

Multiple-conclusion Rules, Hypersequents Syntax and Step Frames

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Abstract

We investigate proof theoretic properties of logical systems via algebraic methods. We introduce a calculus for deriving multiple-conclusion rules and show that it is a Hilbert style counterpart of hypersequent calculi. Using step-algebras we develop a criterion establishing the bounded proof property and finite model property for these systems. Finally, we show how this criterion can be applied to universal classes axiomatized by certain canonical rules, thus recovering and extending known results from both semantically and proof-theoretically inspired modal literature.

Keywords: Multiple-conclusion rules, hypersequents, step algebras, step frames.

1 Introduction

In this paper we continue the proof theoretic investigations of modal logic via algebraic methods which started in [5,4]. In [5,4] the *bounded proof property* (the *bpp*), which is a kind of analytic subformula property, was introduced as a measurement of robustness of proof systems. An algebraic criterion was developed in [5,4] establishing whether a modal system axiomatized by standard rules possesses the *bpp*. Here we extend this research in two directions. First, we investigate more expressive proof systems axiomatized by multiple-conclusion rules for which we develop equivalent systems via hypersequent calculi and prove for them an algebraic criterion for the *bpp*. Second, for a large class of logics (stable logics) we systematically design proof systems that have the *bpp* (see Section 5). Thus, we are at a position to not only check whether a system is robust, but also to *design robust proof systems*, by finding appropriate rules.

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Multiple-conclusion rules recently gained attention in the modal logic literature (see e.g., [18, 16, 3]), because they constitute an essential tool for investigating classes of algebras beyond varieties and because canonical formulae axiomatizations can be nicely developed within this framework. On the other hand and from a completely different research perspective, the proof-theoretic oriented community realized that standard sequent formalisms are insufficient to handle complex logics and moved to more expressive hypersequent calculi (compare for instance the simplicity of communication rules used for the logics of linear frames developed in [1] with the more complex systems needed for cut elimination in the traditional context [10, 14]).

In this paper we connect multiple-conclusion rules and hypersequent calculi. To our best knowledge, no explicit calculus for multiple-conclusion rules has been proposed so far. Note that for semantic investigations such as [16, 3], it is in fact sufficient to specify *abstractly* the properties that a rule system (seen just as a set of rules) should satisfy. On the other hand, a specific calculus for multiple-conclusion rules is needed if we want to make a close comparison with the hypersequent approach. This calculus will play the role of a Hilbert calculus for hyperformulae, i.e., for the syntactic components of a hypersequent. We will introduce such a calculus and investigate it using the techniques developed in [5, 4]. These techniques, based on semantic analysis of ‘step’ structures, have been shown to be rather effective in establishing the bpp. Our long-term proposal is to apply these techniques to obtain the bpp and the *finite model property* (the fmp), thus also decidability, for logics axiomatized by canonical formulae. In this paper, we report a first success in this direction, already covering the bpp and fmp for a continuum of logics, including some of those recently analyzed in [19] via the hypersequent approach.

Proofs of the results from Section 2 will be deferred to the appendix. Proofs of the results from Sections 3 and 4 (requiring routine adjustments from the corresponding proofs in [5, 4]) are included only in the Technical Report [6].

2 A calculus for derived multiple-conclusion rules

Modal formulae are built from propositional variables x, y, \dots by using the Booleans ($\neg, \wedge, \vee, \rightarrow, 0, 1$) and modal operators (\diamond, \square). For simplicity, we take \neg, \wedge, \diamond as primitive connectives, the remaining ones being defined in the customary way (in particular, \square is defined as $\neg\diamond\neg$). We shall also use parameters a, b, \dots instead of variables whenever we want to stress that uniform substitution does not apply to them. Underlined letters stand for tuples of unspecified length and formed by distinct elements, thus for instance, we may use \underline{x} for a tuple such as x_1, \dots, x_n . When we write $\phi(\underline{x})$ we want to stress that ϕ contains at most the variables \underline{x} (and no parameters) and similarly when we write $\phi(\underline{a})$ we want to stress that ϕ contains at most the parameters \underline{a} (and no variables). The same convention applies to sets of formulae: if Γ is a set of formulae and we write $\Gamma(\underline{a})$, we mean that all formulae in Γ are of the kind $\phi(\underline{a})$, etc. We may occasionally replace variables with parameters in a formula: for this, we use the following self-explanatory notation. For a formula $\phi(\underline{x})$ we write $\phi(\underline{a})$

to mean that $\phi(\underline{a})$ is obtained from $\phi(\underline{x})$ by replacing $\underline{x} = x_1, \dots, x_n$ (simultaneously and respectively) by $\underline{a} = a_1, \dots, a_n$. The modal complexity (or the modal degree) of a formula ϕ counts the maximum number of nested modal operators in ϕ (the precise definition is by an obvious induction).

We recall some background on modal algebras, see e.g., [8, Sec. 5.2] or [9, Sec. 7.6] for more details. A *modal algebra* $\mathfrak{A} = (A, \diamond)$ is a Boolean algebra A endowed with a unary operator \diamond satisfying $\diamond(x \vee y) = \diamond x \vee \diamond y$, $\diamond 0 = 0$. Notice that, here and elsewhere, we use the same name for a connective and the corresponding operator in modal algebras (thus, for instance, 0 is zero, \vee is join, etc.). In this way, propositional formulae can be identified with *terms* in the first order language of modal algebras.

From the semantic side, we have the notion of a frame. A *frame* $\mathfrak{F} = (W, R)$ is a set W endowed with a binary relation R . The *dual* of a frame $\mathfrak{F} = (W, R)$ is the modal algebra $\mathfrak{F}^* = (\wp(W), \diamond_R)$, where $\wp(W)$ is the powerset Boolean algebra and \diamond_R is the semilattice morphism associated with R . The latter is defined as follows: for $S \subseteq W$, we have $\diamond_R(S) = \{w \in W \mid R(w) \cap S \neq \emptyset\}$ (here $R(w)$ denotes $\{v \in W \mid (w, v) \in R\}$). It should be noticed that there is a real duality (in the categorical sense) between modal algebras and frames only if we restrict to finite modal algebras and finite frames. If we want a full duality working for arbitrary modal algebras, we must introduce some topological structures on our frames (see, e.g., [8, Sec. 5.5], [9, Sec. 7.5], [17, Ch. 4] or [20]). For the purposes of this paper, however, the duality between finite frames and finite modal algebras will suffice.

2.1 Multiple-conclusion rules

Normal modal logics are an adequate formalism to describe equational classes of modal algebras. However, in this paper we are interested in more general classes. A class of modal algebras is said to be:

- (i) a *variety* iff it is the class of models of a set of equations, i.e., of sentences of the kind $\forall \underline{x} \bigwedge_{i=1}^n \phi_i(\underline{x}) = 1$, where the ϕ_i are modal formulae (aka terms in the first order language of modal algebras);
- (ii) a *quasi-variety* iff it is the the class of models of a set of implications of equations, i.e., of sentences of the kind $\forall \underline{x} (\bigwedge_{i=1}^n \phi_i(\underline{x}) = 1 \rightarrow \psi(\underline{x}) = 1)$, where $\phi_1, \dots, \phi_n, \psi$ are modal formulae;
- (iii) a *universal class* iff it is the class of models of a set of clauses, i.e., of sentences of the kind $\forall \underline{x} (\bigwedge_{i=1}^n \phi_i(\underline{x}) = 1 \rightarrow \bigvee_{j=1}^m \psi_j(\underline{x}) = 1)$, where $\phi_1, \dots, \phi_n, \psi_1, \dots, \psi_m$ are modal formulae.

