$

$\Gamma \Rightarrow \Delta, \varphi \{ \frac{x}{x}, \frac{y}{y} \}$

$\Gamma \Rightarrow \Delta, (\text{RTC}_{x,y} \varphi)(s,t)$

(5)

$\Gamma, \psi(x), \varphi(x,y) \Rightarrow \Delta, \psi \{ \frac{s}{x} \}$

$\Gamma, \psi \{ \frac{s}{x}, \frac{t}{y} \}, (\text{RTC}_{x,y} \varphi)(s,t) \Rightarrow \Delta, \psi \{ \frac{t}{x} \}$

(6)

In all the rules we assume that the terms which are substituted are free for substitution, and that no forbidden capturing occurs. In Rule (6), $x$ should not occur free in $\Gamma$ or $\Delta$, and $y$ should not occur free in $\Gamma, \Delta$ or $\psi$.

Recall that $s = t$ is definable in $\mathcal{L}_{\text{RTC}}$, vide Eq. (1). Moreover, in $\mathcal{L}_{\text{RTC}}$ the rules for equality are admissible since they are derivable using the defining formula. To see this, notice that $s = s$ stands for $(\text{RTC}_{x,y} \bot)(s,s)$ which is provable using Rule (4). To derive the Paramodulation rule we construct the following derivation in $\mathcal{L}_{\text{RTC}}$:

$\Gamma \Rightarrow \Delta, \psi \{ \frac{s}{x} \}$

$\Gamma, \psi \{ \frac{s}{x}, \frac{t}{y} \}, (\text{RTC}_{x,y} \bot)(s,t) \Rightarrow \Delta, \psi \{ \frac{t}{x} \}$

$\Gamma, (\text{RTC}_{x,y} \bot)(s,t) \Rightarrow \Delta, \psi \{ \frac{t}{x} \}$

(6)

(4)

(5)

(6)

(5)

(6)

(5)

(6)

(5)

(6)
3.2 Infinitary Proof Systems

**Definition 8.** The infinitary proof system \( \text{RTC}_G \) for \( L_{\text{RTC}} \) is defined like \( \text{RTC}_G \), but replacing Rule (6) by:

\[
\frac{\Gamma \left\{ \frac{s}{x}, \frac{t}{w} \right\} \Rightarrow \Delta' \left\{ \frac{s}{x}, \frac{t}{w} \right\}}{\Gamma, (\text{RTC}_{x,y} \varphi)(s, z), \varphi \left\{ \frac{z}{x}, \frac{t}{y} \right\} \Rightarrow \Delta}
\]

(7)

where \( z \) is fresh, i.e. \( z \) does not occur free in \( \Gamma, \Delta \), or \( (\text{RTC}_{x,y} \varphi)(s, t) \); and \( \Gamma' \) and \( \Delta' \) are such that \( \Gamma = \Gamma' \left\{ \frac{z}{x}, \frac{t}{w} \right\} \) and \( \Delta = \Delta' \left\{ \frac{z}{x}, \frac{t}{w} \right\} \). We call the formula \( (\text{RTC}_{x,y} \varphi)(s, z) \) in the right-hand premise the ancestor of the principle formula, \( (\text{RTC}_{x,y} \varphi)(s, t) \), in the conclusion.

There is an asymmetry between Rule (5), in which the intermediary is an arbitrary term \( r \), and Rule (7), where we use a variable \( z \). This is necessary to obtain the soundness of the cyclic proof system. It is used to show that when there is a counter-model for the conclusion of a rule, then there is also a counter-model for one of its premises that is, in a sense that we make precise below, ‘smaller’. Using a fresh \( z \) allows us to pick from all counter-models of the premise, whereas if we allowed an arbitrary term \( r \) instead, this might restrict the counter-models we can choose from, i.e. it might only leave ones larger than the smallest one for the conclusion. See Lemma [1] below for more details.

As for the finitary system, the rules for equality are admissible. The following equivalent formulation of the paramodulation rule (used in, e.g., [4,5])

\[
\frac{\Gamma \left\{ \frac{s}{x}, \frac{t}{y} \right\} \Rightarrow \Delta \left\{ \frac{s}{x}, \frac{t}{y} \right\}}{\Gamma \left\{ \frac{t}{x}, \frac{s}{y} \right\}, s = t \Rightarrow \Delta \left\{ \frac{t}{x}, \frac{s}{y} \right\}}
\]

(=L)

is subsumed by Rule (7), of which it can easily be seen to be a special case:

\[
\frac{\Gamma \left\{ \frac{s}{x}, \frac{t}{y} \right\} \Rightarrow \Delta \left\{ \frac{s}{x}, \frac{t}{y} \right\}}{\Gamma \left\{ \frac{t}{x}, \frac{s}{y} \right\}, (\text{RTC}_{x,y} \perp)(s, z), \perp \Rightarrow \Delta \left\{ \frac{t}{x}, \frac{s}{y} \right\}}
\]

(=L')

Proofs in this system are possibly infinite derivation trees. However, not all infinite derivations are proofs: only those that admit an infinite descent argument. Thus we use the terminology ‘pre-proof’ for derivations.

**Definition 9 (Pre-proofs).** An \( \text{RTC}_G^\omega \) pre-proof is a possibly infinite (i.e. non-well-founded) derivation tree formed using the inference rules. A path in a pre-proof is a possibly infinite sequence of sequents \( s_0, s_1, \ldots, (s_n) \) such that \( s_0 \) is the root sequent of the proof, and \( s_{i+1} \) is a premise of \( s_i \) for each \( i < n \).

The following definitions tell us how to track \( \text{RTC} \) formulas through a pre-proof, and allow us to formalize inductive arguments via infinite descent.
Definition 10 (Trace Pairs). Let \( \tau \) and \( \tau' \) be RTC formulas occurring in the left-hand side of the conclusion \( s \) and a premise \( s' \), respectively, of (an instance of) an inference rule. \((\tau, \tau')\) is said to be a trace pair for \((s, s')\) if the rule is:

- the (Subst) rule, and \( \tau = \tau' \theta \) where \( \theta \) is the substitution associated with the rule instance;
- Rule (7), and either:
  a) \( \tau \) is the principle formula of the rule instance and \( \tau' \) is the ancestor of \( \tau \), in which case we say that the trace pair is progressing;
  b) \( s' \) is the left-hand premise and \( \tau' = \tau'' \left\{ \frac{s}{v}, \frac{t}{w} \right\} \) for some \( \tau'' \in \Gamma' \) or
  c) \( s' \) is the right-hand premise and \( \tau = \tau' \).
- any other rule, and \( \tau = \tau' \).

Definition 11 (Traces). A trace is a (possibly infinite) sequence of RTC formulas. We say that a trace \( \tau_1, \tau_2, \ldots, \tau_n \) follows a path \( s_1, s_2, \ldots, s_n \) in a pre-proof \( P \) if, for some \( k \geq 0 \), each consecutive pair of formulas \((\tau_i, \tau_{i+1})\) is a trace pair for \((s_i+k, s_{i+k+1})\). If \((\tau_i, \tau_{i+1})\) is a progressing pair then we say that the trace progresses at \( i \), and we say that the trace is infinitely progressing if it progresses at infinitely many points.

Proofs, then, are pre-proofs which satisfy a global trace condition.

Definition 12 (Infinite Proofs). A RTC\(\omega\) proof is a pre-proof in which every infinite path is followed by some infinitely progressing trace.

Clearly, we cannot reason effectively about such infinite proofs in general. In order to do so we need to restrict our attention to those proof trees which are finitely representable. These are the regular infinite proof trees, which contain only finitely many distinct subtrees. They can be finitely represented as systems of recursive equations or, alternatively, as cyclic graphs [11]. Note that a given regular infinite proof may have many different graph representations. One possible way of formalizing such proof graphs is as standard proof trees containing open nodes (called buds), to each of which is assigned a syntactically equal internal node of the proof (called a companion). Due to space limitation, we elide a formal definition of cyclic proof graphs and rely on the reader’s basic intuitions.

Definition 13 (Cyclic Proofs). The cyclic proof system CRTC\(\omega\) for \( L_{RTC} \) is the subsystem of RTC\(\omega\) comprising of all and only the finite and regular infinite proofs (i.e. those proofs that can be represented as finite, possibly cyclic, graphs).