In order to describe universal classes within a propositional modal language, we shall use multiple-conclusion rules. A *multiple-conclusion rule* (or just a *rule*) is a pair of finite sets of formulae $\langle \Gamma, S \rangle$. If $\Gamma = \{\gamma_1, \dots, \gamma_n\}$, $S = \{\delta_1, \dots, \delta_m\}$, we write the rule $\langle \Gamma, S \rangle$ as Γ/S or as

$$\frac{\gamma_1, \dots, \gamma_n}{\delta_1 \mid \dots \mid \delta_m} (R)$$

The formulae $\Gamma = \{\gamma_1, \dots, \gamma_n\}$ are said to be the *premises* of the rule (R) and the formulae $S = \{\delta_1, \dots, \delta_m\}$ are said to be the *conclusions* of the rule (R) . The multiple-conclusion rule (R) is said to be an *inference rule* or a *single-conclusion rule* iff $m = 1$, i.e., iff it has a single conclusion. A modal algebra $\mathfrak{A} = (A, \diamond)$ validates the multiple-conclusion rule (R) iff it is a model of the clause $\forall \underline{x} (\bigwedge_{i=1}^n \phi_i(\underline{x}) = 1 \rightarrow \bigvee_{j=1}^m \psi_j(\underline{x}) = 1)$. A frame $\mathfrak{F} = (W, R)$ validates (R) iff its dual algebra \mathfrak{F}^* does.

We recall the notion of a rule system from [16]:

Definition 2.1 A set of multiple-conclusion rules K is said to be a *rule system* iff it satisfies the following conditions for every formula ϕ and for every finite sets of formulae Γ, Γ', S, S' :

- (i) $\phi/\phi \in K$;
- (ii) if $\Gamma/S, \phi \in K$ and $\Gamma, \phi/S \in K$, then $\Gamma/S \in K$;
- (iii) if $\Gamma/S \in K$ then $\Gamma, \Gamma'/S, S' \in K$;
- (iv) if $\Gamma/S \in K$ then for every substitution σ , we have that $\Gamma\sigma/S\sigma \in K$.

Above we used obvious conventions about set-theoretic union of finite sets of formulae (e.g., Γ, ϕ stands for $\Gamma \cup \{\phi\}$, moreover Γ, Γ' stands for $\Gamma \cup \Gamma'$, etc.). In addition, we used $\Gamma\sigma$ to denote the set resulting from the application of σ to all members of Γ .

Definition 2.2 A (normal) *modal rule system* is a rule system containing classical tautologies and the distribution schema $\Box(\alpha_1 \rightarrow \alpha_2) \rightarrow (\Box\alpha_1 \rightarrow \Box\alpha_2)$ (as single-conclusion 0-premises rules) as well as necessitation $(\alpha/\Box\alpha)$ and modus ponens $(\alpha, \alpha \rightarrow \beta/\beta)$ rules.

We say that a set of rules K *entails* or *derives* a rule Γ/S (written $K \vdash \Gamma/S$) iff Γ/S belongs to the smallest modal rule system $[K]$ containing K . The following algebraic completeness theorem is proved in [16] (but follows also from our results below):

Theorem 2.3 *Let K be a set of multiple-conclusion rules. Then $K \vdash \Gamma/S$ iff every modal algebra validating all rules in K validates also Γ/S .*

2.2 Hyperformulae and hyperproofs

We now design a calculus for recognizing syntactically the relation $K \vdash \Gamma/S$. We shall actually give two equivalent versions of such a calculus, the latter to be seen just as a Hilbert-style analogue of the well-known hypersequent calculi [1].

A *hyperformula* is a finite set of propositional formulae written in the form

$$\alpha_1 \mid \dots \mid \alpha_n. \quad (1)$$

We use letters S, S_1, S', \dots for hyperformulae; the notation $S \mid S'$ means set union and $S \mid \alpha$ and $\alpha \mid S$ stand for $S \mid \{\alpha\}$ and $\{\alpha\} \mid S$, respectively.

Definition 2.4 Let Γ be a set of propositional modal formulae and let K be a set of multiple-conclusion rules. A *K-hyperproof* (or a *K-derivation* or just

a *derivation*) under assumptions Γ is a finite list of hyperformulae S_1, \dots, S_n such that each S_i in it matches one of the following requirements:

- (i) S_i is of the kind $\alpha \mid S$, where $\alpha \in \Gamma$ or α is a tautology or α is an instance of the distribution schema;
- (ii) S_i is obtained from hyperformulae preceding it by applying a rule from K or the necessitation rule or the modus ponens rule.

We write $\Gamma \vdash_K S$ to mean that there is a K -derivation ending with S .

An important remark is in order for (ii): when we say that S_i is obtained by applying an inference rule, *we include uniform substitution and weakening in the application of the rule*. Thus, if the rule is

$$\frac{\gamma_1, \dots, \gamma_n}{\delta_1 \mid \dots \mid \delta_m} (R)$$

when we say that S_i is obtained from (R) , we mean that there are a hyperformula S and a substitution σ such that S_i is of the kind $S \mid \delta_1\sigma \mid \dots \mid \delta_m\sigma$ and that there are $j_1, \dots, j_n < i$ such that S_{j_1} is of the kind $S \mid \gamma_1\sigma$, and \dots and S_{j_n} is of the kind $S \mid \gamma_n\sigma$ (of course, this applies also to the case $n = 0$, i.e., to zero-premisses rules).

Theorem 2.5 *Let K be a set of multiple-conclusion rules. Then $\Gamma \vdash_K S$ iff the multiple-conclusion rule Γ/S is valid in every modal algebra validating K .*

Corollary 2.6 *Let K be a set of multiple-conclusion rules. For each multiple-conclusion rule Γ/S , we have $K \vdash \Gamma/S$ iff $\Gamma \vdash_K S$.*

Notice that Theorem 2.3 follows from Corollary 2.6 and Theorem 2.5.

2.3 Hypersequent syntax

A *sequent* is a pair of finite sets of formulae written as $\Gamma \Rightarrow \Delta$ and a *hypersequent* is a finite set of sequents written as

$$\Gamma_1 \Rightarrow \Delta_1 \mid \dots \mid \Gamma_n \Rightarrow \Delta_n. \quad (2)$$

In this paper, we are investigating proof theoretic facts that only depends on the modal degree of formulae and on the modal degree of formulae occurring within proofs, thus *we view a sequent $\Gamma \Rightarrow \Delta$ as the formula $\bigwedge \Gamma \rightarrow \bigvee \Delta$ and a hypersequent (2) as the hyperformula*

$$\bigwedge \Gamma_1 \rightarrow \bigvee \Delta_1 \mid \dots \mid \bigwedge \Gamma_n \rightarrow \bigvee \Delta_n. \quad (3)$$

Still, there is an important difference between hyperproofs according to Definition 2.4 and hypersequent calculi e.g., in [1]: once translated into our formalism, the difference is in the possibility of using rules having hyperformulae (and not just formulae) as premisses. We show here that this difference is immaterial because we can translate these more general rules and proofs into our formalism. The translation is effective, does not increase the modal degree of formulae involved in the proofs, but might be harmful for complexity.

We first introduce the definitions needed to make the comparison. A *hyperrule* is a $n * 1$ -tuple of hyperformulae, written as $S_1, \dots, S_k/S$. If H is a set of hyperrules, Γ is a set of hyperformulae and S is a hyperformula, we say that S is *provable from Γ in H* , written $\Gamma \Vdash_H S$ iff there exists a finite list of hyperformulae S_1, \dots, S_n (called a *derivation*) such that each S_i in it matches one of the following requirements:

- (i) S_i is a hyperformula containing a member of Γ , or a tautology, or a formula of the form $\Box(\alpha_1 \rightarrow \alpha_2) \rightarrow (\Box\alpha_1 \rightarrow \Box\alpha_2)$;
- (ii) S_i is obtained from hyperformulae preceding it by applying modus ponens rule $\alpha, \alpha \rightarrow \beta/\beta$, necessitation rule $\alpha/\Box\alpha$, or a hyperrule from H .