Note that it is decidable whether a cyclic pre-proof satisfies the global trace condition, using a construction involving Büchi automata (see, e.g., [118]).

\[\text{Here, } \Gamma', v, w, s \text{ and } t \text{ refer to the instantiations of these same meta-variables appearing in the schema of Rule (7)}.\]
3.3 Soundness and Completeness

The rich expressiveness of TC logic entails that the effective system RTC\(_G\) which is sound w.r.t. the standard semantics, cannot be complete (much like the case for LKID). It is however both sound and complete w.r.t. Henkin semantics.

Theorem 1 (Soundness and Completeness of RTC\(_G\) [10]). RTC\(_G\) is sound for standard semantics, and also sound and complete for Henkin semantics.

We note that the soundness proof of LKID is rather complex since it must handle different types of mutual dependencies between the inductive predicates. For RTC\(_G\) the proof is much simpler due to the uniform rules for the RTC operator.

The infinitary system RTC\(_G^\omega\), in contrast to the finitary system RTC\(_G\), is both sound and complete w.r.t. the standard semantics. To prove soundness, we make use of the following notion of measure for RTC formulas.

Definition 14 (Degree of RTC Formulas). For \(\phi \equiv (RTC_{x,y} \varphi)(s,t)\), we define \(\delta_\phi(M,v) = 0\) if \(v(s) = v(t)\), and \(\delta_\phi(M,v) = n\) if \(v(s) \neq v(t)\) and \(a_0, \ldots, a_n\) is a minimal-length sequence of elements in the semantic domain \(D\) such that \(v(s) = a_0, v(t) = a_n\), and \(M, v[x := a_i, y := a_{i+1}] \models \varphi\) for \(0 \leq i < n\). We call \(\delta_\phi(M,v)\) the degree of \(\phi\) with respect to the model \(M\) and valuation \(v\).

Soundness then follows from the following fundamental lemma.

Lemma 1 (Descending Counter-models). If there exists a standard model \(M\) and valuation \(v\) that invalidates the conclusion \(s\) of (an instance of) an inference rule, then 1) there exists a standard model \(M'\) and valuation \(v'\) that invalidates some premise \(s'\) of the rule; and 2) if \((\tau, \tau')\) is a trace pair for \((s, s')\) then \(\delta_{\tau'}(M', v') \leq \delta_\tau(M, v)\). Moreover, if \((\tau, \tau')\) is a progressing trace pair then \(\delta_{\tau'}(M', v') < \delta_\tau(M, v)\).

Proof. The cases for the standard \(\mathcal{LK}\) rules are straightforward adaptations of those found in e.g. [6].

- The case for Rule (4) follows trivially since it follows immediately from Definition 2 that \(M, v \models (RTC_{x,y} \varphi)(s, s)\) for all \(M\) and \(v\).

- For Rule (5), we take the right-hand premise with model \(M' = M\) and valuation \(v' = v\). Thus, \(M', v' \models \psi\) and \(M', v' \not\models \phi\) for all \(\psi \in \Gamma\) and \(\phi \in \Delta\); moreover, the degree of all RTC formulas in \(\Gamma\) remains unchanged with respect to \(M'\) and \(v'\). Now, since \(M, v \not\models (RTC_{x,y} \varphi)(s, t)\), it follows from Definition 2 that there are no \(a, a' \in D\) such that \(M, v[x := a, y := a'] \models \varphi\). This then gives that \(M', v' \not\models \varphi\).

- For Rule (7), since \(M, v \models (RTC_{x,y} \varphi)(s, t)\) there are two cases to consider:
  (i) If \(v(s) = v(t)\) then we take the left-hand premise with model \(M' = M\) and valuation \(v' = v\), since then swapping occurrences of \(s\) and \(t\) in \(\Gamma\) and \(\Delta\) preserves satisfiability and the degree of any RTC formula in \(\Gamma'\) with respect to \(M'\) and \(v'\) remains the same.
(ii) If on the other hand there are \( a_0, \ldots, a_n \in D \) \( (n > 0) \) such that \( v(s) = a_0 \) and \( v(t) = a_n \) with \( M, v[x := a_i, y := a_{i+1}] \models \varphi \) for \( 0 \leq i < n \), we then take the right-hand premise, the model \( M' = M \) and valuation \( v' = v[z := a_{n-1}] \). Note that, without loss of generality, we may assume a sequence \( a_0, \ldots, a_n \) of minimal length, and thus surmise \( \delta_{\text{RTC}_x, y}(\varphi)(s, t)(M, v) = n \). Since \( z \) is fresh, it follows that \( M', v' \models \varphi \left\{ \frac{z}{s}, \frac{t}{y} \right\} \) and \( M', v'[x := a_i, y := a_{i+1}] \models \varphi \) for \( 0 < i < n - 1 \). If \( n = 1 \) then \( v'(s) = v'(z) = a_0 \) and so \( M, v' \models (\text{RTC}_{x, y}(\varphi))(s, z) \); otherwise this is witnessed by the sequence \( a_0, \ldots, a_{n-1} \). Thus we also have that \( \delta_{\text{RTC}_x, y}(\varphi)(s, z)(M', v') = n - 1 \). To conclude, note it also follows from \( z \) fresh that \( M', v' \models \psi \) for all \( \psi \in \Gamma \) and \( M', v' \not\models \phi \) for all \( \phi \in \Delta \); and furthermore that the degree of any RTC formula in \( \Gamma \) remains unchanged with respect to \( M' \) and \( v' \).

As is standard for infinite descent inference systems \([4,6,12,16,21]\), the above result entails the local soundness of the inference rules (in our case, for standard first-order models). The presence of infinitely progressing traces for each infinite path in a RTC\(^c\) proof ensures soundness via a standard infinite descent-style construction.

**Theorem 2 (Soundness of RTC\(^c\))**. If there is a RTC\(^c\) proof of \( \Gamma \Rightarrow \Delta \), then \( \Gamma \Rightarrow \Delta \) is valid (w.r.t. the standard semantics).

**Proof**. Suppose, for contradiction, that \( \Gamma \Rightarrow \Delta \) is not valid. Then by Lemma 1 there exists an infinite path \( \{s_i\}_{i > 0} \) in the proof and an infinite sequence of model-valuation pairs \( \{(M_i, v_i)\}_{i > 0} \) such that \( (M_i, v_i) \) invalidates \( s_i \) for each \( i > 0 \). Since the proof is a valid RTC\(^c\) proof, this infinite path is followed by an infinitely progressing trace \( \{\tau_i\}_{i > 0} \) for which we can take the degree of each formula with respect to its corresponding counter-model to obtain an infinite sequence of natural numbers \( \{\delta_{\text{RTC}_x, y}(\varphi)(\tau_i)(s, z)\}_{i > 0} \) (for some \( k \geq 0 \)). By Lemma 1 this sequence is decreasing and, moreover, since the trace is infinitely progressing the sequence strictly decreases infinitely often. From the fact that the natural numbers are a well-founded set we derive a contradiction, and thus conclude that \( \Gamma \Rightarrow \Delta \) is indeed valid. \( \square \)

The soundness of the cyclic system is an immediate corollary, since each CRT\(c\) proof is also a RTC\(^c\) proof.

**Corollary 1 (Soundness of CRT\(c\))**. If there is a CRT\(c\) proof of \( \Gamma \Rightarrow \Delta \), then \( \Gamma \Rightarrow \Delta \) is valid (w.r.t. the standard semantics).

Following a standard technique (as used in e.g. \([8]\)), we can show cut-free completeness of RTC\(^c\) with respect to the standard semantics.

**Definition 15 (Schedule)**. A schedule element \( E \) is of the following form:

- a formula of the form \( \neg \varphi, \varphi \land \psi, \varphi \lor \psi, \varphi \rightarrow \psi \).
- a pair of the form \( \langle \forall x \varphi, t \rangle \) or \( \langle \exists x \varphi, t \rangle \), where \( \forall x \varphi \) and \( \exists x \varphi \) are formulas and \( t \) is a term.
The limit of the sequence \((RTC_{x,y}(s,t), r, v, w, z, \Gamma, \Delta)\) where \((RTC_{x,y}(s,t))\) is a formula; \(r\) is a term; \(v, w,\) and \(z\) are variables (with \(v\) and \(w\) distinct) that do not occur free in \((RTC_{x,y}(s,t))\), and \(z\) does not occur free in either \(\Gamma\) or \(\Delta\); and \(\Gamma\) and \(\Delta\) are finite sequences of formulas.

A schedule is a recursive enumeration of schedule elements in which every schedule element appears infinitely often (these exist since our language is countable).

Each schedule corresponds to an exhaustive search strategy for a cut-free proof for each sequent \(\Gamma \Rightarrow \Delta\), via the following notion of a ‘search tree’.

**Definition 16 (Search Tree).** Given a schedule \(\{E_i\}_{i \geq 0}\), for each sequent \(\Gamma \Rightarrow \Delta\) we inductively define an infinite sequence of (possibly open) derivation trees, \(\{T_i\}_{i \geq 0}\), such that \(T_0\) consists of the single open node \(\Gamma \Rightarrow \Delta\), and each \(T_{i+1}\) is obtained by replacing all suitable open nodes in \(T_i\) with applications of first axioms and then the left and right inference rules for the formula in the \(i\)th schedule element.

We give the definition of \(T_{i+1}\) when \(E_i\) is an RTC schedule element, i.e. of the form \((RTC_{x,y}(s,t), r, v, w, z, \Gamma, \Delta)\) (the other cases are similar). \(T_{i+1}\) is then obtained by:

1. first closing as such any open node that is an instance of an axiom;
2. next, replacing every open node \(\Gamma'\), \((RTC_{x,y}(s,t))\Rightarrow \Delta'\) of the resulting tree for which \(\Gamma' \subseteq \Gamma\{\frac{t}{v}, \frac{z}{w}\}\) and \(\Delta' \subseteq \Delta\{\frac{t}{v}, \frac{z}{w}\}\) with the derivation:

\[
\begin{align*}
\Gamma'' \{\frac{t}{v}, \frac{z}{w}\}, \Gamma'' \{\frac{t}{v}, \frac{z}{w}\}, \Gamma'' \{\frac{t}{v}, \frac{z}{w}\} & \Rightarrow \Delta'' \{\frac{t}{v}, \frac{z}{w}\} \\
\Gamma'' \{\frac{t}{v}, \frac{z}{w}\}, \Gamma'' \{\frac{t}{v}, \frac{z}{w}\}, \Gamma'' \{\frac{t}{v}, \frac{z}{w}\} & \Rightarrow \Delta'' \{\frac{t}{v}, \frac{z}{w}\} \\
\Gamma'' \{\frac{t}{v}, \frac{z}{w}\}, \Gamma'' \{\frac{t}{v}, \frac{z}{w}\}, \Gamma'' \{\frac{t}{v}, \frac{z}{w}\} & \Rightarrow \Delta'' \{\frac{t}{v}, \frac{z}{w}\}
\end{align*}
\]

where \(\Gamma'' \subseteq \Gamma\) and \(\Delta'' \subseteq \Delta\) are such that \(\Gamma' = \Gamma'' \{\frac{t}{v}, \frac{z}{w}\}\) and \(\Delta' = \Delta'' \{\frac{t}{v}, \frac{z}{w}\}\).
3. finally, replacing every open node \(\Gamma' \Rightarrow \Delta'\), \((RTC_{x,y}(s,t))\) of the resulting tree with the derivation:

\[
\begin{align*}
\Gamma' \Rightarrow \Delta', \Gamma'' \{\frac{t}{v}, \frac{z}{w}\}, \Gamma'' \{\frac{t}{v}, \frac{z}{w}\}, \Gamma'' \{\frac{t}{v}, \frac{z}{w}\} & \Rightarrow \Delta'' \{\frac{t}{v}, \frac{z}{w}\} \\
\Gamma' \Rightarrow \Delta', \Gamma'' \{\frac{t}{v}, \frac{z}{w}\}, \Gamma'' \{\frac{t}{v}, \frac{z}{w}\}, \Gamma'' \{\frac{t}{v}, \frac{z}{w}\} & \Rightarrow \Delta'' \{\frac{t}{v}, \frac{z}{w}\} \\
\Gamma' \Rightarrow \Delta', \Gamma'' \{\frac{t}{v}, \frac{z}{w}\}, \Gamma'' \{\frac{t}{v}, \frac{z}{w}\}, \Gamma'' \{\frac{t}{v}, \frac{z}{w}\} & \Rightarrow \Delta'' \{\frac{t}{v}, \frac{z}{w}\}
\end{align*}
\]

The limit of the sequence \(\{T_i\}_{i \geq 0}\) is a possibly infinite (and possibly open) derivation tree called the search tree for \(\Gamma \Rightarrow \Delta\) and denoted by \(T_\omega\).

Search trees are, by construction, recursive and cut-free. We construct special sequents out of search trees, called limit sequents, as follows.

**Definition 17 (Limit Sequents).** When a search tree \(T_\omega\) is not an RTC\(_G\) proof, either: (1) it is not even a pre-proof, i.e. it contains an open node; or (2) it is a pre-proof but contains an infinite branch that fails to satisfy the global trace condition. In case [1] it contains an open node to which, necessarily, no schedule element applies (e.g. a sequent containing only atomic formulas), for which we write \(\Gamma_\omega \Rightarrow \Delta_\omega\). In case [2] the global trace condition fails, so there
exists an infinite path \( \{ \Gamma_i \Rightarrow \Delta_i \}_{i>0} \) in \( T_\omega \) which is followed by no infinitely progressing traces; we call this path the untraceable branch of \( T_\omega \). We define \( \Gamma_\omega \) to be the set of formulas \((RTC_{x,y} \perp)(s,t)\) such that there is \( \varphi \) and \( j \) with \((RTC_{x,y} \varphi)(s,t) \in \Gamma_j \) the active formula in an instance of Rule (7) and \( \Gamma_{j+1} \) the left-hand premise of the rule instance. We then define \( \Gamma_\omega = \Gamma_\omega \cup \bigcup_{i>0} \Gamma_i \) and \( \Delta_\omega = \bigcup_{i>0} \Delta_i \), and call \( \Gamma_\omega \Rightarrow \Delta_\omega \) the limit sequent.\[5\]

Note that limit sequents may not technically be sequents since they can be infinite. However when we say that such a limit sequent is provable, we mean that it has a finite subsequent that is provable.

**Lemma 2.** Limit sequents \( \Gamma_\omega \Rightarrow \Delta_\omega \) are not cut-free provable.

**Proof.** Straightforward adaptation of the proof of [6, Lemma 6.3].

As standard, we use a limit sequent to induce a counter-interpretation.

**Definition 18 (Counter-interpretations).** Assume a search tree \( T_\omega \) which is not a \( RTC_\omega \) proof with limit sequent \( \Gamma_\omega \Rightarrow \Delta_\omega \). Let \( \sim \) be the smallest congruence relation on terms such that \((RTC_{x,y} \perp)(s,t) \in \Gamma_\omega \). Define a structure \( M_\omega \) as follows (where \( [t] \) stands for the \( \sim \)-equivalence class of \( t \)):

- The domain of \( M_\omega \) is the set of terms quotiented by the relation \( \sim \).
- For every \( k \)-ary function symbol \( f \) and relation symbol \( q \):
  \[ [f]_{M_\omega}([t_1],\ldots,[t_k]) = [f(t_1,\ldots,t_k)] \]
  \[ [q]_{M_\omega} = \{ ([t_1],\ldots,[t_k]) \mid q(t_1,\ldots,t_k) \in \Gamma_\omega \} \]

We also define a valuation \( \rho_\omega \) for \( M_\omega \) by \( \rho_\omega(x) = [x] \) for all variables \( x \).

The counter-interpretation \( \langle M_\omega, \rho_\omega \rangle \) has the following properties.

**Lemma 3.**
1. If \( M_\omega, \rho_\omega \models \psi \) then \( \Gamma_\omega \Rightarrow \psi \) is cut free provable.
2. If \( \psi \in \Gamma_\omega \) then \( M_\omega, \rho_\omega \models \psi \); and if \( \psi \in \Delta_\omega \) then \( M_\omega, \rho_\omega \not\models \psi \).