Again, ‘to apply a rule $S_1, \dots, S_k/S$ to get S_i ’ means that there is a substitution σ such that S_i is of the kind $\tilde{S} | S\sigma$ and that there are $j_1, \dots, j_n < i$ such that S_{j_1} is of the kind $\tilde{S} | S_1\sigma$ and ... and S_{j_n} is of the kind $\tilde{S} | S_n\sigma$.³

Proposition 2.7 *Let H be a finite set of hyperrules. Then it is possible to produce a set of rules K such that for all Γ, \tilde{S} we have $\Gamma \Vdash_H \tilde{S}$ iff $\Gamma \vdash_K \tilde{S}$.*

Proof. (Sketch, see the appendix for full details) Consider a hyperrule $S_1, \dots, S_k/S$ from H : to obtain K , we simply replace it with the set of rules $\gamma(S_1), \dots, \gamma(S_n)/S$, varying γ among the functions that pick one formula from each S_i , for each $i = 1, \dots, n$. \square

Next we give a few examples. In order to make a more direct link with the current literature, we will use the hypersequent syntax (Gentzen standard sequent rules for classical logic, as well as external structural rules will be always implicitly assumed below). Since in this paper we are interested only in investigating modal degrees of formulae and proofs, in most cases the metavariables Γ, Δ, \dots occurring in the sequent notation below can be replaced by single formulae (hence the rules in Examples 2.8-2.9 can be seen as single rules,⁴ not as schemata standing for infinitely many rules).

Example 2.8 An adequate calculus for **S4** comprises the following two rules (taken from [15])

$$\frac{\Box\Gamma \Rightarrow A_1 \mid \dots \mid \Box\Gamma \Rightarrow A_n}{\Gamma', \Box\Gamma \Rightarrow \Delta, \Box A_1, \dots, \Box A_n} (\Rightarrow\Box)$$

$$\frac{\Box A, A, \Gamma \Rightarrow \Delta}{\Box A, \Gamma \Rightarrow \Delta} (T)$$

where, if $\Gamma = \{\phi_1, \dots, \phi_n\}$, then $\Box\Gamma$ stands for $\{\Box\phi_1, \dots, \Box\phi_n\}$.

³ This notion of a derivation avoids the introduction of side components (in the sense of [1]) when specifying rules: in fact, the side component S is introduced directly when applying the rule.

⁴ This is not the case for Example 2.10 because $\Box\Delta$ on the right of \Rightarrow cannot be replaced by a single formula.

Example 2.9 Let us now consider the universal class of *prime S4.3* algebras: these are the modal algebras validating the above rules and satisfying in addition the clause

$$\forall x \forall y (\Box x \leq \Box y \text{ or } \Box y \leq \Box x).$$

To axiomatize this class, we can add to the above rules the further rule

$$\frac{\tilde{\Gamma}, \Box\Gamma, \Box\Gamma' \Rightarrow \Delta \quad \tilde{\Gamma}', \Box\Gamma', \Box\Gamma \Rightarrow \Delta'}{\tilde{\Gamma}, \Box\Gamma \Rightarrow \Delta \mid \tilde{\Gamma}', \Box\Gamma' \Rightarrow \Delta'} \text{ (Dich)}$$

taken from [15]. Rule (Dich) is nothing but a variant of the communication rule introduced in [1].

Example 2.10 For prime **S5** algebras, we can add to **S4**-rules the following rule taken from [1]

$$\frac{\Box\Gamma, \Gamma' \Rightarrow \Box\Delta, \Delta'}{\Box\Gamma \Rightarrow \Box\Delta \mid \Gamma' \Rightarrow \Delta'} \text{ (S5)}$$

3 Bounded proofs and step frames

From now on, we shall make exclusive reference to the calculus explained in Definition 2.4. We call a *modal calculus* (or simply a *calculus*) a set of multiple-conclusion rules *where only formulae of modal degree at most one occur*.⁵

When we write $\Gamma \vdash_K^n S$ we mean that there is a K -hyperproof under assumptions Γ (see Definition 2.4) in which only formulae of modal complexity at most n occur. We are mostly interested in the semantic characterization of the following property:

Definition 3.1 We say that a calculus K has the *bounded proof property* (the bpp, for short) iff for every hyperformula S of modal complexity at most n and for every Γ containing only formulae of modal complexity at most n , we have

$$\Gamma \vdash_K S \quad \Rightarrow \quad \Gamma \vdash_K^n S.$$

A remarkable consequence of the bpp is explained in the following:

Proposition 3.2 *If a modal calculus K consisting of finitely many rules enjoys the bpp, then the relation $\Gamma \vdash_K S$ (as well as the derivability of rules in K , see Corollary 2.6) is decidable.*

Proof. Since K has the bpp, it is sufficient to prove that $\Gamma \vdash_K^n S$ is decidable, where n is as in Definition 3.1. We show the decidability of the relation $\Gamma \vdash_K^n S$, by bounding the set of formulae that may occur as components ψ_1, \dots, ψ_m of a hyperformula $\psi_1 \mid \dots \mid \psi_m$ included in a derivation witnessing $\Gamma \vdash_K^n S$. Notice

⁵ This property can be assumed without loss of generality, by applying the transformation suggested in [4] (that transformation does not increase the modal degree of proofs). In [4] another property is assumed on rules (namely that variables occurring in them have occurrences inside a modal operator). This property was assumed there to simplify the definition of evaluation in step algebras, but in the present more general context it can have unclear side effects, so we prefer not to assume it anymore.

first that we can freely suppose that only the variables X occurring in Γ, S appear in such a derivation. This is because extra variables can be uniformly replaced by, say, 0.

Let us say that ϕ is equivalent to ψ (written $\phi \approx \psi$) iff $\phi \leftrightarrow \psi$ is provable in the minimum normal modal system (i.e., by using tautologies, modus ponens and necessitation). Notice that the relation \approx is decidable and that, whenever $\phi \approx \psi$ holds, the replacement rule

$$\frac{\phi}{\psi} (Repl)$$

is derivable in K . In fact, $(Repl)$ can be simulated by a derivation having modal degree at most n in case the modal degrees of ϕ, ψ are at most n . In addition, it is well-known (e.g., from the theory of normal forms [12]) that there are finitely many \approx -equivalence classes of the formulae ϕ having at most degree n and built up from the finite set of propositional variables X . We can effectively fix a representative for each of these classes. Let C_n be the set of such representatives.

A canonical substitution σ for a rule $R \in K$ is a substitution σ associating with a variable x a formula $\sigma(x)$ in $C_n \cup C_{n-1}$ in such a way that the formulae occurring as components of hyperformulae from $R\sigma$ are of complexity at most n . Thus, recalling that rules have complexity one, $\sigma(x)$ must be in C_{n-1} if x has an occurrence in R inside a modal operator and $\sigma(x)$ must be in C_n , otherwise. A canonical instance of $R \in K$ is an instance $R\sigma$ of R via a canonical substitution σ . Notice that there are only finitely many canonical instances.

We let Θ be the set of formulae which either occur in $\Gamma \cup S$ or are in C_n or occur in a canonical instance of a rule in K . Again, Θ is finite and has modal complexity at most n .

By induction, we transform a derivation π from K in which formulae of degree at most n occur into a derivation π' in which only members of Θ occur. When building π' , we make use also of the replacement rule $(Repl)$ introduced above.