**Proof.** For \( \psi \) and the first part of \( \psi \), by well-founded induction using the lexicographic ordering of the number of binders (i.e. \( \exists \), \( \forall \), and \( RTC \)) in \( \psi \) and the structure of \( \psi \). Notice that, by definition, \( \rho_\omega(t) = [t] \) for all terms \( t \). In the following, we use \( \equiv \) to denote syntactic equality on terms, in order to avoid confusion with the notation \( s = t \), which we sometimes use as standing for the formula \((RTC_{x,y} \perp)(s,t)\).

\[5\] If \( \psi \) is atomic, by definition \( \psi \in \Gamma_\omega \), and so \( \Gamma_\omega \Rightarrow \psi \) is axiomatic. The cases in which \( \psi \) is a standard compound first-order formula follow straightforwardly by induction. In case \( \psi = (RTC_{x,y} \varphi)(s,t) \), the assumption \( M_\omega, \rho_\omega \models \psi \) entails either:

To be rigorous, we may pick e.g. the left-most open node or untraceable branch.
i) \( \rho_\omega(s) = \rho_\omega(t) \): if \( s \equiv t \) then we conclude with an application of rule (4) to derive \((RTC_{x,y} \varphi)(s,s)\), otherwise \( s \sim t \) which entails by Definition 18 that \((RTC_{x,y} \bot)(s,t) \in \Gamma_\omega \) and then we conclude first with an application of (4) to derive \((RTC_{x,y} \varphi)(s,s)\) and then an application of the derived rule \( \Rightarrow \) to derive \((RTC_{x,y} \bot)(s,t) \Rightarrow (RTC_{x,y} \varphi)(s,t)\).

ii) there exist \([u_0], \ldots, [u_n] \in D_\omega \) such that \( s \equiv u_0, t \equiv u_n \) and \( M, \rho_\omega[x := [u_i], y := [u_{i+1}]] \models \varphi \) for \( 0 \leq i < n \). Thus, for each \( 0 \leq i < n \), \( \rho_\omega \models \varphi \{ u_i, u_{i+1} \} \) and therefore by the I.H. \( \Gamma_\omega \Rightarrow \varphi \{ u_i, u_{i+1} \} \) is cut free provable. Then, using the fact that \( \Gamma_\omega \Rightarrow (RTC_{x,y} \varphi)(s,s) \) is cut-free provable using rule (4) we can construct a cut-free proof of \( \Gamma_\omega \Rightarrow (RTC_{x,y} \varphi)(s,t) \) by consecutive applications of Rule (6).

2. If \( \psi \) is atomic, by construction of \( T_\omega \) it must be of the form \( q(t_1, \ldots, t_k) \), and so by Definition 18 we have that \( M_\omega, \rho_\omega \models q(t_1, \ldots, t_k) \). The cases where \( \psi \) is a first-order formula follow straightforwardly by induction. In case \( \psi = (RTC_{x,y} \varphi)(s,t) \), assume \((RTC_{x,y} \varphi)(s,t) \in \Gamma_\omega \). Note that in this case \( \Gamma_\omega \Rightarrow \Delta_\omega \) is the limit sequent of the untraceable branch in \( T_\omega \). If \( \psi \in \Gamma_\omega \) then by definition, \( \varphi \equiv \bot \) and \( s \sim t \), and so \( \rho_\omega(s) = \rho_\omega(t) \); thus it follows immediately that \( M_\omega, \rho_\omega \models \psi \). If \( \psi \in \bigcup_{i > 0} \Gamma_i \), then by the construction of \( T_\omega \) there is at least one occurrence of rule (7) with active formula \( \psi \) in the untraceable branch; thus either: i) the branch follows the left-hand premise, so there is \((RTC_{x,y} \bot)(s,t) \in \Gamma_\omega \) and \( M_\omega, \rho_\omega \models \psi \) follows as above; or ii) the branch follows the right-hand premise and, since there is no infinitely progressing trace along the untraceable branch, there must be a finite number of distinct variables \( z_1, \ldots, z_n \) such that \( \varphi \{ z_i, z_{i+1} \} \in \Gamma_\omega \), for each \( 0 < i < n \), and \( \varphi \{ z_i, t \} \in \Gamma_\omega \). Then, by the I.H., \( M_\omega, \rho_\omega \models \varphi \{ z_i, z_{i+1} \} \) for each \( 0 < i < n \), and \( M_\omega, \rho_\omega \models \varphi \{ z_i, t \} \). Thus, \( M_\omega, \rho_\omega[x := [z_i], y := [z_{i+1}] \models \varphi \) for each \( 0 < i < n \), and \( M_\omega, \rho_\omega[x := [z_n], y := [t] \models \varphi \). Moreover, the untraceable branch also follows the left-hand premise of rule (7) with active formula \((RTC_{x,y} \varphi)(s,z_1)\); thus it is the case that \((RTC_{x,y} \bot)(s,z_1) \in \Gamma_\omega \) and so \( \rho_\omega(s) = \rho_\omega(z_1) = [z_1] \). \( M_\omega, \rho_\omega \models \psi \) then follows from Definition 2.

For the second part of 2, let \( \psi \in \Delta_\omega \), and assume for contradiction that \( M_\omega, \rho_\omega \models \psi \). Then, by Item 1 \( \Gamma_\omega \Rightarrow \psi \) is cut-free provable. But this entails that \( \Gamma_\omega \Rightarrow \Delta_\omega \) is cut-free provable, which contradicts Lemma 2.

This means that \( M_\omega \) is a counter-model for a sequent whenever the search tree \( T_\omega \) is not a proof, whence completeness follows.

Theorem 3 (Completeness). \( RTC_\omega^c \) is complete for standard semantics.

Proof. Now given any sequent \( S \), if some search tree \( T_\omega \) contracted for \( S \) is not an \( RTC_\omega^c \) proof then it follows from Lemma 3 that \( S \) is not valid (\( M_\omega \) is a counter model for it). Thus if \( S \) is valid, then \( T_\omega \) is a recursive \( RTC_\omega^c \) proof for it.

We obtain admissibility of cut as the search tree \( T_\omega \) is cut-free.

Corollary 2 (Cut admissibility). Cut is admissible in \( RTC_\omega^c \).
3.4 Proof Systems for $L_{RTC}$ with Pairs

The above proof systems are extended to capture ordered pairs as follows.

**Definition 19.** Let $\sigma$ be a signature that contains some constant $c$, and a binary function symbol, denoted by $\langle \rangle$. Let $L_{RTC}$ be a language based on $\sigma$. The proof systems $\langle RTC \rangle_G$, $\langle RTC \rangle_\omega G$, and $\langle CRTC \rangle_\omega G$ are obtained from $RTC_G$, $RTC_\omega G$, $CRTC_\omega G$ (respectively) by the addition of the following rules:

- $\Gamma \Rightarrow \langle x, y \rangle = \langle u, v \rangle, \Delta$
- $\Gamma \Rightarrow x = u \land y = v, \Delta$
- $\Gamma, \langle x, y \rangle = c \Rightarrow \Delta$

The proofs of Theorems 1 and 3 can easily be extended to obtain the following results, for languages with a pairing function. For completeness, the key observation is that the model of the counter-interpretation is one in which every binary function is a pairing function. Thus, it is an admissible structure.

**Theorem 4 (Soundness and Completeness of $\langle RTC \rangle_G$ and $\langle RTC \rangle_\omega G$).** The proof systems $\langle RTC \rangle_G$ and $\langle RTC \rangle_\omega G$ are both sound and complete for the admissible forms of Henkin and standard semantics, respectively.

4 Relating the Finitary and Infinitary Proof Systems

This section discusses the relation between the explicit and the cyclic system for TC. In Section 4.1 we show that the former is contained in the latter. The converse direction, which is much more subtle, is discussed in Section 4.2.

4.1 Inclusion of $RTC_G$ in $CRTC_\omega G$

We can show that the explicit induction rule is derivable in the cyclic system. The key property is that we can derive the explicit induction rule in the cyclic system, as shown below in Figure 1.

**Lemma 4.** Rule [6] is derivable in $CRTC_\omega G$.

This leads to the following result (an analogue to [6, Thm. 7.6]).

**Theorem 5.** $CRTC_\omega G \supseteq RTC_G$, and is thus complete w.r.t. Henkin semantics.

**Proof.** Let $\mathcal{P}$ be a proof in $RTC_G$ and $\mathcal{P}'$ be the corresponding pre-proof in $CRTC_\omega G$ obtained by replacing each instance of Rule [6] by the corresponding instance of the proof schema given in Lemma 1. We argue that $\mathcal{P}'$ is a valid $CRTC_\omega G$ proof. Notice that the only cycles in $\mathcal{P}'$ are internal to the subproofs that simulate Rule [6]. Thus any infinite path in $\mathcal{P}'$ must eventually end up traversing one of these cycles infinitely often. Therefore, it suffices to show that there is an infinitely progressing trace following each such path. This is clearly the case since we can trace the active $RTC$ formulas along these paths, which progress once each time around the cycle, across Rule [7].

□
Lemma 4 is the TC counterpart of [6, Lemma 7.5]. It is interesting to note that the simulation of the explicit LKID induction rule in the cyclic LKID system is rather complex since each predicate has a slightly different explicit induction rule, which depends on the particular productions defining it. Thus, the construction for the cyclic LKID system must take into account the possible forms of arbitrary productions. In contrast, CRTC\textgreek{w} provides a single, uniform way to unfold an RTC formula: the construction given in Fig. 1 is the cyclic representation of the RTC operator semantics, with the variables \(v\) and \(w\) implicitly standing for arbitrary terms (that we subsequently substitute for).

This uniform syntactic translation of the explicit RTC\textgreek{w} induction rule into CRTC\textgreek{w} allows us to syntactically identify a proper subset of cyclic proofs which is also complete w.r.t. Henkin semantics\[6\]. The criterion we use is based on the notion of overlapping cycles. Recall the definition of a basic cycle, which is a path in a (proof) graph starting and ending at the same point, but containing no other repeated nodes. We say that two distinct (i.e. not identical up to permutation) basic cycles overlap if they share any nodes in common, i.e. at some point they both traverse the same path in the graph. We say that a cyclic proof is non-overlapping whenever no two distinct basic cycles it contains contains overlap.

**Definition 20 (Normal Cyclic Proofs).** The normal cyclic proof system NCRTC\textgreek{w} is the subsystem of RTC\textgreek{w} comprising of all and only the non-overlapping cyclic proofs.

Henkin-completeness of the normal cyclic system follows from the fact that the translation of an RTC\textgreek{w} proof into CRTC\textgreek{w}, using the construction shown in Figure 1 results in a proof with no overlapping cycles.

**Theorem 6.** NCRTC\textgreek{w} \(\subseteq\) RTC\textgreek{w}, and is thus complete w.r.t. Henkin semantics.

\[6\] Note it is not clear that a similar complete structural restriction is possible for LKID.
4.2 Inclusions of CRT

C RTCω in RTC

This section addresses the question of whether the cyclic system is equivalent to the explicit one, or strictly stronger. In [6] it was conjectured that for the system with inductive definitions, LKID and CLKIDω are equivalent. Later, it was shown that they are indeed equivalent when containing arithmetics [3,18]. We obtain a corresponding theorem in Section [4.2.1] for the TC systems. However, it was also shown in [2] that in the general case the cyclic system is stronger than the explicit one. We discuss the general case for TC and its subtleties in Section [4.2.2].

4.2.1 The Case of Arithmetics

Let \( L_{RTC} \) be a language based on the signature \( \{0, s, +\} \). Let \( RTC_G + A \) and \( CRTC_\omega G + A \) be the systems for \( L_{RTC} \) obtained by adding to \( RTC_G \) and \( CRTC_\omega G \), respectively, the standard axioms of PA together with the \( RTC \)-characterization of the natural numbers, i.e.:

1. \( s \cdot x = 0 \Rightarrow \)
2. \( s \cdot x = s \cdot y \Rightarrow x = y \)
3. \( x + 0 = x \)
4. \( x + s \cdot y = s \cdot (x + y) \)
5. \( (RTC_{w,u}s \cdot w = u)(0, x) \)

Note that we do not need to assume multiplication explicitly in the signature, nor do we need to add axioms for it, since multiplication is definable in \( L_{RTC} \) and its standard axioms are derivable (vide [1,9]).

Recall that we can express facts about sequences of numbers in PA by using a \( \beta \)-function such that for any finite sequence \( k_0, k_1, ..., k_n \) there is some \( c \) such that for all \( i \leq n \), \( \beta(c, i) = k_i \). Accordingly, let \( B \) be a wff of the language of PA with three free variables which captures in PA a \( \beta \)-function. For each formula \( \varphi \) of the language of PA define \( \varphi^\beta := \varphi \), and define \((RTC_{x,y}\varphi)(s,t))^\beta\) to be:

\[
s = t \lor \exists z, c \cdot B(c, 0, s) \land B(c, s, t) \land \forall u \leq z \cdot \exists v, w \cdot B(c, u, v) \land B(c, s, u, w) \land \varphi^\beta \left\{ \frac{v}{x}, \frac{w}{y} \right\}
\]

The following result, which was proven in [7,9], establishes an equivalence between \( RTC_G + A \) and \( PA_G \) (Gentzen-style system for PA). It is mainly based on the fact that in \( RTC_G + A \) all instances of \( PA_G \) induction rule are derivable.

**Theorem 7 (cf. [9]).** 1. \( \vdash_{RTC_G + A} \varphi \iff \varphi^\beta \).
2. \( \vdash_{RTC_G + A} \Gamma \Rightarrow \Delta \text{ iff } \vdash_{PA_G} \Gamma^\beta \Rightarrow \Delta^\beta. \)

We show a similar equivalence holds between the cyclic system \( CRT_\omega G \) and \( CA_G \), a cyclic system for arithmetic shown to be equivalent to \( PA_G \) [18]. The first part of the result is straightforward.

**Lemma 5.** \( \vdash_{CRT_\omega G + A} \varphi \iff \varphi^\beta \)

**Proof.** This follows from Theorem [7,11] and Theorem [5].
To show the second part, we first show that $\text{CA}_G$ is included in $\text{CRTC}_G \cdot +A$ by giving a construction that directly translates $\text{CA}_G$ proofs into $\text{CRTC}_G \cdot +A$ proofs. Technically, the signature of $\text{CA}_G$ includes the relation symbol $<$ for strict ordering, and the function symbol $\cdot$ for multiplication. As mentioned above, multiplication (and its axioms) are derivable in TC, and the non-strict ordering on natural numbers $s \leq t$ can be expressed in TC as $(\text{RTC}_{w,u} s \cdot w = u)(s,t)$. Therefore, in the following result, we implicitly assume that all $\text{CA}_G$ terms of the form $s \cdot t$ are translated in $\text{CRTC}_G$ as in [19], and formulas of the form $s < t$ as $s \neq t \land (\text{RTC}_{w,u} s \cdot w = u)(s,t)$.

**Lemma 6.** If $\vdash_{\text{CA}_G} \Gamma \Rightarrow \Delta$ then $\vdash_{\text{CRTC}_G \cdot +A} \Gamma \Rightarrow \Delta$.

**Proof.** We define a translation $[\cdot]^*$ on sets of formulas $\Gamma$ as the smallest set of formulas containing $\Gamma$ and $(\text{RTC}_{w,u} s \cdot w = u)(0,t)$ for each term $t$ that appears as a subterm of a formula in $\Gamma$ with no quantified variables (we call such a term free). Notice it suffices to prove $\vdash_{\text{CRTC}_G \cdot +A} \Gamma^* \Rightarrow \Delta$ if $\vdash_{\text{CA}_G} \Gamma \Rightarrow \Delta$, since then we may cut all the added formulas $(\text{RTC}_{w,u} s \cdot w = u)(0,t)$ in $\Gamma^*$ using instances of axiom [3].

We call $\Gamma^* \Rightarrow \Delta$ the *-translation of the sequent $\Gamma \Rightarrow \Delta$, and call the inference rule obtained by applying this translation to the conclusion and each premise the *-translation of the rule. We begin by showing that the *-translations of the axioms and proof rules of $\text{CA}_G$ are derivable in $\text{CRTC}_G \cdot +A$. For the standard rules of $\mathcal{LK}$ and the substitution rule this is trivial, and can be done in such a way that, for the conclusion $\Gamma \Rightarrow \Delta$ and each premise $\Gamma' \Rightarrow \Delta'$, there is a (non-progressing) trace from each $(\text{RTC}_{w,u} s \cdot w = u)(0,t) \in \Gamma^*$ to $(\text{RTC}_{w,u} s \cdot w = u)(0,t') \in (\Gamma')^*$ when $t'$ is a predecessor of $t$ (cf. [18] Def. 1). It remains to show that we can derive the *-translations of the following set of axioms:

\[
\begin{align*}
&t < u, u < v \Rightarrow t < v & \Rightarrow t < u, t = u, u < t \\
&t < u, u < t \Rightarrow \Rightarrow t + 0 = t \\
&t < u, u < st \Rightarrow \Rightarrow t + s \cdot u = s(t + u) \\
&\quad t < 0 \Rightarrow \quad \Rightarrow t \cdot 0 = 0 \\
&\quad t < u \Rightarrow st < s \cdot u \Rightarrow t < s \cdot t \\
&\quad \Rightarrow t < s \cdot t
\end{align*}
\]

and also the following inference rule: $\Gamma, t = sx \Rightarrow \Delta \quad \frac{\Gamma, 0 < t \Rightarrow \Delta}{\Gamma, 0 < t \Rightarrow \Delta} \quad (x \text{ is fresh}).$ We shall call these derivations of the *-translations of the $\text{CA}_G$ rules the simple derivations.