The construction of π' is easy for the base case of derivations of length 1. Let us consider the inductive case of a derivation π ending with the application of a rule R from K (we include also the case in which R is modus ponens or necessitation). For simplicity, let R have a single premise, i.e., R is ϕ/T . Suppose that, in π , the rule is used to infer $T\sigma \mid \tilde{S}$ from $\phi\sigma \mid \tilde{S}$. In π' we derive by induction $\psi \mid \tilde{S}'$, where the formulae in $\psi \mid \tilde{S}'$ belongs to Θ and are equivalent to the formulae in $\phi\sigma \mid \tilde{S}$. Let σ' be a canonical substitution such that $\sigma(x)$ is equivalent to $\sigma'(x)$ for every variable x occurring in R . Then $\phi\sigma$ is equivalent to $\phi\sigma'$ and the formulae in $T\sigma$ are equivalent to the formulae in $T\sigma'$, respectively. By the replacement rule, we can infer $\phi\sigma' \mid \tilde{S}'$ in π' (because $\phi'\sigma \approx \phi\sigma \approx \psi$). Then we infer $T\sigma' \mid \tilde{S}'$ via the rule R . The latter is a hyperformula whose components are all in Θ and are equivalent to the components of the hyperformula $T\sigma \mid \tilde{S}$ inferred by π .

Thus, checking whether $\Gamma \vdash_K^n S$ holds is reduced to checking whether there

is a derivation using formulas from a finite set Θ . Note also that we can assume that a given hyperformula occurs at most once in a derivation (because occurrences following the first one can be removed). The result follows. \square

The following proposition shows that we can limit our consideration to formulae of complexity 1 when checking the bpp.

Proposition 3.3 *A calculus K has the bounded proof property iff for every hyperformula S of modal complexity at most 1 and for every Γ containing only formulae of modal complexity at most 1, we have $\Gamma \vdash_K S \Rightarrow \Gamma \vdash_K^1 S$.*

In the following, we shall adopt the equivalent formulation of the bpp suggested by the above proposition. We shall call finite sets Γ of formulae of modal complexity at most 1, *finite presentations*. It is useful to use parameters (see Section 2) to name the variables occurring in a finite presentation Γ : this is because in a K -hyperproof under assumptions Γ , the formulae in Γ are introduced in the derivation as they are (no substitution applies to them), whereas substitutions are applied to rules in K . Thus, we write $\Gamma(\underline{a})$ to emphasize that at most the parameters \underline{a} occur in Γ and $\Gamma(\underline{a}) \vdash_K S(\underline{a})$ to emphasize that the tuple \underline{a} includes all parameters occurring in both Γ, S .

3.1 Conservative one-step algebras and one-step frames

We first recall the definition of one-step modal algebras and one-step frames from [11] and [7], and define conservative one-step modal algebras and one-step frames.

Definition 3.4 A *one-step modal algebra* is a quadruple $\mathcal{A} = (A_0, A_1, i_0, \diamond_0)$, where A_0, A_1 are Boolean algebras, $i_0 : A_0 \rightarrow A_1$ is a Boolean morphism, and $\diamond_0 : A_0 \rightarrow A_1$ is a semilattice morphism (i.e., it preserves only $0, \vee$). The algebras A_0, A_1 are called the *source* and the *target* Boolean algebras of the one-step modal algebra \mathcal{A} . We say that \mathcal{A} is *conservative* iff i_0 is injective and the union of the images $i_0(A_0) \cup \diamond_0(A_0)$ generates A_1 as a Boolean algebra.

From the dual semantic point of view we have the following:

Definition 3.5 A *one-step frame* is a quadruple $\mathcal{S} = (W_1, W_0, f, R)$, where W_0, W_1 are sets, $f : W_1 \rightarrow W_0$ is a map and $R \subseteq W_1 \times W_0$ is a relation between W_1 and W_0 . We say that \mathcal{S} is *conservative* iff f is surjective and the following condition is satisfied for all $w_1, w_2 \in W_1$:

$$f(w_1) = f(w_2) \ \& \ R(w_1) = R(w_2) \ \Rightarrow \ w_1 = w_2. \quad (4)$$

Similarly to the case of Kripke frames, above we used the notation $R(w_1)$ to mean the set $\{v \in W_0 \mid (w_1, v) \in R\}$ (and similarly for $R(w_2)$). The dual of a finite one-step frame $\mathcal{S} = (W_1, W_0, f, R)$ is the one-step modal algebra $\mathcal{S}^* = (\wp(W_0), \wp(W_1), f^*, \diamond_R)$, where f^* is the inverse image operation and \diamond_R is the semilattice morphism associated with R . The latter is defined as follows: for $S \subseteq W_0$, we have $\diamond_R(S) = \{w \in W_1 \mid R(w) \cap S \neq \emptyset\}$. Conservativity also

carries over from one-step frames to one-step modal algebras (see [4] for a proof of the following proposition):

Proposition 3.6 *A finite one-step frame \mathcal{S} is conservative iff its dual one-step modal algebra \mathcal{S}^* is conservative.*

To complete our list of definitions, let us observe that a one-step modal algebra $\mathcal{A} = (A_0, A_1, i_0, \diamond_0)$ in which we have $A_0 = A_1$ and $i_0 = id$ is nothing but a modal algebra. Similarly, a one-step frame $\mathcal{S} = (W_1, W_0, f, R)$ where we have $W_0 = W_1$ and $f = id$ is just a frame. For clarity, we shall sometimes call modal algebras and frames *standard* or *plain* modal algebras and frames, respectively.

3.2 Inference validation in step algebras

We spell out what it means for a one-step modal algebra and a one-step frame to validate a modal calculus K and a finite presentation Γ (the definition requires little modifications with respect to [4, 5] because we do not restrict to reduced rules).

Let us fix two finite sets of variables $\underline{x} = x_1, \dots, x_n$, $\underline{y} = y_1, \dots, y_m$ and a finite set of parameters $\underline{a} = a_1, \dots, a_m$ (either \underline{x} , \underline{y} or \underline{a} can be empty). An *\underline{a} -augmented* one-step modal algebra $\mathcal{A} = (A_0, A_1, i_0, \diamond_0, \underline{a})$ is a one-step modal algebra together with displayed elements $\underline{a} = \mathbf{a}_1, \dots, \mathbf{a}_m \in A_0$ (these elements will interpret parameters).

Given an \underline{a} -augmented one-step modal algebra as above, an \mathcal{A} -*valuation* is a map associating with each variable $x_i \in \underline{x}$ an element $\mathbf{v}(x_i) \in A_0$ and with each variable $y_j \in \underline{y}$ an element $\mathbf{v}(y_j) \in A_1$. For every formula $\phi(\underline{x})$ of complexity 0, we define $\phi^{\mathbf{v}0} \in A_0$ as follows:

$$\begin{aligned} x_i^{\mathbf{v}0} &= \mathbf{v}(x_i) \quad (\text{for every variable } x_i \in \underline{x}); & a_i^{\mathbf{v}0} &= \mathbf{a}_i \quad (a_i \in \underline{a}); \\ (\phi_1 * \phi_2)^{\mathbf{v}0} &= \phi_1^{\mathbf{v}0} * \phi_2^{\mathbf{v}0} \quad (* = \wedge, \vee); & (\neg\phi)^{\mathbf{v}0} &= \neg(\phi^{\mathbf{v}0}). \end{aligned}$$

For every $\psi(\underline{x}, \underline{y})$ of complexity at most 1 in which the \underline{y} do not have occurrences within the scope of a modal operator, $\psi^{\mathbf{v}1} \in A_1$ is defined as follows:

$$\begin{aligned} x_i^{\mathbf{v}1} &= i_0(\mathbf{v}(x_i)) \quad (\text{for every variable } x_i \in \underline{x}); & a_i^{\mathbf{v}1} &= i_0(\mathbf{a}_i) \quad (a_i \in \underline{a}); \\ y_j^{\mathbf{v}1} &= \mathbf{v}(y_j) \quad (\text{for every variable } y_j \in \underline{y}); & (\diamond\phi(\underline{x}))^{\mathbf{v}1} &= \diamond(\phi^{\mathbf{v}0}); \\ (\psi_1 * \psi_2)^{\mathbf{v}1} &= \psi_1^{\mathbf{v}1} * \psi_2^{\mathbf{v}1} \quad (* = \wedge, \vee); & (\neg\psi)^{\mathbf{v}1} &= \neg(\psi^{\mathbf{v}1}). \end{aligned}$$

It is immediate to see by induction that for every formula $\phi(\underline{x})$ of complexity 0 (in which the \underline{y} do not occur), we have $\phi^{\mathbf{v}1} = i_0(\phi^{\mathbf{v}0})$.