We next show that for each (non-axiomatic) rule of $\text{CA}_G$, we can derive the *-translation of the rule in $\text{CRTC}_G \cdot +A$ with the additional property that for every progressing trace from a term $t$ in the conclusion to a term $t'$ in a premise, there is also a progressing trace from the formula $(\text{RTC}_{w,u} u = s \cdot w)(0,t)$ in the conclusion of the derived rule to the formula $(\text{RTC}_{w,u} u = s \cdot w)(0,t')$ in the corresponding premise.

Consider the open derivation schema shown in Figure [2]. Here, $Nt$ abbreviates the formula $(\text{RTC}_{w,u} u = s \cdot w)(0,t)$, and $t' < t$ abbreviates the translation given
above, i.e. \( s \neq t \land (RTC_{w,u}sw = u)(s, t) \). The symbols \( \dagger \), \( \ddagger \) and \( \star \) denote the (simple) derivations of the \( * \)-translations of the appropriate axioms. Notice that, crucially, in this derivation there is a progressing trace from \( Nt \) in the conclusion to \( Nt' \) in the premise (marked with a \( * \)). There is also a non-progressing trace from each \( RTC \) formula in \( \Gamma \) in the conclusion to its occurrence in \( \Gamma \) in the premise.

For each \( CA_G \) inference rule concluding \( \Gamma \Rightarrow \Delta \), we build a trace-aware derivation of its \( * \)-translation as follows. First, we take the simple derivation of the inference rule. Notice that this gives a non-progressing trace from each \( Nt \) in the conclusion to \( Nt' \) in a premise where \( t' \) is a precursor of \( t \) in the \( CA_G \) inference rule. Now, for every premise \( \Gamma \Rightarrow \Delta \) of the rule, let \( \{ t'_1 < t_1, \ldots, t'_n < t_n \} \) be the set of all such formulas in \( \Gamma \). Notice that \( t'_i < t_i, Nt'_i, Nt_i \subseteq (\Gamma_i)^* \) for each \( i \). To each corresponding premise in the simple derivation we apply, in turn, \( n \) instances of the derivation schema in Figure 2 [2] one for each formula \( t_i < t'_i \). Notice that this combined derivation satisfies the following for each premise \( \Gamma_j \Rightarrow \Delta_j \) and free terms \( t \in \Gamma_j \) and \( t_i \in \Gamma_j \):

1. there is a non-progressing trace from \( Nt \in \Gamma^* \) in the conclusion to \( Nt_i \in (\Gamma_i)^* \) in the premise if \( t_i \) is a precursor of \( t \); and
2. there is a progressing trace from \( N t \in \Gamma^* \) in the conclusion to \( N t_i' \in (\Gamma_i)^* \) in the premise if \( t_i \) is a precursor of \( t \) and \( t_i' < t_i \in \Gamma_i \).

Note that this means the notion of trace in the \( CA_G \) inference rules is exactly mirrored by a \( CRTC_G^\omega \) trace in the trace-aware derivation. Moreover, the properties above hold of all paths from the conclusion to a premise in the trace-aware derivation, in particular those that travel around the internal cycles any finite number of times.

From this construction it follows that we can transform a \( CA_G \) pre-proof, via the trace-aware local \(*\)-translation of each rule, into a \( CRTC_G^\omega + A \) pre-proof with the same global structure. It remains to show that each such \( CRTC_G^\omega + A \) pre-proof resulting from a \( CA_G \) proof is also a \( CRTC_G^\omega + A \) proof. That is, it satisfies the \( CRTC_G^\omega \) global trace condition. Consider an arbitrary infinite path in the \( CRTC_G^\omega + A \) pre-proof. There are two cases to consider:

- The infinite path ends up traversing an infinite path local to the \( (\text{trace-aware}) \)
  \(*\)-translation of an inference rule or \( CA_G \) axiom; in this case notice that each such infinite path has an infinitely progressing trace.
- The infinite path corresponds to an infinite path in the \( CA_G \) proof (possibly interspersed with finite traversals of the cycles local to the trace-aware \(*\)-translation of each rule instance). Since there is an infinitely progressing trace following the path in the \( CA_G \) proof, by the properties above there is also a corresponding infinitely progressing trace following the path in the \( CRTC_G^\omega + A \) pre-proof.

\( \Box \)

This leads immediately to the ‘if’ direction.

**Corollary 3.** If \( \vdash_{CA_G} \Gamma^\beta \Rightarrow \Delta^\beta \) then \( \vdash_{CRTC_G^\omega + A} \Gamma \Rightarrow \Delta \).

**Proof.** We first use Lemma 6 to derive \( \Gamma^\beta \Rightarrow \Delta^\beta \) in \( CRTC_G^\omega + A \), and then combine this with derivations in \( CRTC_G^\omega + A \) of \( \varphi \Rightarrow \varphi^\beta \) (resp. \( \varphi^\beta \Rightarrow \varphi \)) for each \( \varphi \in \Gamma \) (resp. \( \varphi \in \Delta \)), which exist by Lemma 5 with applications of cuts to derive \( \Gamma \Rightarrow \Delta \).

\( \Box \)

In [3], to show the equivalence of the explicit and cyclic systems for \( LKID \), a construction was given which translates \( LKID \) predicates \( P(t) \) into predicates \( P'(t, n) \) with equivalent inductive definitions and an extra parameter \( n \) comprising a ‘stage’ variable. The equivalence is derived by using the cycles in a proof to construct an explicit induction hypothesis over these stage variables. Here, for \( TC \), instead of directly constructing an induction hypothesis for the explicit system, we show that from a \( CRTC_G^\omega + A \) proof we can construct an analogous proof in \( CA_G \) which preserves cycles, and then use the existing equivalence results between \( CA_G, PA_G \) and \( RTC_G \). Our construction is similar to the one given in [3], in that we use a variant of the \( \beta \)-function which introduces a free variable \( n \) (similar to a stage variable), which we are able to trace in the cyclic \( CA_G \) proof. This results in the ‘only if’ direction of the result.

**Lemma 7.** If \( \vdash_{CRTC_G^\omega + A} \Gamma \Rightarrow \Delta \) then \( \vdash_{CA_G} \Gamma^\beta \Rightarrow \Delta^\beta \).
Proof. Similarly to the proof of Lemma ?? above, we define a local translation on proof rules that preserves $\mathsf{CRTC}_{G}^{\omega} + A$ traces as $\mathsf{CA}_{G}$ traces. For this, we use a parameterised variant $\bar{\beta}[n]$ of the $\beta$-translation, which introduces its parameter as a free variable in the translation. It is defined in the same way as the $\beta$-translation, except that it translates formulas of the form $(\mathsf{RTC}_{x,y} \varphi)(s, t)$ as follows:

\[ s = t \lor \exists z, c \cdot n = s z \land B(c, 0, s) \land B(c, s z, t) \land \\
\forall u \leq z \cdot \exists v, w \cdot B(c, u, v) \land B(c, s u, w) \land \varphi^\beta \left\{ \frac{v}{x}, \frac{w}{y} \right\} \]

Notice the use of the original $\beta$-translation for the body of the $\mathsf{RTC}$ formula $\varphi$. We extend the $\bar{\beta}$-translation to sets of formulas, sequents, and inference rules as follows:

- for a set of formulas $\Gamma$, we define $\Gamma^{\bar{\beta}}$ as the set of $\bar{\beta}$-translations of the formulas in $\Gamma$ such that each translation of an $\mathsf{RTC}$ sub-formula introduces a fresh free variable $z$; that is, each distinct $\mathsf{RTC}$ sub-formula is translated using a distinct variable parameter;

- for sequents, we define $(\Gamma \Rightarrow \Delta)^{\bar{\beta}} = \Gamma^{\bar{\beta}} \Rightarrow \Delta^{\bar{\beta}}$ such that the free variable parameters used by the $\bar{\beta}$-translation of the antecedent $\Gamma$ are distinct from the free variables in the succedent $\Delta$;

- for an inference rule with premises $\Gamma_1 \Rightarrow \Delta_1, \ldots, \Gamma_n \Rightarrow \Delta_n$ and conclusion $\Gamma \Rightarrow \Delta$, we define its $\bar{\beta}$-translation as the inference rule with premises $(\Gamma_1 \Rightarrow \Delta_1)^{\bar{\beta}}, \ldots, (\Gamma_n \Rightarrow \Delta_n)^{\bar{\beta}}$ and conclusion $(\Gamma \Rightarrow \Delta)^{\bar{\beta}}$ such that multiple occurrences of the same $\mathsf{RTC}$ sub-formula across the original premises and conclusion are translated using the same free variable parameter in each of the translated premises and conclusion.