Definition 3.7 Suppose that the formulae $\delta_1(\underline{x}, \underline{y}), \dots, \delta_k(\underline{x}, \underline{y}), \gamma_1(\underline{x}, \underline{y}), \dots, \gamma_l(\underline{x}, \underline{y})$ have modal degree at most one and that the variables \underline{y} are the variables not occurring in them inside the scope of a modal operator. We say that a one-step modal algebra \mathcal{A} *validates the multiple-conclusion rule*

$$\frac{\gamma_1, \dots, \gamma_l}{\delta_1 \mid \dots \mid \delta_m} (R)$$

iff for every \mathcal{A} -valuation \mathbf{v} , we have that if $(\phi_1^{\mathbf{v}1} = 1$ and \dots and $\phi_m^{\mathbf{v}1} = 1)$, then $(\gamma_1^{\mathbf{v}1} = 1$ or \dots or $\gamma_n^{\mathbf{v}1} = 1)$. We say that \mathcal{A} *validates a modal calculus* K (written $\mathcal{A} \models K$) iff \mathcal{A} validates all inferences from K .

Notice that it might well be that K_1 and K_2 are equivalent (in the sense that rules from K_1 are derivable in K_2 and vice versa), but that only one of them is validated by a given \mathcal{A} . This phenomenon, however, cannot happen in case \mathcal{A} is standard (i.e., it is a modal algebra).

For formulae $\phi(\underline{a})$ where the variables $\underline{x}, \underline{y}$ do not occur, the valuation \mathbf{v} is not relevant. Thus, in such cases, we may write $\phi^{\mathbf{a}0}, \phi^{\mathbf{a}1}$ instead of $\phi^{\mathbf{v}0}, \phi^{\mathbf{v}1}$, respectively, to stress the fact that the augmentation \mathbf{a} is the essential part of the definition. We write $\mathcal{A} \models \phi(\underline{a})$ for $\phi^{\mathbf{a}1} = 1$ and $\mathcal{A} \models S(\underline{a})$ iff there is a $\phi \in S$ such that $\mathcal{A} \models \phi$. We say that \mathcal{A} *validates the presentation* Γ (in symbols, $\mathcal{A} \models \Gamma(\underline{a})$) iff we have that $\mathcal{A} \models \phi(\underline{a})$ for all $\phi(\underline{a}) \in \Gamma$.

The notion of an \mathcal{S} -valuation for a one-step frame \mathcal{S} is the expected one, namely \mathbf{v} is an \mathcal{S} -valuation iff it is an \mathcal{S}^* -valuation. In the same way the other notions introduced above (augmentation, $\phi^{\mathbf{v}0}, \phi^{\mathbf{v}1}$, validation of a presentation, of an inference, of an axiomatic system) can be extended by duality to one-step frames.

Example 3.8 For the systems **S4**, **S4.3**, **S5**, it can be shown (by applying the ‘step’ variant of modal correspondence theory [5, 4]) that a conservative one-step frame $\mathcal{S} = (W_1, W_0, f, R)$

- validates the rules of Example 2.8 iff it is step-transitive and step-reflexive, where the latter means $f \subseteq R$ and the former means $R \subseteq f \circ \geq_R$ (here \circ is relation composition and $w_1 \geq_R w_2$ is defined to be $R(w_1) \supseteq R(w_2)$);
- validates the rules of Example 2.9 iff it is step-transitive, step-reflexive and step-linear, where the latter means $\forall w_1, w_2 \in W_1 (R(w_1) \subseteq R(w_2) \text{ or } R(w_2) \subseteq R(w_1))$;
- validates the rules of Example 2.10 iff we have $R(w) = W_0$ for all $w \in W_1$.

We can specialize our notions to standard modal algebras and frames. An \underline{a} -augmentation in a modal algebra $\mathfrak{A} = (A, \diamond)$ is a tuple \underline{a} of elements from the support of A , matching the length of \underline{a} . For frames $\mathfrak{F} = (W, R)$, we dually take a tuple from $\wp(W)$, i.e., a tuple of subsets. Given such \underline{a} -augmentation, we can define $\mathfrak{A} \models \Gamma(\underline{a})$ and $\mathfrak{F} \models \Gamma(\underline{a})$ for a presentation $\Gamma(\underline{a})$, just specializing the above definitions (standard modal algebras and frames are special one-step modal algebras and frames). Notice that $\mathfrak{F} \models \Gamma(\underline{a})$ is *global* validity in terms of the Kripke forcing from the modal logic literature, see e.g., [17, Sec. 3.1].

Proposition 3.9 *Let $\mathcal{A} = (A, B, i, \diamond, \underline{a})$ be an augmented conservative one-step modal algebra that validates the modal calculus K and the presentation $\Gamma(\underline{a})$. Then, for every hyperformula $S(\underline{a})$, we have that $\Gamma \vdash_K^1 S$ implies $\mathcal{A} \models S$.*

4 Semantic characterizations of the bpp and the fmp

In this section we first introduce the morphisms of one-step modal algebras and one-step frames.

Definition 4.1 An *embedding* between one-step modal algebras $\mathcal{A} = (A_0, A_1, i_0, \diamond_0)$ and $\mathcal{A}' = (A'_0, A'_1, i'_0, \diamond'_0)$ is a pair of injective Boolean morphisms $h : A_0 \rightarrow A'_0$, $k : A_1 \rightarrow A'_1$ such that

$$k \circ i_0 = i'_0 \circ h \quad \text{and} \quad k \circ \diamond_0 = \diamond'_0 \circ h . \quad (5)$$

$$\begin{array}{ccc} A_0 & \xrightarrow{h} & A'_0 \\ i_0 \downarrow & & \downarrow i'_0 \\ A_1 & \xrightarrow{k} & A'_1 \end{array}$$

$$\begin{array}{ccc} A_0 & \xrightarrow{h} & A'_0 \\ \diamond_0 \downarrow & & \downarrow \diamond'_0 \\ A_1 & \xrightarrow{k} & A'_1 \end{array}$$

Notice that, when \mathcal{A}' is standard (i.e. $A'_1 = A'_0$ and $i'_0 = id$), h must be $k \circ i_0$ and (5) reduces to

$$k \circ \diamond_0 = \diamond'_0 \circ k \circ i_0 . \quad (6)$$

For frames we have the dual definition. In the definition below, we use \circ to denote relational composition: for $R_1 \subseteq X \times Y$ and $R_2 \subseteq Y \times Z$, we have $R_2 \circ R_1 := \{(x, z) \in X \times Z \mid \exists y \in Y ((x, y) \in R_1 \ \& \ (y, z) \in R_2)\}$. Notice that the relational composition applies also when one or both of R_1, R_2 is a function.

Definition 4.2 A *p-morphism* between step frames $\mathcal{F}' = (W'_1, W'_0, f', R')$ and $\mathcal{F} = (W_1, W_0, f, R)$ is a pair of surjective maps $\mu : W'_1 \rightarrow W_1$, $\nu : W'_0 \rightarrow W_0$ such that

$$f \circ \mu = \nu \circ f' \quad \text{and} \quad R \circ \mu = \nu \circ R' . \quad (7)$$

$$\begin{array}{ccc} W'_1 & \xrightarrow{\mu} & W_1 \\ f' \downarrow & & \downarrow f \\ W'_0 & \xrightarrow{\nu} & W_0 \end{array}$$

$$\begin{array}{ccc}
W'_1 & \xrightarrow{\mu} & W_1 \\
R' \downarrow & & \downarrow R \\
W'_0 & \xrightarrow{\nu} & W_0
\end{array}$$

Notice that, when \mathcal{F}' is standard (i.e., $W'_1 = W'_0$ and $f' = id$), ν must be $f \circ \mu$ and (7) reduces to

$$R \circ \mu = f \circ \mu \circ R'. \quad (8)$$

The following definitions introduce the semantic notions needed for our characterization of the bpp.