We show that if $\vdash \mathsf{CRTC}_{G}^{\omega} + A \Gamma \Rightarrow \Delta$ then $\vdash \mathsf{CA}_{G} (\Gamma \Rightarrow \Delta)^{\bar{\beta}}$. We first prove that the $\bar{\beta}$-translation of each $\mathsf{CRTC}_{G}^{\omega} + A$ inference rule can be derived in $\mathsf{CA}_{G}$ in such a way that there is a $\mathsf{CA}_{G}$ trace (progressing or non-progressing, as appropriate) simulating each $\mathsf{CRTC}_{G}^{\omega} + A$ trace present in the original rule, which we do by tracing the free variable parameters in the $\mathsf{CA}_{G}$ rule. We show how this is done for Rule (7); the other rules are straightforward, and do not contain progressing traces.

Take an instance of Rule (7) with contexts $\Gamma$ and $\Delta$, and active formulas $(\mathsf{RTC}_{x,y} \varphi)(s, t)$ and $(\mathsf{RTC}_{x,y} \varphi)(s, z)$ in the conclusion and right-hand premise, respectively. For terms $r$, $s$, and $t$, let:

- $\vartheta(r)$ abbreviate $\exists v, w \cdot B(c, r, v) \land B(c, s r, w) \land \varphi^\beta \left\{ \frac{v}{x}, \frac{w}{y} \right\}$; and

- $A(r, s, t)$ abbreviate $\exists z, c \cdot r = s z \land B(c, 0, s) \land B(c, s z, t) \land \forall u \leq z \cdot \vartheta(u)$.

Additionally, let $\Sigma(r)$ and $\Pi(r)$ abbreviate the follow sequences of formulas:

\[ n = s r, B(c, 0, s), B(c, s r, t), \forall u \leq r \cdot \vartheta(u) \]

\[ n = s s r, B(c, 0, s), B(c, s r, t), B(c, s r, v), B(c, s s r, w), \varphi^\beta \left\{ \frac{v}{x}, \frac{w}{y} \right\}, \forall u \leq r \cdot \vartheta(u) \]

Moreover, note the following.
\[
\varphi^\beta \{ z : \bar{z} \} \Rightarrow \varphi^\beta \{ \bar{z} : z \} \quad (Ax)
\]
\[
s = u, t = w, \varphi^\beta \{ z : \bar{z} \} \Rightarrow \varphi^\beta \{ \bar{z} : z \} \quad (=L)
\]
\[
B(c, 0, s), B(c, s, 0), t, B(c, 0, u), \varphi^\beta \{ z : \bar{z} \} \Rightarrow \varphi^\beta \{ \bar{z} : z \} \quad (B)
\]
\[
B(c, 0, s), B(c, s, 0), t, \forall u \leq 0, \theta(u) \Rightarrow \varphi^\beta \{ z : \bar{z} \} \quad (\leq_0 / \exists L)
\]
\[
\Rightarrow z = 0, \Sigma(z) \Rightarrow \varphi^\beta \{ z : \bar{z} \} \quad (WL / = L)
\]
\[
\text{(a) One step from } s \text{ to } t.
\]
\[
\begin{align*}
H(z') & \Rightarrow s' = z' \quad (=R) \\
H(z') & \Rightarrow B(c, 0, s) \quad (Ax) \\
H(z') & \Rightarrow B(c, s, z') \quad (Ax) \\
H(z') & \Rightarrow \forall u \leq z', \theta(u) \quad (Ax)
\end{align*}
\]
\[
H(z') \Rightarrow s = v \vee A(z', s, v) \\
H(z') \Rightarrow \varphi^\beta \{ \bar{z} : z \} \quad (=L)
\]
\[
H(z') \Rightarrow s' < s \quad (=L) \\
H(z') \Rightarrow s' < n \quad (Ax)
\]
\[
H(z') \Rightarrow B(z, m, m < n \land (s = z \lor A(m, s, z)) \land \varphi^\beta \{ \bar{z} : z \}) \quad (\forall R / = L)
\]
\[
H(z') \Rightarrow B(z, m, m < n \land (s = z \lor A(m, s, z)) \land \varphi^\beta \{ \bar{z} : z \}) \quad (=L)
\]
\[
H(z') \Rightarrow B(z, m, m < n \land (s = z \lor A(m, s, z)) \land \varphi^\beta \{ \bar{z} : z \}) \quad (B / WL)
\]
\[
H(z') \Rightarrow s' < n \\
H(z') \Rightarrow \forall u \leq z', \theta(u) \Rightarrow \psi \\
B(z', z = s, \Sigma(z) \Rightarrow \psi)
\]
\[
\text{(b) Multi-step from } s \text{ to } t.
\]
\[
\begin{align*}
\vdash & z = 0, \Sigma(z) \Rightarrow \psi \\
\vdash & z = s, \Sigma(z) \Rightarrow \psi \\
\vdash & z = 0, \Sigma(z) \Rightarrow \psi \\
\vdash & z = 0, \Sigma(z) \Rightarrow \psi \\
\vdash & z = 0, \Sigma(z) \Rightarrow \psi \\
\vdash & z = 0, \Sigma(z) \Rightarrow \psi
\end{align*}
\]
\[
\Rightarrow z = 0 \lor \exists z'. z = s^0, \Sigma(z) \Rightarrow \psi \quad (\lor L)
\]
\[
\Sigma(z) \Rightarrow \exists z, m < n \land (s = z \lor A(m, s, z)) \land \varphi^\beta \{ \bar{z} : z \} \quad (\exists L)
\]
\[
A(n, s, t) \Rightarrow \exists z, m < n \land (s = z \lor A(m, s, z)) \land \varphi^\beta \{ \bar{z} : z \} \quad (\exists L)
\]
\[
\text{Fig. 3: The core subderivation of the simulation of Rule } [7] \text{ in } \mathsf{CA}_G.
\]

i) We can easily derive \( \Rightarrow z = 0 \lor \exists z'. z = s \) using standard first-order rules and the axioms of \( \mathsf{CA}_G \); we refer to this derivation using \( \vdash \).

ii) Our use of the notation \( \forall u \leq t \) \( \gamma \) technically abbreviates the \( \mathsf{CA}_G \) formula \( \forall u. (u = t \lor u < t) \rightarrow \gamma \), and so we may straightforwardly derive both \( \forall u \leq 0 \cdot \gamma(u) \Rightarrow \gamma(0) \) and \( \forall u \leq s \cdot \gamma(u) \Rightarrow \gamma(s \downarrow) \land \forall u \leq t \cdot \gamma(u) \); for brevity, we refer to an instance of the (Cut) rule that applies these sequents using the labels \( (\leq_0) \) and \( (\leq_s) \), respectively.

20
A(n, s, t) ⇒ ∃z, m . m < n ∧ (s = z ∨ A(m, s, z)) ∧ ϕβ {z, t}

\[ \Gamma^\beta, (s = z ∨ A(m, s, z)), ϕ^\beta \{z, t\} \Rightarrow Δ^\beta \]
\[ \Gamma^\beta, (s = z ∨ A(m, s, z)), ϕ^\beta \{z, t\} \Rightarrow Δ^\beta \]
\[ \Gamma^\beta, m < n \] (s = z ∨ A(m, s, z)), ϕ^β {z, t} \Rightarrow Δ^\beta \]
\[ \Gamma^\beta, m < n \] (s = z ∨ A(m, s, z)), ϕ^β {z, t} \Rightarrow Δ^\beta \]
\[ \Gamma^\beta, s = t \Rightarrow Δ^\beta \]
\[ \Gamma^\beta, s = t ∨ A(n, s, t) \Rightarrow Δ^\beta \]

Fig. 4: A derivation schema simulating Rule (7) in CA_G.

iii) Recall that, since the formula B captures a β-function, we may also derive B(r, s, t), B(r, s, u) ⇒ t = u; we abbreviate instances of (Cut) that apply an instance of this sequent using the label (B).

Using these elements, Fig. 3 shows a derivation of the following sequent, in which we have abbreviated the antecedent formula by ψ:

A(n, s, t) ⇒ ∃z, m . m < n ∧ (s = z ∨ A(m, s, z)) ∧ ϕβ {z, t}

Then, using Fig. 3 as a subderivation, we derive the β-translation of Rule (7) in CA_G as shown in Fig. 4. Note that for any sequence of formulas Σ, we can straightforwardly derive Γ, Σ ⇒ Δ from Γ, Σ ⇒ Δ by first introducing existential quantifiers for the free variable parameters in Σ and then eliminating the terms n < s z with cuts. We abbreviate such a derivation using the label (β). Note that this admits non-progressing traces for all the free variable parameters in Γ. The crucial feature of this derivation is that there is a CA_G trace from the free variable parameter n in the conclusion to m in the right-hand premise, which progresses at the sequent containing the boxed formula m < n. Also, since the context Γ is preserved along the paths to both the left and right premises, all non-progressing traces are simulated as well.