Definition 4.3 Let $\mathcal{A}_0 = (A_0, A_1, i_0, \diamond_0)$ be a one-step modal algebra. A *one-step extension* of \mathcal{A}_0 is a one-step modal algebra $\mathcal{A}_1 = (A_1, A_2, i_1, \diamond_1)$ satisfying $i_1 \circ \diamond_0 = \diamond_1 \circ i_0$. Dually, a *one-step extension* of the one-step frame $\mathcal{S}_0 = (W_1, W_0, f_0, R_0)$ is a one-step frame $\mathcal{S}_1 = (W_2, W_1, f_1, R_1)$ satisfying $R_0 \circ f_1 = f_0 \circ R_1$.

Definition 4.4 A class of one-step modal algebras has the *extension property* iff every conservative one-step modal algebra $\mathcal{A}_0 = (A_0, A_1, i_0, \diamond_0)$ in the class has an extension $\mathcal{A}_1 = (A_1, A_2, i_1, \diamond_1)$ such that i_1 is injective and \mathcal{A}_1 is also in the class. A class of one-step modal frames has the *extension property* iff every conservative one-step frame $\mathcal{S}_0 = (W_1, W_0, f_0, R_0)$ in the class has an extension $\mathcal{S}_1 = (W_2, W_1, f_1, R_1)$ such that f_1 is surjective and \mathcal{S}_1 is also in the class.

Theorem 4.5 *A modal calculus K has the bpp iff the class of finite one-step modal algebras (equivalently, the class of finite one-step frames) validating K has the extension property.*

The characterization of the bpp from Theorem 4.5 may not be easy to handle, because in practical cases one would like to avoid managing one-step extensions and would prefer to work with standard frames instead. This is possible, if we combine the bpp with the finite model property.

Definition 4.6 A modal calculus K has the (global) *finite model property*, the *fmp* for short, if for every tuple \underline{a} of parameters, for every finite set of formulae $\Gamma(\underline{a})$ and for every hyperformula $S(\underline{a})$ we have $\Gamma \not\vdash_K S$ iff there exists a finite \underline{a} -augmented modal algebra \mathfrak{A} such that $\mathfrak{A} \models K$, $\mathfrak{A} \models \Gamma(\underline{a})$ and $\mathfrak{A} \not\models S(\underline{a})$ (equivalently, iff there exists a finite \underline{a} -augmented Kripke frame \mathfrak{F} such that $\mathfrak{F} \models K$, $\mathfrak{F} \models \Gamma(\underline{a})$ and $\mathfrak{F} \not\models S(\underline{a})$).

We are ready for a characterization result:

Theorem 4.7 *A modal calculus K has both the bpp and the fmp iff every finite conservative one-step frame validating K is a p -morphic image of a finite frame validating K (equivalently, iff every finite conservative one-step modal algebra validating K has an embedding into a finite modal algebra validating K).*

Example 4.8 Theorem 4.7 applies to all Examples 2.8-2.10. The construction is the same in all cases and it is rather straightforward: given a finite conservative step frame $\mathcal{S} = (W_1, W_0, f, R)$ validating the rules of the calculus, we can define $\mathfrak{F}' = (W', R')$ and μ so that condition (8) is satisfied as follows:

$$W' := W, \quad \mu := id, \quad w_1 R' w_2 :\Leftrightarrow R(w_1) \supseteq R(w_2).$$

5 Modal stable rules

Canonical formulae for transitive modal logics and intuitionistic logic were introduced by Zakharyashev (see [9] for an overview) who proved that all transitive modal logics and all intermediate logics are axiomatizable by canonical formulae. Jeřábek [16] defined *canonical rules*, which are multiple-conclusion rules generalizing canonical formulas. Jeřábek used these rules for an alternative proof of decidability of admissible rules for intuitionistic logic and transitive modal logics **K4**, **S4**, **S4.3**, etc. However, there are non-transitive modal logics not axiomatizable by canonical formulae and rules. [3] defines *stable canonical rules*, which differ from Zakharyashev's canonical formulae and Jeřábek's canonical rules and proves that every modal logic (including non-transitive ones) is axiomatizable by these rules. In this section we will concentrate on logics axiomatizable by a special subclass of stable canonical rules.

Subframe logics are the logics whose frames are closed under taking subframes. Transitive subframe logics are axiomatizable by a special subclass of canonical formulae called *subframe formulae*, see, e.g., [9]. A similar restriction to stable canonical rules gives a class of *stable logics*. But stable logics are not necessarily transitive. Logics in this class are exactly the logics that are closed under relation-preserving (following [3] we will call such maps *stable*⁶) onto maps. Transitive subframe logics and stable logics enjoy the fmp. Transitive subframe logics enjoy the fmp because they admit selective filtration, and stable logics enjoy the fmp because they admit the standard filtration (see [3] for the details).

In this section we show that all stable logics admit an axiomatization that has the bounded proof property. As we will see below, stable canonical rules will not produce an axiomatization that has the bpp. However, we will modify these rules so that the obtained rules do possess the bpp. This provides a systematic method of producing infinitely many proof calculi that are good (enjoying the bpp) from the proof-theoretic point of view. We remark that Lahav [19] also considers a class of modal logics whose Kripke frames satisfy special first-order conditions. He introduces hypersequent calculi for these logics and proves that these calculi admit cut elimination. It is easy to see that the non-transitive logics studied in [19] are stable logics – their frame classes are closed under stable onto maps. Thus, the class of logics we investigate in this section extends the class of logics studied in [19] in the non-transitive

⁶ In [13] these maps are called *continuos*.

case.⁷ Note, however, that [19] studies cut elimination, whereas we work with the bpp only. Now, if cut elimination gives the subformula property as a by-product, the bpp follows trivially. The converse is not true: we might have the bpp without the subformula property.⁸ However, it should be noticed that the bpp is a strong evidence about the proof-theoretic robustness of a system and supplies a loose notion of analiticity which is sufficient for decidability and which can hold for a wide class of calculi, including cases where the design of cut-eliminating systems looks very problematic.

We start by recalling the definition of modal stable rules. Let $\mathfrak{F} = (F, R_F)$ be a finite frame. For every $a \in F$ we introduce a new propositional variable x_a . The *modal stable rule* of \mathfrak{F} is

$$\frac{\bigvee_{i=1}^n x_{a_i}, \bigwedge_{i \neq j} \neg(x_{a_i} \wedge x_{a_j}), \bigwedge_{i=1}^n (x_{a_i} \rightarrow \Box \bigvee_{b \in R_F(a_i)} x_b)}{\neg x_{a_1} \mid \dots \mid \neg x_{a_n}} (r_{\mathfrak{F}})$$

where we suppose that $F = \{a_1, \dots, a_n\}$.

A *stable embedding* of a modal algebra $\mathfrak{A} = (A, \Diamond)$ into a modal algebra $\mathfrak{B} = (B, \Diamond)$ is an injective Boolean morphism $\mu : A \rightarrow B$ such that we have $\Diamond \mu(x) \leq \mu(\Diamond x)$ for all $x \in A$. For a frame \mathfrak{F} we denote by \mathfrak{F}^* its dual modal algebra and for an algebra \mathfrak{A} we denote by \mathfrak{A}_* the descriptive frame dual to \mathfrak{A} . Recall that a map $f : W \rightarrow W'$ between standard frames (W, R) and (W', R') is called *stable* if for each $x, y \in W$ we have xRy implies $f(x)R'f(y)$.

The following proposition is proved in [3].

Proposition 5.1 *Let $\mathfrak{A} = (A, \Diamond)$ be a modal algebra. Then*

- (i) \mathfrak{A} does not validate $(r_{\mathfrak{F}})$ iff there is a stable embedding of \mathfrak{F}^* into \mathfrak{A} .
- (ii) \mathfrak{A} does not validate $(r_{\mathfrak{F}})$ iff there is a surjective stable map from \mathfrak{A}_* onto \mathfrak{F} .

Our aim is to show that all modal calculi axiomatized by rules of the kind $(r_{\mathfrak{F}})$ have the bounded proof property. Rules $(r_{\mathfrak{F}})$, however, are not good for the bpp, see the counterexample below. We replace rules $(r_{\mathfrak{F}})$ by modified versions.

For each $a \in F$ we just add an extra propositional variable r_a and define the new rule $(r_{\mathfrak{F}}^+)$ by⁹

$$\frac{\bigvee_{i=1}^n x_{a_i}, \bigwedge_{i \neq j} \neg(x_{a_i} \wedge x_{a_j}), \bigwedge_{i=1}^n (x_{a_i} \rightarrow \Box r_{a_i}), \bigwedge_{i=1}^n (r_{a_i} \rightarrow \bigvee_{b \in R_F(a_i)} x_b)}{\neg x_{a_1} \mid \dots \mid \neg x_{a_n}}$$

⁷ The transitive logic **K4** is not stable. The investigation of the bounded proof property of stable logics over **K4** is a topic for future research.

⁸ For example, we show that all stable logics have the bpp. This class contains a continuum of logics [3,2]. Whether all these logics admit natural calculi with cut elimination is an open question.

⁹ For uniformity, we prefer all the r_a to have at least one occurrence located inside a modal operator in the rule $(r_{\mathfrak{F}}^+)$. In order to obtain this, one might add premisses such as $\Box(r_a \vee \neg r_a)$.

Lemma 5.2 *Rules $(r_{\mathfrak{F}}^+)$ and $(r_{\mathfrak{F}})$ are inter-derivable.*

Proof. On the one hand, $(r_{\mathfrak{F}})$ can be obtained from $(r_{\mathfrak{F}}^+)$ by applying the substitution $r_{a_i} \mapsto \bigvee_{b \in R_F(a_i)} x_b$. On the other hand, we apply necessitation and distribution to the premise $\bigwedge_{i=1}^n (r_{a_i} \rightarrow \bigvee_{b \in R(a_i)} x_b)$ and then transitivity of implication to obtain $\bigwedge_{i=1}^n (x_{a_i} \rightarrow \square \bigvee_{b \in R_F(a_i)} x_b)$. \square

Notice that the above fragment of a derivation, when plugged into a hyperproof, may increase the modal degree (if the substitution used to apply the rule $(r_{\mathfrak{F}})$ replaces the x_b with formulae of, say, modal degree 1, we obtain formulae of modal degree 2 when we use $(r_{\mathfrak{F}})$ to simulate $(r_{\mathfrak{F}}^+)$). This is why $(r_{\mathfrak{F}}^+)$ is preferable to $(r_{\mathfrak{F}})$ from the point of view of the modal complexity analysis of proofs.

Theorem 5.3 *A modal calculus comprising only rules of the kind $(r_{\mathfrak{F}}^+)$ enjoys the bpp and fmp.*

Proof. We use Theorem 4.7. Let $\mathcal{S} = (W_1, W_0, f, R)$ be a finite conservative one-step frame validating $(r_{\mathfrak{F}}^+)$. Consider the standard frame (W_1, \tilde{R}) where \tilde{R} is defined by

$$w\tilde{R}w' \text{ iff } wRf(w') \quad (9)$$

(i.e. we have $\tilde{R} = f^\circ \circ R$, where f° is the converse of f , seen as a relation). This is a finite Kripke frame having \mathcal{S} as a p-morphic image. In fact, (8) is satisfied by taking $\mu := id$ because $f \circ \tilde{R} = f \circ f^\circ \circ R = R$, (we used that $f \circ f^\circ = id$, which holds by the surjectivity of f).

We now show that (W_1, \tilde{R}) validates $(r_{\mathfrak{F}})$ (recall that $(r_{\mathfrak{F}}^+)$ is equivalent to it in standard frames because the two rules are inter-derivable): to this aim, we prove that if there is a surjective R -preserving map μ from (W_1, \tilde{R}) onto $\mathfrak{F} = (F, R_F)$, then \mathcal{S} does not validate $(r_{\mathfrak{F}}^+)$, contrary to the hypothesis. Suppose there is such a μ . Define now a valuation \mathbf{v} by taking $\mathbf{v}(x_a) = \{w \mid \mu(w) = a\} \subseteq W_1$ and

$$\mathbf{v}(r_a) = \{v \in W_0 \mid \forall w (f(w) = v \Rightarrow aR_F\mu(w))\}.$$

The definition is well defined because the variables having at least an occurrence inside a modal operator are precisely the r_a 's, so these variables are evaluated as subsets of W_0 and the other ones as subsets of W_1 . Thus \mathbf{v} evaluates to 1 the formulae $\bigvee_{i=1}^n x_{a_i}$ and $\bigwedge_{i \neq j} \neg(x_{a_i} \wedge x_{a_j})$, whereas $\neg x_{a_1}, \dots, \neg x_{a_n}$ are not evaluated to 1 (because μ is surjective). It remains to check that for every $a \in F$, we have (i) $x_a^{\mathbf{v}1} \subseteq \square r_a^{\mathbf{v}1}$ and (ii) $r_a^{\mathbf{v}1} \subseteq (\bigvee_{b \in R_F(a)} x_b)^{\mathbf{v}1}$. Now (i) holds by (9) and because μ is stable: if $w \in x_a^{\mathbf{v}1}$ and wRv then $v \in r_a^{\mathbf{v}1}$ because if $f(w') = v$ then $w\tilde{R}w'$ and consequently $a = \mu(w)R_F\mu(w')$. To prove (ii), pick $w \in f^*(r_a)$; we have in particular $aR_F\mu(w)$, thus $w \in (\bigvee_{b \in R_F(a)} x_b)^{\mathbf{v}1}$. \square

From Lemma 5.2, we immediately obtain the following result from [3]:

Corollary 5.4 *A modal calculus comprising only rules of the kind $(r_{\mathfrak{F}})$ enjoys the finite model property.*

The following counter-example shows that we really need to replace $(r_{\mathfrak{F}})$ by $(r_{\mathfrak{F}}^+)$ to obtain the bpp.

Example 5.5 Consider the two element reflexive chain

$$\mathfrak{F} := \boxed{b \circ \longrightarrow \circ a}$$

The rule $(r_{\mathfrak{F}})$ simplifies to

$$\frac{x_a \rightarrow \Box x_a}{x_a \mid \neg x_a}$$

This rule is validated in a step frame $\mathcal{S} = (W_1, W_0, f, R)$ iff for every proper subset $a \subseteq W_0$ (i.e., for every subset different from \emptyset, W_0) there is $w \in W_1$ such that $f(w) \in a$ and $R(w) \not\subseteq a$. In a standard frame (W, S) this means that every pair of elements of W are connected via an S -path (to see this, consider as a the set of points which are reachable in $n \geq 0$ steps by any given point and show that such an a must be total). It is not difficult to check that putting $W_1 := \{w_1, w_2\}, W_0 := \{v\}, f(w_1) := f(w_2) := v, R(w_1) := \{v\}, R(w_2) := \emptyset$, we obtain a finite conservative one-step frame that validates $(r_{\mathfrak{F}})$ but cannot be a p-morphic image of a standard frame validating it (because in the latter there cannot be terminal points and any pre-image of w_2 along a p-morphism must be such by (8)). Since the fmp holds for the modal calculus axiomatized by the rule $(r_{\mathfrak{F}})$ according to Corollary 5.4, it is clear that it is the bpp that fails for it (failure of the bpp can also be directly checked by using Theorem 4.5 instead of Theorem 4.7 and Corollary 5.4).