Now, using the derivations of the local β-translations of the inference rules, from a CRTC_G+A pre-proof, we can build a CA_G pre-proof with the same global structure. For each bud in the resulting CA_G pre-proof, we first apply an instance of the substitution rule that substitutes each free variable parameter of its companion with the free variable parameter of its corresponding β-translation instance. Notice that this is possible, since the parameter variable is unique for the β-translation of each RTC sub-formula. We can then form a cycle in the CA_G pre-proof.

Since the CRTC_G traces for each rule are simulated by the CA_G derived rules, for each trace following a (finite or infinite) path in the CRTC_G+A pre-proof
there is a trace following the corresponding path in the \( CA_G \) pre-proof containing a progression point for each progression point in the \( CRTC^{\omega}_G + A \) trace. From this it follows that if the \( CRTC^{\omega}_G + A \) pre-proof satisfies the \( (RTC^{\omega}_G) \) global trace condition, then its translation satisfies the \( CA_G \) global trace condition. Finally, we derive \( \Gamma^\beta \Rightarrow \Delta^\beta \) from \( \Gamma^{\bar{\beta}} \Rightarrow \Delta^\beta \) as described above.

\[ \Box \Box \]

**Theorem 8.** \( \vdash_{CRTCA^{\omega}_G + A} \Gamma \Rightarrow \Delta \iff \vdash_{CA_G} \Gamma^{\bar{\beta}} \Rightarrow \Delta^\beta \).

**Proof.** By Corollary 3 and Lemma 7.

These results allow us to show an equivalence between the finitary and cyclic systems for TC with arithmetic.

**Theorem 9.** \( RTC_G^A \) and \( CRTC^{\omega}_G + A \) are equivalent.

**Proof.** The fact that \( RTC_G^A \subseteq CRTC^{\omega}_G + A \) follows immediately from Theorem 5. For the converse, suppose \( \Gamma \Rightarrow \Delta \) is provable in \( CRTC^{\omega}_G + A \). By Theorem 8 we get that \( \vdash_{CA_G} \Gamma^{\bar{\beta}} \Rightarrow \Delta^\beta \). Using the equivalence between \( CA_G \) and \( PA_G \), we obtain \( \vdash_{PA_G} \Gamma^{\bar{\beta}} \Rightarrow \Delta^\beta \). Then we conclude using Theorem 7(2).

\[ \Box \Box \]

Note that the result above can easily be extended to show that adding the same set of additional axioms to both \( RTC_G^A \) and \( CRTC^{\omega}_G + A \) results in equivalent systems. Also note that in the systems with pairs, to embed arithmetics there is no need to explicitly include addition and its axioms. Thus, by only including the signature \( \{0, s\} \) and the corresponding axioms for it we can obtain that \( \langle RTC \rangle_G^A \) and \( \langle CRTC \rangle^{\omega}_G + A \) are equivalent.

In [3], the equivalence result of [18] was improved to show it holds for any set of inductive predicates containing the natural number predicate \( N \). On the one hand, our result goes beyond that of [3] as it shows the equivalence for systems with a richer notion of inductive definition, due to the expressiveness of TC. On the other hand, TC does not support restricting the set of inductive predicates, i.e. the \( RTC \) operator may operate on any formula in the language. To obtain a finer result which corresponds to that of [3] we need to further explore the transformations between proofs in the two systems. This is left for future work.

### 4.2.2 The General Case

As mentioned, the general equivalence conjecture between LKID and CLKID\(^\omega\) was refuted in [2], by providing a concrete example of a statement, namely the 2-Hydra statement, which is provable in the cyclic system but not in the explicit one. However, a careful examination of this counter-example reveals that it only refutes a strong form of the conjecture, according to which both systems are based on the same set of productions. In fact, already in [2] it is shown that if the explicit system is extended by another inductive predicate, namely one expressing the \( \leq \) relation, then the 2-Hydra statement becomes provable. Therefore, the less strict formulation of the question, namely whether for any proof in CLKID\(^\omega\) there is a proof in LKID\(\phi'\) for some \( \phi' \supseteq \phi \), has not yet been resolved. Notice
that in TC the equivalence question is of this weaker variety, since the RTC operator 'generates' all inductive definitions at once. That is, there is no a priori restriction on the inductive predicates one is allowed to use. Indeed, the 2-Hydra statement can be expressed in \( \mathcal{L}_{RTC} \) and proved in \( \mathcal{CRTC}^G_\omega \). However, this does not produce a counter-example for TC since it is also provable in \( \mathcal{RTC}^G_\omega \), due to the fact that \( s \leq t \) is definable via the \( \mathcal{RTC} \) formula \( \psi = \langle RTC \_w,u \_s \_w = u \rangle \langle s,t \rangle \).

Despite our best efforts, we have not yet managed to settle this question, which does not seem to be straightforward. What is more, we do not have a strong conviction regarding the answer. One possible approach to solving it is the semantical one, i.e. exploiting the fact that the explicit system is known to be sound w.r.t. Henkin semantics. This is what was done in \[2\]. Thus, to show strict inclusion one can construct a different statement which is provable in \( \mathcal{CRTC}^G_\omega \), and use a similar methodology to provide a Henkin model for TC that is not a model of the statement. However, constructing a TC Henkin model appears to be non-trivial, due to its rich inductive power. Alternatively, to prove equivalence, one could show that \( \mathcal{CRTC}^G_\omega \) is also sound w.r.t. Henkin semantics. Here, again, proving this does not seem to be straightforward.

In our setting, there is also the question of the inclusion of \( \mathcal{CRTC}^G_\omega \) in \( \mathcal{NCRTC}^G_\omega \), which amounts to the question of whether overlapping cycles can be eliminated. Moreover, we can ask if \( \mathcal{NCRTC}^G_\omega \) is included in \( \mathcal{RTC}^G_\omega \), independently of whether this also holds for \( \mathcal{CRTC}^G_\omega \). Again, the semantic approach described above may prove fruitful in answering these questions.

5 Conclusions and Future Work

We developed a natural infinitary proof system for transitive closure logic which is cut-free complete for the standard semantics and subsumes the explicit system. We further explored its restriction to cyclic proofs which provides the basis for an effective system for automating inductive reasoning. In particular, we syntactically identified a subset of cyclic proofs that is Henkin-complete. A summary of the proof systems we have studied in this paper, and their interrelationships, is shown in Figure 5. Where an edge between systems is labelled with an inclusion \( \subseteq \), this signifies that a proof in the source system is already a proof in the destination system.

As mentioned in the introduction, as well as throughout the paper, this research was motivated by other work on systems of inductive definitions, particularly the LKID framework of \[6\], its infinitary counterpart LKID\(^\omega \), and its cyclic subsystem CLKID\(^\omega \). In terms of the expressive power of the underlying logic, TC (assuming pairs) subsumes the inductive machinery underlying LKID. This is because for any inductive predicate \( P \) of LKID, there is an \( \mathcal{L}_{RTC} \) formula \( \psi \) such that for every standard admissible structure \( M \) for \( \mathcal{L}_{RTC} \), \( P \) has the same interpretation as \( \psi \) under \( M \). This is due to Thm. 3 in \[11\] and the fact that the interpretation of \( P \) must necessarily be a recursively enumerable set. As for the converse inclusion, for any positive \( \mathcal{L}_{RTC} \) formula there is a production of a corresponding LKID inductive definition. However, the RTC operator can also
be applied on complex formulas (whereas LKID productions only consider atomic predicates). This indicates that TC might be more expressive. It was noted in [6, p. 1180] that complex formulas may be handled by stratifying the theory of LKID, similar to [15], but the issue of relative expressiveness of the resulting theory is not addressed. While we strongly believe it is the case that TC is strictly more expressive than the logic of LKID, proving so is left for future work. Also left for future research is establishing the comparative status of the corresponding formal proof systems.

In addition to the open question of the (in)equivalence of \( RTC_G \) and \( CRTC_G^\omega \) in the general case, discussed in Section 4.2, several other questions and directions for further study naturally arise from the work of this paper. We would like to implement our cyclic proof system in order to investigate the practicalities of using TC logic to support automated inductive reasoning. Also, the uniformity provided by the transitive closure operator suggests that coinductive reasoning can be incorporated into the formal system. Determining whether, and to what extent, this is indeed the case is left for future work.

References