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Appendix

For the proof of the algebraic completeness Theorem 2.5, we need a couple of lemmas:

Lemma .6 *Weakening is admissible: we have $\Gamma \vdash_K S \Rightarrow \Gamma \vdash_K S \mid S'$, for every S' .*

Proof. Trivial by induction on the length of derivation. \square

Lemma .7 *Let Γ be a set of formulae, α a formula, S a hyperformula and K a set of multiple-conclusion rules. If $\Gamma \cup \{\alpha\} \vdash_K S$ and $\Gamma \vdash_K \alpha \mid S$, then $\Gamma \vdash_K S$.*

Proof. Assume $\Gamma \vdash_K \alpha \mid S$. Using weakening, by induction on proof length, it is easy to see that $\Gamma \cup \{\alpha\} \vdash_K \tilde{S}$ implies $\Gamma \vdash_K S \mid \tilde{S}$ for every \tilde{S} . The claim now follows because $S \mid S$ is equal to S (hyperformulae are defined as sets of formulae). \square

Theorem 2.5 *Let K be a set of multiple-conclusion rules. Then $\Gamma \vdash_K S$ iff the multiple-conclusion rule Γ/S is valid in every modal algebra validating K .*

Proof. One direction is trivial. For the other direction, let us suppose that $\Gamma \vdash_K S$ does not hold. By Zorn's lemma, pick $\tilde{\Gamma}$ to be a maximal set of formulae containing Γ such that $\tilde{\Gamma} \not\vdash_K S$. We claim that for every hyperformula $\alpha_1 \mid \dots \mid \alpha_n$

$$\tilde{\Gamma} \vdash_K \alpha_1 \mid \dots \mid \alpha_n \mid S \quad \Rightarrow \quad \exists i \tilde{\Gamma} \vdash_K \alpha_i. \quad (.1)$$

In fact, if this does not hold, by the maximality of $\tilde{\Gamma}$, we have both that $\tilde{\Gamma} \vdash_K \alpha_1 \mid \dots \mid \alpha_n \mid S$ and that $\tilde{\Gamma} \cup \{\alpha_1\} \vdash_K S$. By the above lemma, this implies $\tilde{\Gamma} \vdash_K \alpha_2 \mid \dots \mid \alpha_n \mid S$. Repeating the argument n -times, we obtain $\tilde{\Gamma} \vdash_K S$, contradiction.

Now notice that Lemma .7 and the maximality of $\tilde{\Gamma}$ imply that if $\tilde{\Gamma} \vdash_K \alpha$, then $\alpha \in \tilde{\Gamma}$ and $\Box\alpha \in \tilde{\Gamma}$ (the latter is because necessitation rule is mentioned in condition (ii) of Definition (2.4)). In addition, $\tilde{\Gamma}$ contains Γ and is disjoint from S , by condition (i) of Definition (2.4). Thus, if we put

$$\alpha_1 \approx \alpha_2 \quad \Leftrightarrow \quad \alpha_1 \leftrightarrow \alpha_2 \in \tilde{\Gamma}$$

we can introduce on the set of equivalence classes a modal algebra structure $\mathfrak{A} = (A, \diamond)$. Since Γ is included in $\tilde{\Gamma}$ and is disjoint from S , \mathfrak{A} does not validate Γ/S . By the claim (.1) and condition (ii) of Definition (2.4), it is evident that \mathfrak{A} validates all rules from K . \square

Corollary 2.6 *Let K be a set of multiple-conclusion rules. For each multiple-conclusion rule Γ/Δ , we have $K \vdash \Gamma/\Delta$ iff $\Gamma \vdash_K \Delta$.*

Proof. It is sufficient to observe that (I) if $\Gamma \vdash_K \Delta$, then Γ/Δ belongs to every modal rule system \mathcal{K} containing K and that (II) $\{\Gamma/\Delta \mid \Gamma \vdash_K \Delta\}$ is a modal rule system extending K .

Claim (II) is immediate from Lemmas .6, .7.

Claim (I) is by induction on the length of the K -hyperproof witnessing $\Gamma \vdash_K \Delta$: for instance, if the K -hyperproof ends with an application of the necessitation rule according to Definition 2.4(ii), then from $\Gamma/\alpha, S \in \mathcal{K}$ (this holds by induction hypothesis) and from the fact that the necessitation rule belongs to every modal rule system, from conditions (iii) and (ii) of Definition 2.1, we obtain $\Gamma/\Box\alpha, S \in \mathcal{K}$.¹⁰ \square

We now fill the missing details for the proof of Proposition 2.7:

Proposition 2.7 *Let H be a finite set of hyperrules. Then it is possible to produce a set of rules K such that for all Γ, \tilde{S} we have $\Gamma \Vdash_H \tilde{S}$ iff $\Gamma \vdash_K \tilde{S}$.*

Proof. Consider a hyperrule $S_1, \dots, S_k/S$ from H : to obtain K , we simply replace it with the set of rules $\gamma(S_1), \dots, \gamma(S_n)/S$, varying γ among the functions that pick one formula from each S_i , for each $i = 1, \dots, n$.

The right-to-left claim of the proposition is immediate by weakening. To show the left-to-right direction, we use the argument below. Suppose H' is obtained from H by replacing the hyperrule $S_1, \dots, S_n/S$ with the pair of rules

$$S'_1, S_2, \dots, S_n/S, \quad S''_1, S_2, \dots, S_n/S \quad (.2)$$

where we suppose that S'_1, S''_1 are both not empty and such that $S_1 = S'_1 \cup S''_1$. We claim that we have $\Gamma \Vdash_H S$ iff $\Gamma \Vdash_{H'} S$ (clearly, the statement of the proposition follows from an iterated application of this claim). Again that $\Gamma \Vdash_H \tilde{S} \Leftarrow \Gamma \Vdash_{H'} \tilde{S}$ holds is trivial by weakening. Now suppose that we have $\Gamma \Vdash_H \tilde{S}$. In the derivation witnessing this, there will possibly be lines labelled by $S_1\sigma \mid T, \dots, S_n\sigma \mid T$ justifying a line labelled $S\sigma \mid T$ via the use of the hyperrule $S_1, \dots, S_n/S$. The derivation can be corrected so to use the rules (.2) instead (iterated corrections will eliminate any use of the rule $S_1, \dots, S_n/S$). We first produce (by weakening) derivations of $S_2\sigma \mid S''_1\sigma \mid T$ and \dots and $S_n\sigma \mid S''_1\sigma \mid T$. These hyperformulae, combined with $S'_1\sigma \mid S''_1\sigma \mid T$ yield a derivation of $S''_1\sigma \mid S\sigma \mid T$ by applying the first hyperrule from (.2). By weakening again, we produce now derivations of $S_2\sigma \mid S\sigma \mid T$ and \dots and $S_n\sigma \mid S\sigma \mid T$. These hyperformulae, combined with $S''_1\sigma \mid S\sigma \mid T$ yields a derivation of $S\sigma \mid S\sigma \mid T$ by applying the second hyperrule from (.2) and we are done because $S\sigma \mid S\sigma \mid T$ is equal to $S\sigma \mid T$ (hyperformulae are sets, not multisets). \square

¹⁰ Notice that we added S to α because, according to the remark following Definition 2.4, when we apply the necessitation rule $\alpha/\Box\alpha$, then we deduce $\Box\alpha \mid S$ from a proof line containing the hyperformula $\alpha \mid S$.